



Proximity Gaps for Reed–Solomon Codes

Eli Ben-Sasson* Dan Carmon* Yuval Ishai† Swastik Kopparty ‡
 Shubhangi Saraf§

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Abstract

A collection of sets displays a *proximity gap* with respect to some property if for every set in the collection, either (i) all members are δ -close to the property in relative Hamming distance or (ii) only a tiny fraction of members are δ -close to the property. In particular, no set in the collection has roughly half of its members δ -close to the property and the others δ -far from it.

We show that the collection of affine spaces displays a proximity gap with respect to Reed–Solomon (RS) codes, even over small fields, of size polynomial in the dimension of the code, and the gap applies to any δ smaller than the Johnson/Guruswami–Sudan list-decoding bound of the RS code. We also show near-optimal gap results, over fields of (at least) *linear* size in the RS code dimension, for δ smaller than the unique decoding radius. Finally, we discuss several applications of our proximity gap results to distributed storage, multi-party cryptographic protocols, and concretely efficient proof systems.

We prove the proximity gap results by analyzing the execution of classical algebraic decoding algorithms for Reed–Solomon codes (due to Berlekamp–Welch and Guruswami–Sudan) on a *formal element* of an affine space. This involves working with Reed–Solomon codes whose base field is an (infinite) rational function field. Our proofs are obtained by developing an extension (to function fields) of a strategy of Arora and Sudan for analyzing low-degree tests.

*StarkWare Industries Ltd. {eli,dancar}@starkware.co

†Computer Science Department, Technion. Supported by ERC Project NTSC (742754), NSF-BSF grant 2015782, BSF grant 2018393, and a grant from the Ministry of Science and Technology, Israel and Department of Science and Technology, Government of India. Work done in part while participating in the Simons Institute program on Proofs, Consensus, and Decentralizing Society. yuvali@cs.technion.ac.il

‡Department of Mathematics and Department of Computer Science, Rutgers University. Research supported in part by NSF grants CCF-1540634 and CCF-1814409 and BSF grant 2014359. Work done in part while attending a workshop at the Simons Institute program on Proofs, Consensus, and Decentralizing Society. swastik.kopparty@gmail.com

§Department of Mathematics and Department of Computer Science, Rutgers University. Research supported in part by NSF grants CCF-1540634 and CCF-1909683, BSF grant 2014359, a Sloan research fellowship and the Simons Collaboration on Algorithms and Geometry. shubhangi.saraf@gmail.com

1 Introduction

A variety of protocols, arising in the contexts of interactive proofs, distributed storage and cryptography, give rise to the following problem regarding proximity to a linear code $V \subset \mathbb{F}_q^n$ over a finite field \mathbb{F}_q of minimal relative distance δ_V . These myriad protocols assume oracle access to a batch of vectors $\mathbf{u} = \{u_0, \dots, u_l\} \subset \mathbb{F}_q^n$ and their soundness requires that each and every vector u_i be close to V in relative Hamming distance. Furthermore, soundness deteriorates as a function of the largest distance between some vector u_i and the code V . Thus, we seek protocols that minimize the number of queries to the entries of the vectors in \mathbf{u} , while maximizing the probability of recognizing when some vector u_i is significantly far from V .

The linearity of V suggests a natural heuristic, first explored by Rothblum, Vadhan and Wigderson [RVW13]: sample a uniformly random vector u' in the span of \mathbf{u} (denoted $\text{span}(\mathbf{u})$) and view the distance between u' and V , denoted $\Delta(u', V)$, as a proxy for the maximal distance between some member of \mathbf{u} and V . To argue soundness, we would like to show that if even a single u_i is δ -far from (all members of) V , then a randomly chosen u' is also far from V . Indeed, the paper [RVW13] that suggested this heuristic also showed for any V , that whenever a single u_i is δ -far from V , then nearly all samples u' are at least $\delta/2$ -far from V . Here and henceforth, we use Δ to denote relative Hamming distance and say “ u is δ -close to V ”, denoted $\Delta(u, V) \leq \delta$, when $\Delta(u, v) \leq \delta$ for some $v \in V$; otherwise we say “ u is δ -far from V ” (denoted $\Delta(u, V) > \delta$).

Notice that the result above incurs a $2\times$ degradation in the proximity parameter δ : the worst-case assumption — that some u_i is δ -far from V — implies an average-case distance that is only $\delta/2$. Eliminating the proximity degradation is easy when the field size is exponential in the code length. More concretely, if $q \gg 2^{nH(\delta)}$, where H is the binary entropy function, then a union bound over agreement sets shows that for $\delta < \delta_V$, if u_i is δ -far from V then so are nearly all $u' \in \text{span}(\mathbf{u})$. However, exponential field size is prohibitively large in the context of the motivating applications. Obtaining similar results over fields of sub-exponential size appears to be much more challenging.

A number of works looked at this question and were able to remove the degradation in δ with polynomial field size. Ames et al. [AHIV17] showed that for proximity parameters δ that are smaller than half of the unique-decoding radius of V (i.e., when $\delta < \delta_V/4$), nearly all $u' \in \text{span}(\mathbf{u})$ are δ -far from V . The proximity bound was subsequently improved to $\delta < \delta_V/3$ by Roth and Zémor [RZ18]. Ben-Sasson et al. [BKS18] showed similar results for δ above the unique decoding radius, holding for any $\delta < 1 - \sqrt[4]{1 - \delta_V}$, and the state of the art was given in [BGKS20], holding for any $\delta < 1 - \sqrt[3]{1 - \delta_V}$. In fact, this latter result was shown to be tight for certain RS codes, in particular, of maximal blocklength $n = q$.

Ames et al., who were the first to show that in certain cases the average-case distance of $u' \in \text{span}(\mathbf{u})$ from V is nearly-always equal to the worst-case distance of $u_i \in \mathbf{u}$ from V , also raised the following intriguing question, which is at the focus of our investigation here: For which codes and what range of δ does the following statement hold?

If some $u^ \in \text{span}(\mathbf{u})$ is δ -far from V , then so are nearly all $u' \in \text{span}(\mathbf{u})$.*

One implication of our main result is that when V is an RS code over a sufficiently large field — polynomially large in the code’s blocklength — and when δ is smaller than the Johnson/Guruswami–Sudan list decoding bound, the above phenomenon holds. We refer to it as a *proximity gap*, as explained next.

1.1 Gaps and proximity gaps

When a “gap” is mentioned in theoretical computer science, it usually refers to a situation where all objects under consideration must fall into one of two categories, and these categories display a large gap according to some metric. Striking examples are given by PCP reductions whose outputs are constraint satisfaction problems that lie in one of two categories: satisfiable instances in which some assignment satisfies all constraints, and unsatisfiable instances in which all assignments fail to satisfy more than an ϵ fraction of constraints. Another gap example underlies randomized algorithms. For instance, the Miller–Rabin primality test relies on a gap between primes and composites: in the latter case (composites), at least three-quarters of the integers serve as composite witnesses whereas for primes none do, leading to a “gap” of measure $3/4$.

Our result can be phrased as a *proximity gap* according to the following definition.

Definition 1.1 (Proximity gap). *Let $\mathsf{P} \subset \Sigma^n$ be a property and $\mathsf{C} \subset 2^{\Sigma^n}$ be a collection of sets. Let Δ be a distance measure on Σ^n . We say that C displays a (δ, ϵ) -proximity gap with respect to P under Δ if every $\mathsf{S} \in \mathsf{C}$ satisfies one, and only one, of the following:*

1. $\Pr_{s \in \mathsf{S}}[\Delta(s, \mathsf{P}) \leq \delta] = 1$.
2. $\Pr_{s \in \mathsf{S}}[\Delta(s, \mathsf{P}) \leq \delta] \leq \epsilon$.

We call δ the proximity parameter and ϵ is the error parameter. By default, Δ denotes the relative Hamming distance measure.

Using this definition we can state our main result. Throughout this paper, \mathbb{F}_q denotes the field of size q , and $\text{RS}[\mathbb{F}_q, \mathcal{D}, k]$ is the RS code of dimension $k + 1$ and blocklength $n = |\mathcal{D}|$ containing as its codewords the polynomials of degree $\leq k$, evaluated on \mathcal{D} . We use ρ to denote the rate $\rho = \frac{k+1}{n}$ of the code.

The following result has two parts and each part has its own proof. The first part holds only below the unique decoding radius but has a smaller error parameter, denoted ϵ_{U} ; the second part holds for proximity parameters up to the Johnson/Guruswami–Sudan bound but has a larger error bound ϵ_{J} (the proof of the second part is also significantly harder).

Theorem 1.2 (Proximity Gap for RS codes). *The collection $\mathsf{C}_{\text{Affine}}$ of affine spaces in $\mathbb{F}_q^{\mathcal{D}}$ displays a (δ, ϵ) -proximity gap with respect to the RS-code $V := \text{RS}[\mathbb{F}_q, \mathcal{D}, k]$ of blocklength n and rate $\rho = \frac{k+1}{n}$, for any $\delta \in (0, 1 - \sqrt{\rho})$, and $\epsilon = \epsilon(q, n, \rho, \delta)$ defined as the following piecewise function:*

- **Unique decoding bound:** For $\delta \in \left(0, \frac{1-\rho}{2}\right]$, the error parameter ϵ is

$$\epsilon = \epsilon_{\text{U}} = \epsilon_{\text{U}}(q, n) := \frac{n}{q}. \quad (1.1)$$

- **Johnson bound:** For $\delta \in \left(\frac{1-\rho}{2}, 1 - \sqrt{\rho}\right)$, setting $\eta := 1 - \sqrt{\rho} - \delta$, the error parameter ϵ is

$$\epsilon = \epsilon_{\text{J}} = \epsilon_{\text{J}}(q, n, \rho, \delta) := \frac{(k+1)^2}{\left(2 \min\left(\eta, \frac{\sqrt{\rho}}{20}\right)\right)^7 q} = O\left(\frac{1}{(\eta\rho)^{O(1)}} \cdot \frac{n^2}{q}\right) \quad (1.2)$$

There are two striking aspects to this result. First, the proximity parameter δ can take any value smaller than the famous Johnson/Guruswami–Sudan bound, which is the largest distance for which we know of efficient (list) decoding algorithms. (Looking ahead, the Guruswami–Sudan

algorithm will play a crucial, though non-algorithmic, role in our proofs.) Second, the size of the field needed to achieve this result is relatively small — linear in the blocklength when δ is below the unique decoding radius $\delta < (1 - \rho)/2$ and, for fixed rate, quadratic in blocklength for larger δ up to the list decoding bound.

1.2 Concentration bounds

Theorem 1.2 implies the following concentration bound, saying that for any affine space in which the element farthest from the RS code is within the Johnson/Guruswami–Sudan radius, nearly all elements are at exactly the same distance from the code(!).

For two sets $U, V \subset \Sigma^n$ define the *divergence*¹ of U from V as $D(U, V) := \max_{u \in U} \Delta(u, V)$.

Corollary 1.3 (Concentration bounds). *Let V, q, n, k and ρ be as defined in Theorem 1.2. Let $U \subset \mathbb{F}_q^D$ be an affine space over \mathbb{F}_q and denote $\delta^* := D(U, V)$. If δ^* is smaller than the Johnson/Guruswami–Sudan bound, then nearly all elements of U have distance exactly δ^* from the code. In other words, if $\delta^* \in (0, 1 - \sqrt{\rho})$, then*

$$\Pr_{u \in U} [\Delta(u, V) \neq \delta^*] \leq \epsilon,$$

where $\epsilon = \epsilon(q, n, \rho, \delta^*)$ is as defined in Theorem 1.2.

When the divergence of U from the RS code V is greater than the Johnson/Guruswami–Sudan bound ($\delta^* > 1 - \sqrt{\rho}$) we may still use Theorem 1.2 to conclude that nearly all elements of U are $\approx (1 - \sqrt{\rho})$ -far from V , but what remains an interesting open problem is whether nearly all members of U are maximally far (δ^* -far) from V . An example from [BGKS20] show that this need not be the case for RS codes where $q = O(n)$.

1.3 Correlated agreement

Next, we state the main technical theorem proved in the paper. For proximity parameters in the unique decoding regime this is proved in Theorem 4.1, and for proximity parameters in the list decoding regime this is proved in Theorem 5.1.

Theorem 1.4 (Main Theorem — Correlated agreement over lines (informal)). *Let V, q, n, k and ρ be as defined in Theorem 1.2. For $u_0, u_1 \in \mathbb{F}_q^D$, if $\delta \in (0, 1 - \sqrt{\rho})$ and*

$$\Pr_{z \in \mathbb{F}_q} [\Delta(u_0 + z \cdot u_1, V) \leq \delta] > \epsilon,$$

where ϵ as defined in Theorem 1.2, then there exists $\mathcal{D}' \subset \mathcal{D}$ and $v_0, v_1 \in V$ satisfying

- **Density:** $|\mathcal{D}'|/|\mathcal{D}| \geq 1 - \delta$, and
- **Agreement:** v_0 agrees with u_0 and v_1 agrees with u_1 on all of \mathcal{D}' .

Remark 1.5 (Sampling from extension fields). One may sample z from a finite extension field $\mathbb{F}_{q'}$ of \mathbb{F}_q . In this case, the statement above holds with ϵ_U and ϵ_V modified by replacing q with q' in the denominators of Eqs. (1.1) and (1.2), respectively. Notice that even in this setting, the vectors v_0, v_1 deduced to exist in Theorem 1.2 belong to $\text{RS}[\mathbb{F}_q, \mathcal{D}, k]$, not just in $\text{RS}[\mathbb{F}_{q'}, \mathcal{D}, k]$, because v_0, v_1 have high agreement with $u_0, u_1 \in \mathbb{F}_q^D$. The ability to sample from a larger field (and incur smaller error) applies to the other statements of this section but for simplicity we state all of them using a single field \mathbb{F}_q to both define V and sample z from.

¹ Notice that divergence is not symmetric as can be seen, e.g., when U is a strict subset of V .

Motivated by applications (described later), we generalize the theorem above to two interesting cases: (i) low-degree parameterized curves, and (ii) higher-dimensional affine spaces; details follow.

Correlated agreement over parameterized curves The first extension of Theorem 1.4 is of particular importance for two reasons. First, it leads to derandomized testing of verifiable secret sharing and distributed storage protocols (cf. Section 7.1). Second, it improves the soundness analysis of the Fast RS IOPP (FRI) protocol [BBHR18b], which is used in concretely efficient and transparent (public coin) proof systems [BBHR18a, BBHR19, BCR⁺18, BCG⁺19, COS19]. We discuss this application in Sections 3.2 and 7.2.

Let $\mathbf{u} = \{u_0, \dots, u_l\} \subset \mathbb{F}_q^{\mathcal{D}}$. The *parameterized curve* of degree l that is generated by \mathbf{u} is the following collection of vectors in $\mathbb{F}_q^{\mathcal{D}}$:

$$\text{curve}(\mathbf{u}) := \left\{ u_z := \sum_{i=0}^l z^i \cdot u_i \mid z \in \mathbb{F}_q \right\}.$$

Theorem 1.6 (Correlated agreement for low-degree parameterized curves). *Let V, q, n, k and ρ be as defined in Theorem 1.2. Let $\mathbf{u} = \{u_0, \dots, u_l\} \subset \mathbb{F}_q^{\mathcal{D}}$. If $\delta \in (0, 1 - \sqrt{\rho})$ and*

$$\Pr_{u \in \text{curve}(\mathbf{u})} [\Delta(u, V) \leq \delta] > l \cdot \epsilon,$$

where ϵ is as defined in Theorem 1.2, then there exists $\mathcal{D}' \subset \mathcal{D}$ and $v_0, \dots, v_l \in V$ satisfying

- **Density:** $|\mathcal{D}'|/|\mathcal{D}| \geq 1 - \delta$, and
- **Agreement:** for all $i \in \{0, \dots, l\}$, the functions u_i and v_i agree on all of \mathcal{D}' .

Correlated agreement for affine spaces The second generalization of our Main Theorem 1.4, extends it to an entire affine space. Theorem 1.2 follows directly from the following statement. Notice that Main Theorem 1.4 is actually a case of the following result (for 1-dimensional spaces). However, we stated that special case separately because we prove it first, and from it deduce the more general case (see Section 6.3).

Theorem 1.7 (Correlated agreement over affine spaces). *Let V, q, n, k and ρ be as defined in Theorem 1.2. For $u_0, u_1, \dots, u_l \in \mathbb{F}_q^{\mathcal{D}}$ let $U = u_0 + \text{span}\{u_1, \dots, u_l\} \subset \mathbb{F}_q^{\mathcal{D}}$ be an affine subspace. If $\delta \in (0, 1 - \sqrt{\rho})$ and*

$$\Pr_{u \in U} [\Delta(u, V) \leq \delta] > \epsilon,$$

where ϵ is as defined in Theorem 1.2, then there exists $\mathcal{D}' \subset \mathcal{D}$ and $v_0, \dots, v_l \in V$ satisfying

- **Density:** $|\mathcal{D}'|/|\mathcal{D}| \geq 1 - \delta$, and
- **Agreement:** for all $i \in \{0, \dots, l\}$, the functions u_i and v_i agree on all of \mathcal{D}' .

Furthermore, in the unique decoding regime $\delta \in \left(0, \frac{1-\rho}{2}\right]$, there exists a unique maximal \mathcal{D}' satisfying the above, with unique v_i .

Correlated agreement (Theorem 1.7) is a sufficient condition for proximity gaps with the same error and proximity parameters (Theorem 1.2). We leave as open problems (i) whether correlated agreement is also a necessary condition for a proximity gap. And, if the answer to this question is negative, an intriguing possibility arises: (ii) obtaining proximity gaps for $\delta > 1 - \sqrt{\rho}$ while bypassing the correlated agreement approach we took here.

Organization of the rest of the paper: We start with an overview of the proof of Main Theorem 1.4 in Section 2. In Section 3 we survey several applications of our results. Section 4 gives the (simpler) proof of the unique decoding radius part of Main Theorem 1.4. Section 5 gives the proof of the (harder) list decoding radius part of that theorem, by reducing it to a different, more parameterized format (Appendix A provides the preliminary algebraic setup for the proof). In Section 6 we prove the generalizations of Main Theorem 1.4 to curves (Theorem 1.6) and higher dimensional affine spaces (Theorem 1.7). Section 7 concludes with more details on selected applications — Verifiable Secret Sharing (VSS) and Fast RS IOPs of Proximity (FRI).

2 Proof overview

In this section, we give an overview of our proof strategy of our main result, Theorem 1.4.

Recall the setup. $V = \text{RS}[\mathbb{F}_q, \mathcal{D}, k]$ of degree k polynomials evaluated at the points of $\mathcal{D} \subseteq \mathbb{F}_q$, where $|\mathcal{D}| = n$. We have functions $u_0, u_1 : \mathcal{D} \rightarrow \mathbb{F}_q$ such that for many $z \in \mathbb{F}_q$, the function $u_0 + zu_1$ is δ -close to V . We want to deduce that u_0 and u_1 are themselves close to V .

The main conceptual idea of our analysis is to work with the function field $\mathbb{K} = \mathbb{F}_q(Z)$ with a formal variable Z , and to study the various received words $u_0 + zu_1$ for the code V simultaneously by considering the *formal received word* $w = u_0 + Zu_1 : \mathcal{D} \rightarrow \mathbb{K}$ for the (big field) Reed–Solomon code $\text{RS}[\mathbb{K}, \mathcal{D}, k]$. It turns out that showing that w is close to a (well-structured) codeword of this Reed–Solomon code is sufficient to show that u_0 and u_1 are both close to the original Reed–Solomon code V . With this viewpoint, our proof strategy is to *run a decoding algorithm for Reed–Solomon codes* on this received word $w = u_0 + Zu_1$. Our goal is to analyze the execution of this algorithm to show that it succeeds in finding a nearby Reed–Solomon codeword. We do such an analysis by relating it to the execution of that decoding algorithm on the various received words $u_0 + zu_1$ for the Reed–Solomon code V over the small field \mathbb{F}_q .

This strategy is instantiated with two different decoding algorithms for Reed–Solomon codes: the Berlekamp–Welch unique decoding algorithm, and the Guruswami–Sudan list decoding algorithm [GS99]. Both instantiations give rise to intriguing algebraic questions about polynomials, which we resolve using nontrivial tools from algebraic geometry and the theory of algebraic function fields.

Instantiation with the Berlekamp–Welch Algorithm

Over a field \mathbb{F} and an evaluation domain \mathcal{D} , given a received word $r : \mathcal{D} \rightarrow \mathbb{F}$, the Berlekamp–Welch decoding algorithm for finding the (unique) nearby polynomial $P(X) \in \mathbb{F}[X]$ close to r works as follows. First it searches for low-degree polynomials $A(X), B(X) \in \mathbb{F}[X]$ such that for each $x \in \mathcal{D}$:

$$A(x)r(x) = B(x).$$

Then the nearby polynomial $P(X)$ is recovered as $B(X)/A(X)$ (which a priori may be a rational function).

In our setting, we first run the Berlekamp–Welch algorithm with received word $w = u_0 + Zu_1 : \mathcal{D} \rightarrow \mathbb{K}$ over the big field $\mathbb{K} = \mathbb{F}_q(Z)$ (we will sometimes view this as a function $w(x, z)$ with $w : \mathcal{D} \times \mathbb{F}_q \rightarrow \mathbb{F}_q$). Our goal is to find a nearby Reed–Solomon codeword (low-degree polynomial) $P(X) \in \mathbb{K}[X]$ which has the special form $P_0(X) + ZP_1(X)$, where each $P_i(X) \in \mathbb{F}_q[X]$. The first step of the Berlekamp–Welch algorithm gives us $A(X), B(X) \in \mathbb{K}[X] = \mathbb{F}_q(Z)[X]$. Making the Z dependence explicit, we write these as $A(X, Z), B(X, Z)$. This gives us a candidate, namely $A(X, Z)/B(X, Z)$, for being a Reed–Solomon codeword close to w . We will show two things: that

$A(X, Z)/B(X, Z)$ is a polynomial in $\mathbb{F}_q(Z)[X]$ (a priori it is only a rational function), and that it is close to w .

The crucial step is to substitute $Z = z$ into $A(X, Z)$ and $B(X, Z)$ for various values of $z \in \mathbb{F}_q$. Letting $w_z = u_0 + zu_1 : \mathcal{D} \rightarrow \mathbb{K}$ (the result of substituting $Z = z$ into w), it turns out that $A(X, z), B(X, z) \in \mathbb{F}_q[X]$ are what we would get if we run the Berlekamp–Welch algorithm (over the small field \mathbb{F}_q) on received word w_z . In particular, for many z we get that $B(X, z)$ is divisible by $A(X, z)$ in $\mathbb{F}_q[X]$, and $B(X, z)/A(X, z)$ equals the Reed–Solomon codeword close to w_z . This then allows us to use the Polishchuk–Spielman lemma (a strengthening of the classical Bezout theorem, which deduces divisibility of bivariate polynomials from divisibility of univariate restrictions) to conclude that $B(X, Z)/A(X, Z)$ is in fact a polynomial $P(X, Z)$ in $\mathbb{K}[X]$ of low degree in X .

The final step is to show that $P(X, Z)$, when viewed as a function from \mathcal{D} to \mathbb{K} , is close to w , and that the Z dependence of $P(X, Z)$ is simple (just linear in Z). This is again achieved by considering Z substitutions. We know that for many z , $P(X, z)$ is the degree at most k polynomial $P_z(X)$ that is close to w_z . This means that the X degree of $P(X, z)$ is at most k , and that for many $x \in \mathcal{D}$ and z there is agreement between $P(x, z)$ and $w_z(x) = w(x, z)$. On the other hand, for any $x \in \mathcal{D}$, $w(x, \cdot)$ is a linear function, and $P(x, \cdot)$ is a low degree rational function, and so they cannot agree on too many points unless the low degree rational function $P(x, \cdot)$ formally equals the linear function $w(x, \cdot)$. Therefore this formal equality must happen for many $x \in \mathcal{D}$, i.e., $P(\cdot, Z)$ is close to w . Finally, by simple linear algebra, if $P(x, Z)$ is linear in Z for many x , we conclude that $P(X, Z)$ is linear in Z . This gives us our desired conclusion.

Instantiation with the Guruswami–Sudan Algorithm

Over a field \mathbb{F} and an evaluation domain \mathcal{D} , given a received word $r : \mathcal{D} \rightarrow \mathbb{F}$, the Sudan and Guruswami–Sudan decoding algorithms for finding all nearby polynomials $P(X) \in \mathbb{F}[X]$ close to r work as follows. First one searches for a low-degree polynomial $Q(X, Y) \in \mathbb{F}[X, Y]$ such that for each $x \in \mathcal{D}$,

$$Q(x, r(x)) = 0.$$

(This is the Sudan algorithm; for the Guruswami–Sudan algorithm we ask that Q vanishes at each $(x, r(x))$ with high multiplicity.) Then every nearby polynomial $P(X)$ turns out to have the property that $Y - P(X)$ divides $Q(X, Y)$ in the bivariate polynomial ring $\mathbb{F}_q[X, Y]$. This means that all such $P(X)$ can be found by factoring $Q(X, Y)$.

In our setting, we run the Guruswami–Sudan algorithm with received word $w = u_0 + Zu_1 : \mathcal{D} \rightarrow \mathbb{K}$ over the big field $\mathbb{K} = \mathbb{F}_q(Z)$. Our goal is to find a nearby low-degree polynomial $P(X) \in \mathbb{K}[X]$ which has the special form $P_0(X) + ZP_1(X)$, where each $P_i(X) \in \mathbb{F}_q[X]$. The first step of the Guruswami–Sudan algorithm gives us a bivariate polynomial $Q(X, Y) \in \mathbb{K}[X, Y]$ such that $Q(x, w(x)) = 0$ for each $x \in \mathcal{D}$. Again, we write $Q(X, Y)$ as $Q(X, Y, Z) \in \mathbb{F}_q(Z)[X, Y]$ to make the Z dependence explicit (and we can clear denominators in Z without affecting the vanishing property).

Substituting $Z = z$, we get that $Q(x, w_z(x), z) = 0$ for each $x \in \mathcal{D}$. This means that the polynomial $Q_z(X, Y) \in \mathbb{F}_q[X, Y]$ given by $Q_z(X, Y) = Q(X, Y, z) \in \mathbb{F}_q[X, Y]$ is the bivariate polynomial we would have found while running the Guruswami–Sudan algorithm with received word $w_z : \mathcal{D} \rightarrow \mathbb{F}_q$ over the small field \mathbb{F}_q . Since for many $z \in \mathbb{F}_q$ we have that w_z is close to some codeword $P_z(X) \in \mathbb{F}_q[X]$ of the Reed–Solomon code V , we get that $Y - P_z(X)$ divides $Q(X, Y, z)$ for many $z \in \mathbb{F}_q$. We would like to deduce from this that over the big field \mathbb{K} there is a low-degree polynomial $P(X) \in \mathbb{K}[X]$ such that $Y - P(X)$ divides $Q(X, Y)$ in $\mathbb{K}[X, Y]$ (and furthermore, this $P(X)$ is close to w and has a simple Z dependence).

This is the most involved (and interesting) part of the analysis. We will factor $Q(X, Y, Z)$ completely into linear factors in Y .

$$Q(X, Y, Z) = C(X, Z)(Y - \gamma_1(X, Z))(Y - \gamma_2(X, Z)) \cdots (Y - \gamma_D(X, Z)). \quad (2.1)$$

This is natural to do, because we are searching for factors that are linear in Y . Then we substitute $Z = z$ into this, and we should see $P_z(X)$ as one of the factors.

However, getting such a factorization for $Q(X, Y, Z)$ may not be possible with polynomials $\gamma_i(X, Z)$, and we have to look (far) beyond. What kind of objects should we think of the γ_i as? After getting the $\gamma_i(X, Z)$, we would like to (a) argue about when $\gamma_i(X, Z)$ is a polynomial in X , and (b) substitute $Z = z$ into it and inspect the resulting object. To enable these, we will express $\gamma_i(X, Z)$ in the ring $R = \overline{\mathbb{K}}[[X]]$, the ring of power series in X , whose coefficients are in the algebraic closure of $\mathbb{K} = \mathbb{F}_q(Z)$. The power series in X representation allows us to see when γ_i is a polynomial in X , and the coefficients being simply algebraic functions in Z (such as $\sqrt{Z^3 + Z + 1}$) allows us to reason about substitutions $Z = z$. Having decided on R , it is a simple application of Hensel lifting (after possibly a random shift) to show that a factorization as in (2.1) is possible with the $\gamma_i \in R$.

Rather than describe what happens in full generality, we just sketch what would happen in a special case with most of the action. Suppose \mathbb{F}_q is not of characteristic 2, and we have:

$$Q(X, Y, Z) = Y^2 - (Z^3 + Z + 1)(1 - ZX).$$

Going to the ring R , and letting $\alpha = \sqrt{Z^3 + Z + 1} \in \overline{\mathbb{K}}$, it turns out that $Q(X, Y, Z)$ factors as:

$$\begin{aligned} Q(X, Y, Z) &= \left(Y - \sqrt{Z^3 + Z + 1} \sqrt{1 - ZX} \right) \cdot \left(Y + \sqrt{Z^3 + Z + 1} \sqrt{1 - ZX} \right) \\ &= \left(Y - \left(\alpha - \frac{\alpha \cdot Z}{2} X - \frac{\alpha \cdot Z^2}{16} X^2 + \dots \right) \right) \left(Y + \left(\alpha - \frac{\alpha \cdot Z}{2} X - \frac{\alpha \cdot Z^2}{16} X^2 + \dots \right) \right) \end{aligned}$$

where we used the Taylor series expansion for $\sqrt{1 - ZX}$. Now substitute $Z = z$ for $z \in \mathbb{F}_q$. Substituting values into algebraic functions like α is a slightly delicate operation (which square root do you choose? how do you make these choices consistent for different algebraic functions?), but it can be done using basic concepts from the theory of algebraic function fields. Another tool that we need from the theory of algebraic function fields is an analogue of the degree of a polynomial, to measure complexity of algebraic functions and bound the number of their zeroes. In this sketch we avoid going into any such details.

Doing the substitution gives us:

$$\begin{aligned} Q_z(X, Y) &= Q(X, Y, z) \\ &= \left(Y - \left(\alpha(z) - \frac{\alpha(z) \cdot z}{2} X + \dots + c_i \alpha(z) z^i X^i + \dots \right) \right) \times \\ &\quad \left(Y + \left(\alpha(z) - \frac{\alpha(z) \cdot z}{2} X + \dots + c_i \alpha(z) z^i X^i + \dots \right) \right). \end{aligned}$$

By properties of the Guruswami–Sudan decoding algorithm, we know for all “good” $z \in \mathbb{F}_q$ where w_z is close to some low degree polynomial P_z , we must have that $Y - P_z(X)$ divides $Q_z(X, Y)$. Given the factorization above, one of the following must occur:

1. $P_z(X) = \left(\alpha(z) - \frac{\alpha(z) \cdot z}{2} X - \frac{\alpha(z) \cdot z^2}{16} X^2 + \dots + c_i \alpha(z) z^i X^i + \dots \right),$
2. $P_z(X) = - \left(\alpha(z) - \frac{\alpha(z) \cdot z}{2} X - \frac{\alpha(z) \cdot z^2}{16} X^2 + \dots + c_i \alpha(z) z^i X^i + \dots \right).$

Whichever power series ends up equaling $P_z(X)$, the coefficient of X^{k+1} in that power series must equal 0. In our particular example, we deduce that $c_{k+1}\alpha(z)z^{k+1} = 0$ for some constant c_{k+1} . Assuming c_{k+1} is nonzero in \mathbb{F}_q , we get that $\alpha(z)z^{k+1} = 0$ for every good z . Finally we use the fact that a *nonzero* algebraic functions of low “degree” like $\alpha(Z)Z^{k+1} = \sqrt{Z^3 + Z + 1} \cdot Z^{k+1}$ cannot vanish at too many points z . This means that there cannot be too many good z , contradicting our hypothesis. We conclude that $Q(X, Y, Z)$ cannot equal $Y^2 - (Z^3 + Z + 1)(1 - ZX)$!

A very similar argument derives a contradiction unless $Q(X, Y, Z)$ has a factor of the form $Y - P(X)$ for some $P(X) \in \overline{\mathbb{K}}[X]$ of degree at most k . The only twist is that we may have to focus on the coefficient of some different power X^{k+c} in the power series than the coefficient of X^{k+1} (in case the coefficient of X^{k+1} in the power series is identically 0). To make this argument work, we need to estimate the “degree” of the algebraic functions that appear as coefficients in these power series. This involves a careful study of the Hensel lifting process, especially its effect on the complexity of its coefficients.

The final part of the argument, showing that some $Y - P(X)$ factor of $Q(X, Y, Z)$ is such that $P(X)$ has high agreement with w and all the coefficients of $P(X)$ are linear polynomials in Z , is similar to what happened in the unique decoding case. Instead of using the fact that a low degree rational function and a linear function cannot have high agreement unless they are equal, we use the fact that a low degree algebraic function and a linear function cannot have high agreement unless they are equal. This completes our sketch of the proof.

Technical issues When we actually implement the argument, there are some technical changes we make (both for simplicity and for optimizing parameters). First, we do not do the proof by contradiction, but instead show how to find the factor of the form $Y - P(X)$. Next, instead of directly doing Hensel lifting with Q , we factor Q into irreducible factors over $\mathbb{F}_q[X, Y, Z]$ and focus on a single irreducible factor that is “responsible” for many of the P_z . This helps in that we do not need to factor arbitrarily messy Q ’s completely into linear factors, but only those which have the property that $Q(X, Y, z)$ has a linear factor of the form $Y - P_z(X)$. Finally, instead of arguing over the algebraic closure $\overline{\mathbb{K}}$, we go to a small algebraic extension \mathbb{L} of \mathbb{K} which is rich enough to express all the coefficients of the relevant power series. These changes lead to some simplifications and quantitative improvements in our proofs.

Relationship with the Arora-Sudan low degree test [AS03] A beautiful and fundamental paper of Arora and Sudan [AS03], analyzed the “line vs. line” low degree test for multivariate polynomials in the high error regime. The heart of their paper is a theorem that says that if a function $f : \mathbb{F}_q^2 \rightarrow \mathbb{F}_q$ is such that for most lines L given by $Y = aX + b$ in \mathbb{F}_q^2 the univariate function obtained from restricting f to L (denoted $f|_L$) is close to a low degree univariate polynomial, then f is itself close to a low degree bivariate polynomial. This is closely related to our theorem which deduces a similar conclusion about a received word $w : \mathcal{D} \times \mathbb{F}_q \rightarrow \mathbb{F}_q$, also based on restrictions to lines. Our proof is heavily influenced by the proof in [AS03] (which in turn builds on fundamental results on polynomial factorization and the Hilbert irreducibility theorem by Kalfoten [Kal85, Kal95]). There is one crucial difference in our proof. Our approach is spearheaded by the idea of running all arguments over the big field $\mathbb{K} = \mathbb{F}_q(Z)$ (as opposed to treating Z as another variable over \mathbb{F}_q just like X and Y , as is done in [AS03]). This difference affects our proofs in a tangible sense: our proofs are based on bivariate interpolation over the big field \mathbb{K} rather than trivariate interpolation over the small field \mathbb{F}_q . Inside the analysis, our proofs use power series in one variable over function fields rather than power series in two variables over finite extensions of \mathbb{F}_q . This leads to more involved algebraic tools being needed for our proof (most seriously the use of algebraic function

fields), but also yields three improvements. First, our result is about axis parallel restrictions $Z = z$ (for $z \in \mathbb{F}_q$) instead of more general linear restrictions $Z = aX + b$ (for $a, b \in \mathbb{F}_q$). This simpler form of restriction is important for our applications. Second, our result deduces structure all the way up to the Johnson radius, while the result in Arora-Sudan is to a smaller radius (polynomially worse in terms of agreement parameter). Third, our result works over fields that are quadratic in the degree of the polynomials involved whereas the Arora-Sudan result requires fields that are quartic (at least) in the degree.

3 Applications

Our proximity gap results are motivated by the following general setting. There are several purported codewords $\mathbf{u} = \{u_1, \dots, u_l\} \subset \mathbb{F}_q^n$ of an RS code V . A verifier would like to be assured that they are all close to V . This is done by taking a random linear combination of the u_i and checking its proximity to V . The analysis of this simple test, which arises naturally in a variety of application scenarios, turns out to be surprisingly challenging. Indeed, it is closely related to the proximity gap problem we study in this work.

This batch verification problem arises in two kinds of settings: a *distributed* setting, where entries of \mathbf{u} are split between multiple servers and may not be known to any single entity, and a *centralized* setting, where \mathbf{u} is entirely known to a prover and can be queried by a verifier. We briefly explain the role of proximity gaps in these two types of applications.

In the distributed setting, the coefficients of the random linear combination is either generated by a single verifier or jointly via a distributed coin tossing protocol. Each server then responds with its own share of the output. Verification succeeds if the joint output is a codeword, or alternatively it is close to the code. Examples for applications in the distributed setting include Verifiable Secret Sharing (see Section 7.1) and secure multiparty computation protocols, such as those from [DI06, IPS09]. These applications typically rely on unique decoding and can thus benefit from our near-optimal analysis for this regime. In this type of applications, the main challenge is protecting against an *adaptive* adversary who may choose which servers to corrupt after seeing the coefficients of the random linear combination. To defeat such an adversary, we need to ensure that if at least one of the u_i is far from the code, then (with high probability) so is their random linear combination. If this were not the case, an adaptive adversary could eliminate all inconsistencies by corrupting a small number of servers. Proximity gaps rule out this kind of attack.

In the centralized setting, \mathbf{u} is known to a prover and can be queried by the verifier. A typical realization is using a tree-based succinct cryptographic commitment that binds the prover to a uniquely defined \mathbf{u} and yet enables efficient local opening of symbols queried by the verifier. In this case, the verifier challenges the prover by choosing the coefficients r_i of the random linear combination. The prover, who claims that all u_i are codewords in V , must respond with a valid codeword $u \in V$. The verifier checks that u agrees with $u' = r_1u_1 + \dots + r_lu_l$ by querying a random entry of u and the corresponding entries of \mathbf{u} and checking their consistency. (To amplify soundness, the verifier can query several random entries of u .) Here too, proximity gaps guarantee that if one of the u_i is far from V , then (with high probability) so is u' . This ensures that the verifier detects an inconsistency with high probability. Examples for applications in the centralized setting include communication-efficient proof systems [RVW13, AHIV17, BBHR18b], homomorphic commitment schemes [CDD⁺16], and secure two-party computation protocols [IPS08, HIMV19]. See more in Section 3.2 below.

An appealing feature of the simple “random linear combination” test is that it can be implemented with low communication and computation costs. In particular, in the distributed setting

it suffices for each server to send a single field element to the verifier. In both settings, communicating the l random coefficients is typically not a bottleneck. This random challenge can be made shorter either by using a cryptographic pseudorandom generator or unconditionally by using simple derandomization techniques. In particular, one can generate all coefficients as distinct powers of a single random field element and appeal to the parameterized curves variant of the proximity gap theorem (Theorem 1.6).

Our new proximity gaps imply a tighter analysis of applications that test proximity to RS codes. Generally speaking, in the distributed setting the improved proximity gap bounds imply a constant-factor improvement in the *resilience threshold*, namely the number of corrupted parties that can be tolerated. In the centralized setting, one typically gets constant-factor savings in the overall communication and computation costs. While often ignored in theory-oriented research, the latter kind of improvements can be very significant in the context of practical succinct proof systems.

Why RS codes? Reed–Solomon codes are commonly used in distributed storage, efficient proof systems, and cryptographic protocols. They are useful because of their MDS property, near-linear encoding, and efficient (list)-decoding algorithms. A more qualitative feature of RS codes, which is commonly used in proof systems and cryptography, is the following *multiplication-friendliness* property: when $n = |\mathcal{D}| > 2k$, the pointwise products of codewords in $V = \text{RS}[\mathbb{F}_q, \mathcal{D}, k]$ span a linear code that has nontrivial minimal distance, namely the code $\text{RS}[\mathbb{F}_q, \mathcal{D}, 2k]$.

We now give more concrete examples of applying proximity gaps to analyze batch-verification tasks that arise in different application scenarios.

3.1 Distributed storage and cryptography

Distributed storage. Consider a scenario in which l users encode their inputs using a length- n RS code $V = \text{RS}[\mathbb{F}_q, \mathcal{D}, k]$, where server i stores the i -th symbol of each of the l codewords. Suppose that some of the nl symbols were corrupted, say by a transient malware that overwrites a subset of the symbols before being discovered and eliminated. A verifier would like to get a quick estimate of the amount of damage caused by the malware. A natural idea is to have the servers communicate a random linear combination u' of the potentially corrupted codewords u_j . Using the basic proximity gap result (Theorem 1.2), if at least one of u_j is δ -far from the code (for $\delta \leq \frac{1-\rho}{2}$ or $\delta < 1 - \sqrt{\rho}$), then u' is δ -far from the code except with small failure probability (at most n/q for $\delta \leq \frac{1-\rho}{2}$). Thus, for sufficiently large \mathbb{F}_q , the distance of u' from V provides a reliable upper bound on the maximal relative distance of a vector u_i from V within the proximity bounds of Theorem 1.2. This estimate is not too pessimistic in the sense that if only a μ -fraction of the servers were affected, the upper bound obtained by the test is no bigger than μ .

Distributed proximity test for Interleaved RS codes. The above analysis leaves something to be desired: if u' is within (sufficiently small) distance δ from V , the verifier is only assured that each u_j is *individually* within distance δ from V . In some applications, we would like to get the stronger guarantee that in such an event there is a δ -fraction of the coordinates whose removal makes *all* u_j consistent with V . Moreover, we would like to identify this set of coordinates, which is uniquely defined in the unique decoding regime. This is useful even in the above distributed storage scenario, but will be even more useful for the applications we discuss next. The stronger feature can be conveniently captured using the notion of an *Interleaved Reed–Solomon* (IRS) code. In an $\text{IRS}(V, l)$ code, the codewords are $l \times n$ matrices in which each row is a codeword in V . The symbols of such a codeword are the matrix columns. Namely, a codeword consists of n symbols in

\mathbb{F}_q^ℓ . The following theorem, which follows easily from Theorem 1.7, phrases the stronger guarantee provided by the refined analysis in terms of proximity testing for IRS codes. We state it for the unique decoding regime, which suffices (and is sometimes required) for the applications we discuss next. For u within the unique decoding radius of V , we denote by $\Gamma(u, V)$ the set of coordinates on which u disagrees with the closest codeword from V .

Theorem 3.1 (Distributed proximity test for Interleaved RS codes). *Let $V = \text{RS}[\mathbb{F}_q, \mathcal{D}, k]$ for $|\mathcal{D}| = n$ and $\mathbf{V} = \text{IRS}(V, l)$. We view codewords in V and \mathbf{V} as vectors in \mathbb{F}_q^n and matrices in $\mathbb{F}_q^{l \times n}$ respectively. Let $\rho = \frac{k+1}{n}$ and $\delta \leq \frac{1-\rho}{2}$. Let $\mathbf{u} \in \mathbb{F}_q^{l \times n}$ and let $u' = r^T \mathbf{u}$ where $r \in_R \mathbb{F}_q^l$.*

- *Completeness: If $\Delta(\mathbf{u}, \mathbf{V}) \leq \delta$ then $\Pr[\Delta(u', V) \leq \delta] = 1$ and moreover $\Pr[\Gamma(u', V) \neq \Gamma(\mathbf{u}, \mathbf{V})] \leq n/q$.*
- *Soundness: If $\Delta(\mathbf{u}, \mathbf{V}) > \delta$ then $\Pr[\Delta(u', V) \leq \delta] \leq n/q$.*

We refer to the above test as distributed because it can be implemented with low communication complexity in the distributed setting, where each server holds a different column of \mathbf{u} . One can similarly obtain an affine version with the same guarantee, where \mathbf{u} has an additional row u_0 that is always added to u' (i.e., with coefficient $r_0 = 1$), and the code \mathbf{V} is extended to by $\text{IRS}(V, l + 1)$. This affine version is useful for zero-knowledge variants of the test, where a single random $u_0 \in V$ is used for blinding u_1, \dots, u_l . This is used in the cryptographic applications we discuss next.

General cryptographic protocols. Theorem 3.1 serves as a useful tool for analyzing cryptographic protocols in the presence of an *adaptive* adversary who can dynamically choose the set of corrupted parties. For instance, it shows that secure multiparty computation protocols from [DI06, IPS09] are adaptively secure when the adversary can corrupt roughly 1/3 of the parties. The best previous proximity gaps from [RZ18, BKS18, BGKS20] could only get up to 1/4 corruption threshold in the same setting. Adaptive security, in turn, is crucial for the general transformation from [IKOS09, IPS08] of these honest-majority protocols to two-party protocols and protocols for dishonest majority. Indeed, this is the context that gave rise to proximity gap in the analysis of the Liger zero-knowledge proof system [AHIV17], which applies a variant of the transformation from [IKOS09] to a variant of the protocol from [DI06]. We give a detailed exposition of the application of proximity gaps to *verifiable secret sharing*, which serves as a basis for the above results on secure multiparty computation, in Section 7.

3.2 Soundness of the Fast RS IOPP (FRI) protocol

FRI is an Interactive Oracle Proof of Proximity (IOP of Proximity, or IOPP) as defined in [RRR16, BCS16]. An IOP is an interactive protocol in which the verifier has oracle access to messages sent by the prover, so she need not read and store those messages but may query random entries of them. FRI is one of a family of protocols for testing proximity to the RS code (an “RS proximity testing” (RPT) protocol). Its purpose is to check whether a received word $f : \mathcal{D} \rightarrow \mathbb{F}_q$ belongs to a pre-specified RS code $V := \text{RS}[\mathbb{F}_q, \mathcal{D}, k]$ and to reject words that are δ -far from the code with high probability and low query complexity. Due to its efficiency it is used as a building block in several recent succinct zero knowledge protocols including scalable and transparent (public coins) arguments of knowledge (STARKs) [BBHR18a, BBHR19], Aurora [BCR⁺18] and its succinct version [BCG⁺19], and Fractal [COS19], to name a few. These systems have been shown by Chiesa

et al. to be sound in the quantum random oracle model (hence are “plausibly post-quantum secure”) [CMS19]. Therefore, understanding the concrete soundness error of FRI, denoted ϵ_{FRI} , is of significant practical value, in addition to being a theoretically interesting question.

Consider the case of f that is maximally far from V , i.e., $\Delta(f, V) \approx 1 - \rho$ (this holds, e.g., for random f , with high probability). Fix a target soundness error bound $2^{-\lambda}$ (in concrete settings, λ is the “security parameter”, often fixed to $\lambda = 128$). The communication complexity of FRI is dominated by the number t of iterations of the QUERY phase, so the question at hand is:

How many iterations t of the QUERY phase are needed to obtain $\epsilon_{\text{FRI}} \leq 2^{-\lambda}$?

The initial analysis of [BBHR18b] required a number t that is quite large, and does not tend to 0 even for tiny rates ρ . This was improved by [BKS18] to $t \approx 4\lambda \log \frac{1}{\rho}$, and then by [BGKS20] to $t \approx 3\lambda \log \frac{1}{\rho}$. Sadly, that paper also showed that this bound is tight, at least when the field size q equals the code’s blocklength n . Our main result regarding FRI (Theorem 7.2) shows that for $q \gg n^2$ we can reduce the number t of iterations by 33% to $t \approx 2\lambda \log \frac{1}{\rho}$, which leads to communication complexity that is at least 33% shorter, for provable soundness settings. The actual savings in the provable soundness case are likely larger, due to smaller field size and the ability of the improved analysis to operate with any sequence of oracle sizes in the FRI COMMIT phase (as discussed after the statement of Theorem 7.2).

4 Correlated Agreement over Lines — Unique Decoding Radius

In this section we prove the correlated agreement result for proximity parameters that are below the unique decoding radius, corresponding to the $\epsilon = \epsilon_{\text{U}}$ part of Theorem 1.4. In this case, where $\delta \in \left(0, \frac{1-\rho}{2}\right]$, our result holds even with fields that are merely *linear* in the blocklength of the code. More importantly, the proof will present several ideas, in simplified form, that will appear again in the proof of harder, list decoding regime, result (Theorem 5.1).

As usual, let \mathbb{F}_q be the finite field of size q , let $\mathcal{D} \subseteq \mathbb{F}_q$ be an evaluation domain of size $|\mathcal{D}| = n$, let $k \leq n$, and let $V = \text{RS}[\mathbb{F}_q, \mathcal{D}, k]$ be the Reed–Solomon code of rate $\rho = \frac{k+1}{n}$.

Theorem 4.1. *Suppose $\delta \leq (1 - \rho)/2$. Let $u_0, u_1 : \mathcal{D} \rightarrow \mathbb{F}_q$ be functions. Let*

$$S = \{z \in \mathbb{F}_q : \Delta(u_0 + zu_1, V) \leq \delta\}$$

and suppose $|S| > n$. Then $S = \mathbb{F}_q$. Furthermore there are $v_0, v_1 \in V$ such that for all $z \in \mathbb{F}_q$,

$$\Delta(u_0 + zu_1, v_0 + zv_1) \leq \delta$$

and in fact

$$|\{x \in \mathcal{D} : (u_0(x), u_1(x)) \neq (v_0(x), v_1(x))\}| \leq \delta|\mathcal{D}|.$$

Remark 4.2. Since $\delta \leq \frac{1-\rho}{2}$ is within the unique decoding regime, the above v_0, v_1 , which are simultaneously and separately δ -close to u_0, u_1 , are also unique.

4.1 The Berlekamp–Welch decoder

Our proof will be based on the Berlekamp–Welch decoding algorithm. Let \mathbb{F} be a (general) field and $\mathcal{D} \subseteq \mathbb{F}$. For an integer k , consider the Reed–Solomon code $V = \text{RS}[\mathbb{F}, \mathcal{D}, k]$. We will be instantiating the Berlekamp–Welch decoder for RS codes over two different fields: the “standard” field \mathbb{F}_q and the

field of rational functions $\mathbb{K} = \mathbb{F}_q(Z)$ in the formal variable Z . We now give a quick description of the Berlekamp–Welch decoding algorithm and some useful aspects of it.

Given a received word $w : \mathcal{D} \rightarrow \mathbb{F}$, where $\mathcal{D} \subseteq \mathbb{F}$, and an error parameter $e = \delta n \leq \frac{n-k-1}{2}$, the Berlekamp–Welch decoder finds the unique (if any) polynomial $P(X) \in \mathbb{F}[X]$ such that $\Delta(w, P) \leq e$.

The first step of the Berlekamp–Welch decoder is to set up a homogeneous system of linear equations to find polynomials $A(X), B(X) \in \mathbb{F}[X]$ with $\deg(A) \leq e$, $\deg(B) \leq k + e$ such that:

$$A(x)w(x) = B(x)$$

for all $x \in \mathcal{D}$.

Lemma 4.3. *The homogeneous system of linear equations above has the following properties:*

1. Suppose $\Delta(w, V) \leq \delta$. Then the system of equations has a nonzero solution.
2. Suppose $\Delta(w, V) \leq \delta$. Then for any nonzero solution $A(X), B(X)$ to the system of equations, we have that $A(X)$ divides $B(X)$ (in $\mathbb{F}[X]$), and furthermore the element of V which realizes the distance is $B(X)/A(X)$.
3. If $A(X), B(X)$ is a nonzero solution to the system of equations such that $A(X)$ divides $B(X)$ (in $\mathbb{F}[X]$), then the polynomial $P(X) = B(X)/A(X)$ has the property² that $\Delta(w, P) \leq \delta$.

These properties above are well known and we omit the proof.

4.2 The Polishchuk–Spielman lemma

Another ingredient that will appear in our proof is a version of the Polishchuk–Spielman lemma [PS94]. The version we state below is a variation of [Spi95, Lemma 4.2.18], and we include a derivation in Appendix D.

Lemma 4.4. *Let $A(X, Z), B(X, Z) \in \mathbb{F}_q[X, Z]$ be polynomials. Suppose there are at least n_X choices of $x \in \mathbb{F}_q$ such that $A(x, Z)$ divides $B(x, Z)$ as polynomials in $\mathbb{F}_q[Z]$, and at least n_Z choices of $z \in \mathbb{F}_q$ such that $A(X, z)$ divides $B(X, z)$ as polynomials in $\mathbb{F}_q[X]$. If*

1. $\deg_X(A) + \deg_X(B) < n_X$,
2. $\deg_Z(A) + \deg_Z(B) < n_Z$,
3. $\frac{\deg_X(B)}{n_X} + \frac{\deg_Z(B)}{n_Z} < 1$,

then $A(X, Z)$ divides $B(X, Z)$ as polynomials in $\mathbb{F}_q[X, Z]$.

² Note that $P(X)$ may have degree larger than k . The best we can say about the degree of $P(X)$ is that it is at most $k + e$.

4.3 Proof of Theorem 4.1

By definition of S , for each $z \in S$, we have a polynomial $P_z(X) \in \mathbb{F}_q[X]$ with $\deg(P_z) \leq k$ such that $\Delta(u_0 + zu_1, P_z) \leq \delta$.

Our strategy is to run the Berlekamp–Welch decoder over the field $\mathbb{K} = \mathbb{F}_q(Z)$ of rational functions in the formal variable Z .

First define a received word

$$w : \mathcal{D} \rightarrow \mathbb{K}$$

given by:

$$w(x) = u_0(x) + Zu_1(x).$$

We sometimes also use the notation $w(x, Z)$ to denote $u_0(x) + Zu_1(x)$.

We will try to find a polynomial $P(X, Z) \in \mathbb{F}_q[X, Z]$ of the form $P(X, Z) = v_0(X) + Zv_1(X)$, where $\deg_X(P) \leq k$, such that

$$P(x, Z) = w(x)$$

for at least $n - e$ choices of $x \in \mathcal{D}$.

4.3.1 Step 1: Finding $A(X, Z), B(X, Z)$

The first step of the Berlekamp–Welch algorithm is to find nonzero $A(X, Z), B(X, Z) \in \mathbb{K}[X]$ of degrees $\leq e$ and $\leq k + e$ (in the variable X) respectively such that

$$A(x, Z)w(x) = B(x, Z) \tag{4.1}$$

for all $x \in \mathcal{D}$. Setting this up as a homogeneous linear system over \mathbb{K} , we get an $n \times (k + 2e + 2)$ matrix $M(Z)$ with entries being polynomials in Z , of degree ≤ 1 for the $e + 1$ columns of the A -variables, and degree 0 for the $k + e + 1$ columns of the B variables.

We now show that $M(Z)$ has rank $< k + 2e + 2$ over \mathbb{K} . Fix any $(k + 2e + 2) \times (k + 2e + 2)$ minor of $M(Z)$, and consider its determinant $R(Z) \in \mathbb{F}_q[Z]$. We will show that $R(Z) = 0$. This then implies that $M(Z)$ has rank $< k + 2e + 2$ over \mathbb{K} .

For any $z \in S$, consider $M(z)$. This is the homogeneous linear system that arises when we run the Berlekamp–Welch decoder with received word $u_0 + zu_1 \in \mathbb{F}_q^n$ over the field \mathbb{F}_q . By definition of S we know that $\Delta(u_0 + zu_1, V) \leq \delta$, and so Item 1 of Lemma 4.3 tells us that this linear system has a nonzero solution. Therefore $M(z)$ has rank $< k + 2e + 2$. Thus for each $z \in S$, $R(z) = 0$. Now notice that $\deg(R) \leq e + 1$. Since $|S| > e + 1$, we conclude that $R(Z) = 0$ formally, as desired.

Solving the linear system (4.1) over \mathbb{K} by Cramer’s rule (and clearing the denominator), we get $A(X, Z), B(X, Z) \in \mathbb{F}_q[X, Z]$ with $\deg_Z(A) \leq e$, $\deg_Z(B) \leq e + 1$, $\deg_X(A) \leq e$, $\deg_X(B) \leq k + e$ and:

$$A(x, Z)w(x) = B(x, Z)$$

for all $x \in \mathcal{D}$. Using our alternate notation for w , we get:

$$A(x, Z)w(x, Z) = B(x, Z)$$

for all $x \in \mathcal{D}$. Note that in particular, $A(x, Z)$ divides $B(x, Z)$ (as polynomials in $\mathbb{F}_q[Z]$) for all $x \in \mathcal{D}$.

4.3.2 Step 2: Dividing $B(X, Z)$ by $A(X, Z)$ in $\mathbb{F}_q[X, Z]$

Now if $z \in S$, we know that the function $w(\cdot, z) : \mathcal{D} \rightarrow \mathbb{F}_q$ has distance $\leq e$ from V . Consider the Berlekamp–Welch system of linear equations associated with this received word: namely, we consider the space of all pairs of polynomials $E(X), F(X)$ of degrees $\leq e, \leq k + e$ respectively, such that

$$E(x)w(x, z) = F(x)$$

for all $x \in \mathcal{D}$. We see that any solution $E(X), F(X)$ of this system falls into one of two cases:

- $E(X) = 0$, in which case $F(x) = 0$ for all $x \in \mathcal{D}$, and so $F(X) = 0$ too.
- $E(X)$ is nonzero, in which case Item 2 of Lemma 4.3 tells us that $E(X)$ divides $F(X)$, and $\frac{F(X)}{E(X)} = P_z(X)$.

In both these cases, $E(X)$ divides $F(X)$.

For any $z \in S$, $A(X, z), B(X, z)$ are polynomials that satisfy the properties of E, F above, thus $A(X, z)$ divides $B(X, z)$.

To recap, we've seen that $A(x, Z)$ divides $B(x, Z)$ for $|\mathcal{D}| = n$ values of x , and $A(X, z)$ divides $B(X, z)$ for $|S|$ values of z . Since

$$\begin{aligned} \deg_X(A) + \deg_X(B) &\leq k + 2e < n = |\mathcal{D}|, \\ \deg_Z(A) + \deg_Z(B) &\leq 2e + 1 \leq n < |S|, \text{ and} \\ \frac{\deg_X B}{|\mathcal{D}|} + \frac{\deg_Z B}{|S|} &< \frac{k + e}{n} + \frac{e}{n} < 1, \end{aligned}$$

we may apply the Polishchuk–Spielman lemma 4.4 to get that $A(X, Z)$ divides $B(X, Z)$ in $\mathbb{F}_q[X, Z]$, and define $P(X, Z) = B(X, Z)/A(X, Z) \in \mathbb{F}_q[X, Z]$.

4.3.3 Step 3: $P(X, Z)$ has X -degree at most k

We will now bound the X degree of $P(X, Z) \in \mathbb{F}_q[X, Z]$. The idea is to substitute values for Z .

Observe that $\deg_X(P) \leq k + e$, and $\deg_Z(P) \leq e + 1$. Write $P(X, Z) = \sum_{i=0}^{k+e} P_i(Z)X^i$, where each $P_i(Z) \in \mathbb{F}_q[Z]$ is of degree at most $e + 1$.

Let $S' = \{z \in S : A(X, z) \neq 0\}$. Since $\deg_Z A \leq e$, we have $|S'| \geq |S| - e$.

For $z \in S'$, the discussion above about solutions E, F to the Berlekamp–Welch system of linear equations tells us that $B(X, z)/A(X, z) = P_z(X)$. Thus $P(X, z) = P_z(X)$, and so $P_i(z) = 0$ for all $i > k$.

Since $|S'| > e + 1$, this implies that $P_{k+1}(Z), \dots, P_{k+e}(Z)$ are all identically zero. Thus $P(X, Z)$ has X -degree at most k .

4.3.4 Step 4: $P(x, Z) = w(x, Z)$ for many x

Now we show that P is close to w . We are effectively using Item 3 of Lemma 4.3 here, but we reprove it because we need to keep track of some other information.

Let $\mathcal{D}' = \{x \in \mathcal{D} : A(x, Z) \neq 0\}$. Since $A(X, Z) \neq 0$ and $\deg_X(A) \leq e$, we have $|\mathcal{D}'| \geq n - e > k$.

Note that for any $x \in \mathcal{D}$, we have both $A(x, Z)w(x, Z) = B(x, Z)$ by the decoder's construction, as well as $A(x, Z)P(x, Z) = B(x, Z)$. Comparing the two equations yields

$$A(x, Z)P(x, Z) = A(x, Z)w(x, Z), \quad x \in \mathcal{D}.$$

If $x \in \mathcal{D}'$, then $A(x, Z)$ is not identically zero and is invertible in \mathbb{K} . We thus conclude that

$$P(x, Z) = w(x, Z) \tag{4.2}$$

for all $x \in \mathcal{D}'$.

4.3.5 Step 5: $P(X, Z)$ has Z -degree at most 1

Let $\{x_0, \dots, x_k\} \subseteq \mathcal{D}'$ be any set of $k + 1$ distinct elements of \mathcal{D}' . Let $v_0(X), v_1(X) \in \mathbb{F}_q[X]$ be the unique degree $\leq k$ interpolations of u_0, u_1 at the points x_0, \dots, x_k . Observe that for each $0 \leq i \leq k$ we have

$$P(x_i, Z) = w(x_i, Z) = u_0(x_i) + Zu_1(x_i) = v_0(x_i) + Zv_1(x_i).$$

It follows that the two degree $\leq k$ polynomials $P(X, Z), v_0(X) + Zv_1(X) \in \mathbb{K}[X]$ agree at $k + 1$ points — thus they must be identical, i.e. $P(X, Z) = v_0(X) + Zv_1(X)$ identically.

Using equation (4.2) again, we find for all $x \in \mathcal{D}'$

$$w(x, Z) = P(x, Z) = v_0(x) + Zv_1(x),$$

and since $|\mathcal{D} \setminus \mathcal{D}'| \leq e = \delta|\mathcal{D}|$, we have

$$\Delta(u_0 + Zu_1, v_0 + Zv_1) \leq \delta,$$

as claimed. This completes the proof of Theorem 4.1. \square

5 Correlated Agreement over Lines — List Decoding Radius

In this section we prove the large distance part of the correlated agreement theorem (Theorem 1.4), corresponding to larger proximity parameters $\delta \in \left(\frac{1-\rho}{2}, 1 - \sqrt{\rho}\right)$. First, we state the theorem in a slightly different form that will be easier to work with.

Theorem 5.1. *Let $u_0, u_1 : \mathcal{D} \rightarrow \mathbb{F}_q$, let $m \geq 3$, define*

$$\delta_0(\rho, m) := 1 - \sqrt{\rho} - \frac{\sqrt{\rho}}{2m}, \tag{5.1}$$

and let $\delta \leq \delta_0(\rho, m)$. Define

$$S = \{z \in \mathbb{F}_q : \Delta(u_0 + zu_1, V) \leq \delta\} \tag{5.2}$$

and suppose

$$|S| > \frac{(1 + \frac{1}{2m})^7 m^7}{3\rho^{3/2}} n^2. \tag{5.3}$$

Then u_0, u_1 are simultaneously δ -close to V , i.e. $\exists v_0, v_1 \in V$ such that

$$|\{x \in \mathcal{D} : (u_0(x), u_1(x)) = (v_0(x), v_1(x))\}| \geq (1 - \delta)|\mathcal{D}|.$$

The above version easily implies the large distance part of Theorem 1.4 with the coarser

$$\epsilon_J = O\left(\frac{1}{(\eta\rho)^{O(1)}} \cdot \frac{n^2}{q}\right)$$

bound by setting $m = O\left(\frac{\sqrt{\rho}}{\eta}\right)$. For the more precise bound on ϵ_J , we need to be a little careful, and we do this in the following theorem.

Theorem 5.2 (Correlated agreement over lines — alternative formulation). *Let $u_0, u_1 : \mathcal{D} \rightarrow \mathbb{F}_q$. Let $\delta, \eta > 0$ satisfy $\eta \leq \frac{\sqrt{\rho}}{20}$ and $\delta \leq \delta_0(\rho, \eta) := 1 - \sqrt{\rho} - \eta$, and suppose*

$$\mathbb{P}_{z \in \mathbb{F}_q}(\Delta(u_0 + zu_1, V) \leq \delta) > \frac{\rho^2 n^2}{(2\eta)^7 q} =: \epsilon_J. \quad (5.4)$$

Then u_0, u_1 are simultaneously δ -close to V , i.e. $\exists v_0, v_1 \in V$ such that

$$|\{x \in \mathcal{D} : (u_0(x), u_1(x)) = (v_0(x), v_1(x))\}| \geq (1 - \delta)|\mathcal{D}|.$$

Proof of Theorem 5.2 from Theorem 5.1. Set $m = \left\lceil \frac{\sqrt{\rho}}{2\eta} \right\rceil \geq 10$, and note that $\delta \leq \delta_0(\rho, \eta) < \delta_0(\rho, m)$. Define S as in Theorem 5.1, and observe that (5.3) is satisfied:

$$|S| > \epsilon_J q = (2\eta)^{-7} \rho^2 n^2 > \left(\frac{m-1}{\sqrt{\rho}} \right)^7 \rho^2 n^2 = \left(1 - \frac{1}{m}\right)^7 \frac{m^7}{\rho^{3/2}} n^2 > \frac{\left(1 + \frac{1}{2m}\right)^7 m^7}{3\rho^{3/2}} n^2,$$

where in the last step we use $3\left(1 - \frac{1}{m}\right)^7 > \left(1 + \frac{1}{2m}\right)^7$, which holds for $m \geq 10$. Thus we may apply Theorem 5.1, and conclude (u_0, u_1) is δ -close to (v_0, v_1) , as claimed. \square

5.1 The Guruswami–Sudan decoder

Our proof will be based on the Guruswami–Sudan decoding algorithm. Let \mathbb{F} be a (general) field and $\mathcal{D} \subseteq \mathbb{F}$. Let V be the Reed–Solomon code $\text{RS}[\mathbb{F}, \mathcal{D}, k]$. Let $\rho = \frac{k+1}{n}$ denote its rate. We will be instantiating the Guruswami–Sudan decoder for RS codes over two different fields: the “standard” field \mathbb{F}_q and the field of rational functions $\mathbb{K} = \mathbb{F}_q(Z)$ in the formal variable Z . We now give a quick description of the Guruswami–Sudan decoding algorithm and some useful aspects of it.

First some definitions related to bivariate polynomials. The (a, b) -weighted degree of a monomial $X^i Y^j$ is $ai + bj$. The (a, b) -weighted degree of a polynomial $Q(X, Y) \in \mathbb{F}[X, Y]$ is the maximal (a, b) -weighted degree of all its non-zero monomials. The vanishing multiplicity of a polynomial $Q(X, Y) \in \mathbb{F}[X, Y]$ at a point $(x, y) \in \mathbb{F}^2$ is the smallest m such that the shifted polynomial $Q(x + X, y + Y)$, written as:

$$Q(x + X, y + Y) = \sum_{i,j} a_{i,j} X^i Y^j$$

has $a_{ij} = 0$ for all (i, j) with $i + j < m$. We denote the vanishing multiplicity of Q at (x, y) by $\text{mult}(Q, (x, y))$.

Given a received word $w : \mathcal{D} \rightarrow \mathbb{F}$ and a multiplicity parameter m , the Guruswami–Sudan decoder first solves a homogeneous system of linear equations to find a nonzero polynomial $Q(X, Y) \in \mathbb{F}[X, Y]$ with $(1, k)$ -weighted degree less than $D_X(m)$ (for a certain function $D_X(m)$, specified later), such that:

$$\text{mult}(Q, (x, w(x))) \geq m$$

for all $x \in \mathcal{D}$.

The key properties of this system of linear equations that enable decoding are given by the following lemma.

Lemma 5.3. *Let $\delta_0(\rho, m) = 1 - \sqrt{\rho} - \frac{\sqrt{\rho}}{2m}$. With $D_X(m) = (m + \frac{1}{2})\sqrt{\rho}n$, the system of linear equations set up above has the following properties:*

1. *The system has a nonzero solution $Q(X, Y)$.*

2. For any nonzero solution $Q(X, Y)$ of the above system, and for any polynomial $P(X) \in V$ such that $\Delta(w, P) \leq \delta_0(\rho, m)$, we have that $Y - P(X)$ divides $Q(X, Y)$ in the polynomial ring $\mathbb{F}[X, Y]$.

Note that these choices of δ_0 and D_X are not quite optimal. The optimal values are only slightly better, but their formulas are longer and messier, and we opt for simplicity in favor of optimization.

5.2 Proof of Theorem 5.1

By definition of S , for each $z \in S$, we have a polynomial $P_z(X) \in \mathbb{F}_q[X]$ with $\deg(P_z) \leq k$ such that $\Delta(u_0 + zu_1, P_z) \leq \delta$.

Our strategy is to run the Guruswami–Sudan decoder over the field $\mathbb{K} = \mathbb{F}_q(Z)$ of rational functions in the formal variable Z .

First define a received word

$$w : \mathcal{D} \rightarrow \mathbb{K}$$

given by:

$$w(x) = u_0(x) + Zu_1(x).$$

We sometimes also use the notation $w(x, Z)$ to denote $u_0(x) + Zu_1(x)$.

We will try to find a polynomial $P(X, Z) \in \mathbb{F}_q[X, Z]$ of the form $P(X, Z) = v_0(X) + Zv_1(X)$, where $\deg_X(P) \leq k$, such that

$$P(x, Z) = w(x)$$

for at least $n - e$ choices of $x \in \mathcal{D}$.

5.2.1 Step 1: Interpolating $Q(X, Y, Z)$

Let $D_X = D_X(m) = (m + \frac{1}{2})\sqrt{\rho n}$. The first step of the Guruswami–Sudan decoding algorithm is to find a nonzero polynomial $Q(X, Y) \in \mathbb{K}[X, Y]$:

$$Q(X, Y) = \sum_{i+k \cdot j \leq D_X} Q_{ji}(Z) X^i Y^j,$$

where each $Q_{ji}(Z)$ lies in the big field \mathbb{K} , such that $Q(X, Y)$ has a zero of multiplicity m at $(x, w(x))$ for each $x \in \mathcal{D}$.

This is possible when the number of available monomials, which is at least $\frac{k}{2} \left(\left(\frac{D_X}{k} + \frac{1}{2} \right)^2 - \frac{1}{4} \right)$, exceeds the number of homogeneous linear equations $\binom{m+1}{2} n$. Indeed, this happens for our choice of D_X .

Solving this system of equations for a nonzero solution using Cramer’s rule, and clearing Z denominators, we get such a $Q(X, Y) \in \mathbb{K}[X, Y]$ where each coefficient $Q_{ji}(Z)$ is in fact an element of $\mathbb{F}_q[Z]$ (i.e., a polynomial instead of just a rational function) with controlled degree.

Explicitly, we get:

Claim 5.4. *There is a nonzero polynomial $Q(X, Y) \in \mathbb{K}[X, Y]$ with $(1, k)$ -weighted degree less than D_X such that for each $x \in \mathcal{D}$, we have:*

$$\text{mult}(Q, (x, w(x))) \geq m,$$

and furthermore:

•

$$\deg_X(Q) < D_X = (m + \frac{1}{2})\sqrt{\rho n}. \tag{5.5}$$

- $D_Y := \deg_Y(Q)$ satisfies:

$$D_Y \leq \frac{D_X}{k} = \frac{m + \frac{1}{2}}{\sqrt{\rho}}. \quad (5.6)$$

- Each coefficient $Q_{ji}(Z)$ of $Q(X, Y)$ is in $\mathbb{F}_q[Z]$.
- $D_{YZ} := \deg_{Y,Z}(Q)$ (which is the total Y, Z degree of Q) satisfies:

$$D_{YZ} \leq \frac{(m + \frac{1}{2})^3}{6\sqrt{\rho}} n. \quad (5.7)$$

Only the bound on D_{YZ} needs to be discussed, and it is explained and proven in detail in Appendix B.

Claim 5.4 allows us to express the lower bound on $|S|$ from (5.3) in terms of D_X, D_Y, D_{YZ} , as

$$|S| > \frac{(1 + \frac{1}{2m})^7 m^7}{3\rho^{3/2}} n^2 = 2 \left(\frac{(m + \frac{1}{2})}{\sqrt{\rho}} \right)^3 ((m + \frac{1}{2})\sqrt{\rho}n) \left(\frac{(m + \frac{1}{2})^3}{6\sqrt{\rho}} n \right) > 2D_Y^3 D_X D_{YZ}. \quad (5.8)$$

5.2.2 Step 2: $Q(X, Y, z)$ is divisible by $Y - P_z(X)$ for many z

Recall that for every $z \in S$ there exists a polynomial $P_z(X) \in \mathbb{F}_q[X]$ of degree at most k with distance at most δ from $u_0 + zu_1$, i.e. $(x, P_z(x))$ equals $(x, w_z(x))$ for at least $n(1 - \delta)$ values of $x \in \mathcal{D}$. In each such point of agreement, the univariate polynomial $Q(X, P_z(X), z) \in \mathbb{F}_q[X]$ must then have a zero of order m , thus it has at least $mn(1 - \delta)$ roots counted with multiplicity. On the other hand, since Q is chosen to have $(1, k)$ -weighted degree less than D_X , and $\deg P_z(X) \leq k$, we have $\deg_X(Q(X, P_z(X), z)) < D_X$. Thus the polynomial must be identically zero if $D_X \leq mn(1 - \delta)$. Indeed this holds, because

$$1 - \delta \geq 1 - \delta_0 = \left(1 + \frac{1}{2m} \right) \sqrt{\rho} = \frac{D_X}{mn}.$$

Thus $Q(X, Y, z)$ is divisible by $Y - P_z(X)$ for each $z \in S$.

5.2.3 Step 3: Finding a good x_0 to start Hensel lifting

We now begin the process of finding a power series solution $Y = \gamma(X) \in \overline{\mathbb{K}}[[X]]$ to $Q(X, Y, Z) = 0$ (thought of as a bivariate equation $Q(X, Y) = 0$ with coefficients in \mathbb{K}). To find the power series solution, we will start at a suitable solution (x_0, α_0) of $Q(X, Y) = 0$, and then use Hensel lifting. In this section, our goal is to find such a “suitable” (x_0, α_0) .

Considering $Q(X, Y, Z)$ as a polynomial in Y over $\mathbb{F}_q[X, Z]$, it can be uniquely factored as

$$Q(X, Y, Z) = C(X, Z) \prod_i R_i(X, Y^{p^{f_i}}, Z)^{e_i}, \quad (5.9)$$

where p is the characteristic of \mathbb{F}_q , $f_i \geq 0$, $e_i \geq 1$, and each $R_i(X, Y, Z)$ is irreducible and separable³. In this section we prove the following claim:

³ $R(X, Y, Z)$ being separable in Y means it does not have repeated roots in the variable Y , in any extension field. This is equivalent to $\text{disc}_Y(R(X, Y, Z)) \neq 0$. For an irreducible polynomial in Y , it is also equivalent to the Y -derivative being not identically 0, or to the polynomial not being representable as a polynomial in Y^p .

Claim 5.5. *There exists $x_0 \in \mathbb{F}_q$ such that for all i ,*

$$\text{disc}_Y(R_i(X, Y, Z))(x_0) \neq 0 \in \mathbb{F}_q[Z].$$

Before we prove the claim, let us explain its use and motivation. The Hensel lift (described in more detail in Appendix A.4) shows that any *simple* root⁴ $Y = \alpha_0 \in \overline{\mathbb{K}}$ of $Q(x_0, Y, Z)$ can be uniquely lifted to a power series solution $\gamma(X) \in \overline{\mathbb{K}}[[X - x_0]]$ with free coefficient α_0 , by iteratively finding solutions to $Q(X, Y, Z) \equiv 0 \pmod{(X - x_0)^s}$ with increasing s .

However, it may be that $Q(X, Y)$ has no simple roots. This could happen, for example, if the factors of $Q(X, Y)$ appear with multiplicity. To resolve this, we instead focus on an irreducible factor R_i of Q . Even after focusing on R_i , it might still be the case that the particular root α_0 of $R_i(x_0, Y, Z)$ which we would wish to lift non-simple. We avoid this issue by requiring that that $R_i(x_0, Y, Z)$ is separable in Y (i.e., all of its roots in Y are simple). This happens if $\text{disc}_Y R_i(x_0, Y, Z) \neq 0$. Claim 5.5 exactly guarantees the existence of an element x_0 such that this occurs for all possible R_i . For all future sections, we will fix any such x_0 arbitrarily.

Henceforth, we will assume for simplicity that Q does not have inseparable irreducible factors, i.e. that $f_i = 0$ for all i . Note that since any inseparable factor has Y -degree at least p , this is necessarily the case if the characteristic is larger than D_Y , for example in the case where \mathbb{F}_q is a prime field with $p = q$. The general case, including the possibility of inseparable factors, is very similar, but has more technicalities, which are discussed in detail in Appendix C.

Proof of Claim 5.5. Since all of the polynomials $R_i(X, Y, Z)$ are separable in Y , the discriminants $\text{disc}_Y(R_i(X, Y, Z)) \in \mathbb{F}_q[X, Z]$ are non-zero polynomials. We need to find an $x_0 \in \mathbb{F}_q$ which makes all these discriminants evaluate to nonzero polynomials in $\mathbb{F}_q[Z]$. This will simply follow from a bound on the sum of degrees of all the R_i , which we would like to show is less than q . A crude bound on the sum of degrees (which would require a stronger bound on $|S|$) is easy to give. The rest of the proof is just a more careful bound.

Define

$$\text{disc}_Y^*(Q) := \prod_i \text{disc}_Y(R_i(X, Y, Z)) \in \mathbb{F}_q[Z][X].$$

It suffices to show that $\deg_X \text{disc}_Y^*(Q) < q$.

The discriminant $\text{disc}_Y(Q)$ is possibly the zero polynomial (if any $e_i > 1$), but formally it can be considered as a polynomial in the Y -coefficients $Q_j(X, Z) = \sum_{i < D_X - k \cdot j} Q_{ji}(Z)X^i$, and its degree can be computed from the Sylvester matrix regardless of whether it vanishes or not. This polynomial is formally divisible by $\text{disc}_Y^*(Q)$, and therefore we can bound

$$\begin{aligned} \deg_X(\text{disc}_Y^*(Q)) &\leq \deg_X(\text{disc}_Y(Q)) = (D_Y - 1)(2D_X - 2 - kD_Y) \\ &\leq \left(\left\lceil \frac{D_X}{k} \right\rceil - 2 \right) \left(2D_X - \left\lceil \frac{D_X}{k} \right\rceil \cdot k + k - 2 \right) \\ &\leq \frac{D_X(D_X - 1)}{k} < \frac{(m + \frac{1}{2})^2 \rho n^2}{\rho n} = (m + \frac{1}{2})^2 n. \end{aligned}$$

Note that from (5.3), we in particular have $q \geq |S| > \frac{(1 + \frac{1}{2m})^7 m^7}{3\rho^{3/2}} n^2 > (m + \frac{1}{2})^2 n$, and we have our desired inequality. \square

⁴ A root is simple when it has multiplicity 1.

5.2.4 Step 4: Focusing on a useful factor $R_i(X, Y, Z)$

Our main goal in the following sections is to show that at least one of the factors $R_i(X, Y, Z)$ is of the form $Y - P(X, Z)$, with $P \in \mathbb{F}_q[X, Z]$ a polynomial of X -degree at most k and Z -degree at most 1. Note that for such a factor, we will have in particular that for any $z \in \mathbb{F}_q$, $Y - P(X, z)$ is a factor of $R_i(X, Y, z)$, and $P(x_0, z)$ is a rational root of $R_i(x_0, Y, z)$. We will see that a converse is also true: if such factors and roots of R_i exist for sufficiently many z 's, then R_i must be divisible by (and thus of the form) $Y - P(X, Z)$. In this section we will make use of this fact in order to focus on a useful R_i . More precisely, we prove the following claim:

Claim 5.6. *There exists a factor $R = R_i$ of Q , and an irreducible factor $H(Y, Z)$ of $R(x_0, Y, Z)$, such that the set $S_{x_0, R, H}$ of z values for which both R and H vanish at P_z , i.e.*

$$S_{x_0, R, H} = \{z \in S : R(X, P_z(X), z) \equiv 0 \text{ and } H(P_z(x_0), z) = 0\},$$

is sufficiently large; more precisely, such that

$$|S_{x_0, R, H}| > 2D_Y^2 D_X D_Y Z. \quad (5.10)$$

Proof. After substituting $X \mapsto x_0$, each of the irreducible $R_i(X, Y, Z)$ can be factored as

$$R_i(x_0, Y, Z) = C_i(Z) \prod_j H_{ij}(Y, Z),$$

where $H_{ij} \in \mathbb{F}_q[Z][Y]$ are irreducible, separable in Y , and with positive Y -degree. For Q , this yields the factorization

$$Q(x_0, Y, Z) = \left(C(x_0, Z) \prod_i C_i(Z) \right) \prod_{i,j} H_{ij}(Y, Z).$$

In particular, the number of H_{ij} is at most D_Y .

For any $z \in S$, the polynomial $P_z(X)$ satisfies $Q(X, P_z(X), z) = 0$, i.e. $Y - P_z(X) \mid Q(X, Y, z)$, thus there is some i such that $Y - P_z(X) \mid R_i(X, Y, z)$, or equivalently $R_i(X, P_z(X), z) = 0$. Substituting $X \mapsto x_0$ yields also $Y - P_z(x_0) \mid R_i(x_0, Y, z)$, thus there is some j such that $Y - P_z(x_0) \mid H_{ij}(Y, z)$ and equivalently $H_{ij}(P_z(x_0), z) = 0$. Therefore we have $z \in S_{x_0, R_i, H_{ij}}$, by definition. Let (i, j) be the most common pair appearing in this process and set $R = R_i, H = H_{ij}$. Since the total number of pairs (i, j) is at most D_Y , and using (5.8), by the pigeonhole principle we find that

$$|S_{x_0, R, H}| \geq \frac{|S|}{D_Y} > 2D_Y^2 D_X D_Y Z,$$

as claimed. □

As mentioned above, our proof will *eventually* show that both R and H must in fact be linear in Y , with $R(X, Y, Z) = Y - P(X, Z)$, where P is linear in Z and of degree k in X and $H = R(x_0, Y, Z) = Y - P(x_0, Z)$. We will also see that $P(x_0, z) = P_z(x_0)$ for almost all $z \in S_{x_0, R, H}$. We will reach this point only later — for now we assume R, H have Y -degrees $d, d_H \leq D_Y$, total Y, Z degrees at most $D = D_Y Z$, and are not necessarily monic.

5.2.5 Step 5: Interlude — the algebraic function field \mathbb{L} and the power series $\gamma(X)$

Our next step is to find a root of H (if needed, by artificially adding it to the field $\mathbb{F}_q(Z)$), and then to lift it to a power series solution $Y = \gamma(X)$ to $R(X, Y, Z) = 0$. This process is carried out in Appendix A, which also provides the required setup and definitions from the theory of algebraic extensions of function fields.

We strongly encourage reading Appendix A at this point, as the analysis of the aforementioned γ in the next sections will use the following objects introduced and discussed there: the function field \mathbb{L} , the ring of regular functions \mathcal{O} , the special polynomials/algebraic functions ζ, ξ, W, \tilde{H} , and the power series γ itself, with its coefficients α_t and their numerators β_t . We will also use the definition of the rational substitution maps π_z (that allow us to substitute z into the regular algebraic functions in \mathcal{O}) and the weight function $\Lambda(\cdot)$ (that bounds the number of zeroes of a regular algebraic function) from that appendix, as well as Lemma A.1.

At the end of the day, the power series $\gamma(X)$ will be shown to be of the form $P(X, Z) \in \mathbb{F}_q[X, Z]$ with X - and Z -degrees at most k and 1, respectively. However, to reach that point, we will analyze $\gamma(X)$ as having coefficients in an algebraic extension of $\mathbb{F}_q(Z)$, and unbounded X -degree.

5.2.6 Step 6: Bounding the X -degree of γ

In this section we show that the power series solution $\gamma = \sum_{t=0}^{\infty} \alpha_t (X - x_0)^t$ to $R(X, Y, Z) = 0$ is in fact a finite polynomial in X of degree k . In other words, we prove:

Claim 5.7. *For all $t > k$, $\alpha_t = 0$. Equivalently,*

$$\gamma = \gamma_k = \sum_{t=0}^k \alpha_t (X - x_0)^t.$$

The claim is proved in two steps: first for all $k < t < D_X$, by showing that $\pi_z(\alpha_t)$ is well defined and vanishes for sufficiently many substitutions; then for all $t \geq D_X$ as well, by observing that γ_k is already a root of $R(X, Y, Z)$.

Proof. For each $z \in S_{x_0, R, H}$, from $H(P_z(x_0), z) = 0$ we also get $\tilde{H}(W(z)P_z(x_0), z) = 0$, and thus we have a substitution map $\pi_z : \mathcal{O} \rightarrow \mathbb{F}_q$ with $t_z = W(z)P_z(x_0)$. The denominators appearing in the α_t are all powers of W and ξ , so we can evaluate α_t at any z which is not a root of W or ξ , i.e. in the set

$$S' = S_{x_0, R, H} \setminus \{z : W(z) = 0 \text{ or } \pi_z(\xi) = 0\}.$$

Using lemma A.1 and the bounds on $\deg W, \Lambda(\xi)$ from claim A.2, as well as (5.10), we find

$$\begin{aligned} |S'| &\geq |S_{x_0, R, H}| - (\deg W + d_H \Lambda(\xi)) \\ &\geq |S_{x_0, R, H}| - ((d-1)d_H + 1)(D - d_H + 1) + 1 \\ &> |S_{x_0, R, H}| - d_H d D > 2D_Y^2 D_X D_{YZ} - D_Y^2 D_{YZ} \\ &= D_Y^2 D_{YZ} (2D_X - 1). \end{aligned} \tag{5.11}$$

Note that to apply the lemma we also used $\xi \neq 0$ in \mathbb{L} .

Now for each $z \in S'$, we may apply π_z to γ , since we removed all poles of γ from S' , and find $R(X, \pi_z(\gamma), z) = \pi_z(R(X, \gamma, Z)) = \pi_z(0) = 0$. On the other hand, note that $\pi_z(\alpha_0) = \pi_z(T/W) = t_z/W(z) = P_z(x_0)$, that $R(X, P_z(X), z) = 0$, and that

$$\pi_z(\zeta) = \pi_z\left(\frac{\partial R}{\partial Y}\left(x_0, \frac{T}{W}, Z\right)\right) = \frac{\partial R}{\partial Y}\left(x_0, \frac{t_z}{W(z)}, z\right) = \frac{\partial R}{\partial Y}(x_0, P_z(x_0), z)$$

is well-defined and non-zero. We've seen that $\pi_z(\gamma), P_z(X) \in \mathbb{F}_q[[X - x_0]]$ are power series solutions to $R(X, Y, z) = 0$, with the same free coefficient $\pi_z(\alpha_0) = P_z(x_0)$, which is a simple root of

$R(x_0, Y, z)$. Thus, they must be identically equal, by the uniqueness of the Hensel lift with a given starting simple root. In other words, we have

$$\mathbb{F}_q[[X - x_0]] \ni \sum_{t=0}^{\infty} \pi_z(\alpha_t)(X - x_0)^t = \pi_z(\gamma) = P_z(X) \in \mathbb{F}_q[X - x_0]$$

and in particular, $\pi_z(\alpha_t) = 0$ for all $z \in S'$ and all $t > k$, since $\deg(P_z) \leq k$. Thus we also find $\pi_z(\beta_t) = 0$ for all $k < t < D_X$. Additionally for such t 's we have by claim A.2

$$\Lambda(\beta_t) < (2t + 1)dD \leq dD(2D_X - 1),$$

and from (5.11) it follows that

$$|S'| > D_Y^2 D_{YZ}(2D_X - 1) \geq d_H dD(2D_X - 1) > d_H \Lambda(\beta_t).$$

We can therefore apply lemma A.1 to find that indeed $\beta_t = 0$ and $\alpha_t = 0$ in \mathbb{L} .

We thus have that the degree k polynomial

$$\gamma_k = \sum_{t=0}^k \alpha_t (X - x_0)^t = \sum_{t=0}^{D_X-1} \alpha_t (X - x_0)^t = \gamma_{D_X-1} \in \mathbb{L}[X]$$

satisfies $\gamma \equiv \gamma_k \pmod{(X - x_0)^{D_X}}$, and therefore

$$R(X, \gamma_k, Z) \equiv 0 \pmod{(X - x_0)^{D_X}};$$

but, $R(X, \gamma_k, Z) \in \mathbb{L}[X]$ is a polynomial of degree $< D_X$, since by construction Q has $(1, k)$ -weighted degree less than D_X and so do its factors, and therefore $R(X, \gamma_k, Z) = 0$ identically. By the uniqueness of the lifting, we thus find $\gamma = \gamma_k \in \mathbb{L}[X]$, as claimed. \square

5.2.7 Step 7: Bounding the Z -degree of γ

In the previous section we've seen that $\gamma = \gamma_k \in \mathbb{L}[X]$ is a polynomial of degree at most k in X , whose coefficients lie in \mathbb{L} , an extension field of $\mathbb{F}_q(Z)$. We've also seen that $\pi_z(\gamma) = P_z(X)$ for all $z \in S'$. In this section we will show that the coefficients of γ are all in fact simply linear polynomials in Z , and thus obtain that $\gamma = P(X, Z) \in \mathbb{F}_q[X, Z]$ with X -degree at most k and Z -degree at most 1:

Claim 5.8. *The exists degree $\leq k$ polynomials $v_0, v_1 \in \mathbb{F}_q[X]$, such that*

$$\gamma = v_0(X) + Z \cdot v_1(X) =: P(X, Z).$$

We do this by showing that $\gamma(x)$ agrees with the linear function $w(x, Z)$ on at least $k + 1$ values of x , using sufficiently many Z -substitutions at each x , and then use the fact γ can be interpolated from such values of $\gamma(x)$, and this interpolation is also linear in Z . Details follow.

We consider good pairs of $x \in \mathcal{D}, z \in S'$ satisfying $w(x, z) = P_z(x)$. We define the sets of x 's which are good for each $z \in S'$ and vice versa, that is,

$$\mathcal{D}_z = \{x \in \mathcal{D} : w(x, z) = P_z(x)\},$$

$$S'_x = \{z \in S' : w(x, z) = P_z(x)\} = \{z \in S' : x \in \mathcal{D}_z\}.$$

By the definitions of S and P_z , we have $|\mathcal{D}_z| \geq n - e$ for each $z \in S'$, where $e = \lfloor \delta n \rfloor$. We make the following claims regarding the sizes of the sets S'_x :

Claim 5.9. *Suppose $|S'_x| > (2k + 1)d_H dD$. Then $\gamma(x) = w(x, Z)$, and in particular, $\gamma(x)$ is a linear polynomial in $\mathbb{F}_q[Z]$.*

Claim 5.10. *There exists a set $\mathcal{D}_{\text{top}} = \{x_1, \dots, x_{k+1}\} \subset \mathcal{D}$ of $k+1$ points of \mathcal{D} , satisfying $|S'_{x_j}| > (2k+1)d_H dD$ for all $1 \leq j \leq k+1$.*

Before proving the two claims, let us first deduce Claim 5.8 from them:

Proof Claim 5.8. Observe that from Claim 5.9 and Claim 5.10 it follows that $\gamma(x) = w(x, Z) = u_0(x) + Z \cdot u_1(Z)$ is linear in Z for every $x = x_j \in \mathcal{D}_{\text{top}}$. But, since $\gamma(X)$ is a polynomial of degree at most k , it can be interpolated from its values in any $k+1$ points, and this interpolation only involves operations over \mathbb{F}_q . Thus the interpolated polynomial will also have coefficients which are linear in Z .

More concretely, let $v_0(X), v_1(X) \in \mathbb{F}_q[X]$ be the unique polynomials of degree at most k interpolating $u_0(x), u_1(x)$ at the points of \mathcal{D}_{top} . Then $\gamma(X)$ and $v_0(X) + Z \cdot v_1(X)$ are two polynomials in $\mathbb{L}[X]$ of degree at most k which agree on at least $k+1$ evaluations, since

$$\gamma(x_j) = w(x_j, Z) = u_0(x_j) + Z \cdot u_1(x_j) = (v_0(X) + Z \cdot v_1(X))(x_j)$$

for each $x_j \in \mathcal{D}_{\text{top}}$. It follows that γ and $v_0 + Z \cdot v_1$ are identically equal as polynomials in $\mathbb{L}[X]$, as claimed. \square

We now proceed to prove the claims:

Proof of Claim 5.9. Since $\pi_z(\gamma) = P_z(X)$ for each $z \in S'$, by definition of S'_x we have

$$\pi_z(\gamma(x)) = P_z(x) = w(x, z) = u_0(x) + z \cdot u_1(x)$$

for each $z \in S'_x$, or equivalently

$$\pi_z(\gamma(x) - (u_0(x) + Z \cdot u_1(x))) = 0. \quad (5.12)$$

On the other hand, we can write

$$\begin{aligned} \gamma(x) - (u_0(x) + Z \cdot u_1(x)) &= \left(\frac{1}{W^{k+1} \xi^{e_k}} \sum_{t=0}^k \beta_t (x - x_0)^t W^{k-t} \xi^{e_k - e_t} \right) - (u_0(x) + Z \cdot u_1(x)) \\ &= \frac{1}{W^{k+1} \xi^{e_k}} \left(\beta(x) - (u_0(x) + u_1(x) \cdot Z) W^{k+1} \xi^{e_k} \right), \end{aligned} \quad (5.13)$$

where $\beta(x) := \sum_{t=0}^k \beta_t (x - x_0)^t W^{k-t} \xi^{e_k - e_t} \in \mathcal{O}$, which by Claim A.2 has weight

$$\Lambda(\beta(x)) = 1 + (k+1)\Lambda(W) + e_k \Lambda(\xi) \leq (2k+1)dD,$$

and so does $\tilde{\beta}(x) := \beta(x) - (u_0(x) + u_1(x) \cdot Z) W^{k+1} \xi^{e_k}$. From (5.12) and (5.13) we have $\pi_z(\tilde{\beta}(x)) = 0$ for all $z \in S'_x$, with $|S'_x| > (2k+1)d_H dD \geq d_H \Lambda(\tilde{\beta}(x))$, by assumption. By Lemma A.1 it follows that $\tilde{\beta}(x) = 0$, and thus $\gamma(x) = u_0(x) + Z \cdot u_1(x) = w(x, Z)$ identically in \mathbb{L} , as claimed. \square

Proof of Claim 5.10. Let $\mathcal{D}_{\text{top}} = \{x_1, \dots, x_{k+1}\} \subset \mathcal{D}$ be the set of the $x_j \in \mathcal{D}$ with the $k+1$ largest sizes of $|S'_{x_j}|$, breaking ties arbitrarily. We first claim that for each $1 \leq j \leq k+1$,

$$|S'_{x_j}| \geq \frac{n-k-1-e}{n-k-1} |S'| \geq \frac{1-\rho-\delta}{1-\rho} |S'|.$$

This follows by way of contradiction, for otherwise we would have:

$$\begin{aligned} e|S'| &\geq \sum_{z \in S'} |\mathcal{D} \setminus \mathcal{D}_z| = \sum_{x \in \mathcal{D}} |S' \setminus S'_x| \geq \sum_{x \in \mathcal{D} \setminus \mathcal{D}_{\text{top}}} (|S'| - |S'_x|) \\ &> (n-k-1) \left(1 - \frac{n-k-1-e}{n-k-1} \right) |S'| = e|S'|. \end{aligned}$$

From (5.11) we have

$$|S'_{x_j}| \geq \frac{1-\rho-\delta}{1-\rho} |S'| > \frac{1-\rho-\delta}{1-\rho} D_Y^2 D_{YZ} (2D_X - 1) = \frac{1-\rho-\delta}{1-\rho} (2D_X - 1) d_H d D,$$

so to conclude $|S'_{x_j}| > (2k+1)d_H d D$ it suffices to show that $2D_X - 1 > \frac{1-\rho}{1-\rho-\delta} (2k+1)$. And indeed for $m \geq 3$, we have

$$\begin{aligned} 2D_X - 1 &= (2m+1)\sqrt{\rho}n - 1 > 2 \cdot 3\sqrt{\rho}n - 1 > \frac{1+\sqrt{\rho}}{\sqrt{\rho}}(3\rho n) - 1 \\ &> \frac{1-\rho}{\sqrt{\rho}(1-\sqrt{\rho})}(2k+1) > \frac{1-\rho}{\sqrt{\rho} + \frac{\sqrt{\rho}}{2m} - \rho}(2k+1) \\ &= \frac{1-\rho}{1-\delta_0(\rho, m) - \rho}(2k+1) \geq \frac{1-\rho}{1-\rho-\delta}(2k+1) \end{aligned}$$

as needed. \square

5.2.8 Step 8: Proving the correlated agreement between u_i and v_i

We've found a polynomial $\gamma = v_0(X) + Z \cdot v_1(X)$ of the required degrees satisfying $Q(X, \gamma, Z) = 0$. To finish the proof of Theorem 5.1, it now remains only to be seen that $\gamma(x, Z)$ and $w(x, Z)$ agree identically on all but $\delta|\mathcal{D}|$ values of $x \in \mathcal{D}$.

Define $\mathcal{D}' = \{x \in \mathcal{D} : |S'_x| \geq 2\}$. Note that for each $x \in \mathcal{D}'$, we must have $v_0(x) = u_0(x)$, $v_1(x) = u_1(x)$ (and thus $\gamma(x, Z) = w(x, Z)$), since

$$\begin{aligned} v_0(x) + z \cdot v_1(x) &= \pi_z(\gamma(x)) = \pi_z(\gamma)(x) = P_z(x) = w(x, z) \\ &= u_0(x) + z \cdot u_1(x) \end{aligned}$$

for at least 2 different values of $z \in S'_x$. Note that, from (5.3), (5.6), (5.7) as well as $m \geq 2$ and $\rho < 1$ we have

$$|S'| \geq \frac{|S|}{D_Y} - D_Y^2 D_{YZ} > \frac{|S|}{2D_Y} > \frac{(m + \frac{1}{2})^6}{6\rho} n^2 > 2n^2,$$

and therefore the size of $\mathcal{D} \setminus \mathcal{D}'$, i.e. the number of $x \in \mathcal{D}$ for which $|S'_x| \leq 1$, is at most

$$\frac{e|S'|}{|S'|-1} < \left(1 + \frac{2}{|S'|}\right)e < \left(1 + \frac{1}{n^2}\right)e < e + \frac{1}{n} < e + 1,$$

from which we find

$$|\{x \in \mathcal{D} : (u_0(x), u_1(x)) = (v_0(x), v_1(x))\}| \geq |\mathcal{D}'| \geq (1-\delta)|\mathcal{D}|$$

as we wanted to show. \square

Remark 5.11. Note that from $\gamma = v_0(X) + Z \cdot v_1(X) = P(X, Z)$ solving $R(X, \gamma, Z) = 0$ we find that $Y - P(X, Z) \mid R(X, Y, Z)$. But since R is irreducible, it indeed follows that $R = Y - P(X, Z)$ is monic and linear in Y , and so is H , as was mentioned earlier; and, as mentioned, we only reach this conclusion near the end of the proof. It is an interesting open problem whether this conclusion can be reached without passing through the various extension fields as our proof required.

6 Correlated Agreement in Generalized Settings

Both Theorems 4.1 and 5.1 considered two functions $u_0, u_1 : \mathcal{D} \rightarrow \mathbb{F}$ and the corresponding affine line $\{u_0 + zu_1 : z \in \mathbb{F}_q\}$ inside the linear plane $\text{span}\{u_0, u_1\}$. A generalization of these theorems that is particularly important to the soundness analysis of the FRI protocol in Theorem 7.2 is obtained by considering $l + 1$ functions $u_0, \dots, u_l : \mathcal{D} \rightarrow \mathbb{F}$, and the 1-dimensional, degree l parameterized curve

$$\{u_0 + zu_1 + z^2u_2 + \dots + z^lu_l : z \in \mathbb{F}_q\}$$

inside the linear space $\text{span}\{u_0, \dots, u_l\}$. This curve can also be viewed as a function $w(\cdot, Z) : \mathcal{D} \rightarrow \mathbb{K}$ given by

$$w(x, Z) = u_0(x) + u_1(x)Z + u_2(x)Z^2 + \dots + u_l(x)Z^l.$$

The two theorems can then be generalized as follows:

Theorem 6.1. *Suppose $\delta \leq (1 - \rho)/2$. Let $u_0, u_1, \dots, u_l : \mathcal{D} \rightarrow \mathbb{F}_q$ be functions. Let*

$$S = \{z \in \mathbb{F}_q : \Delta(u_0 + zu_1 + \dots + z^lu_l, V) \leq \delta\}$$

and suppose $|S| > l \cdot n$. Then for all $z \in \mathbb{F}_q$ we have

$$\Delta(u_0 + zu_1 + \dots + z^lu_l, V) \leq \delta,$$

and furthermore there are $v_0, \dots, v_l \in V$ such that for all $z \in \mathbb{F}_q$,

$$\Delta(u_0 + zu_1 + \dots + z^lu_l, v_0 + zv_1 + \dots + z^lv_l) \leq \delta$$

and in fact

$$|\{x \in \mathcal{D} : (u_0(x), \dots, u_l(x)) \neq (v_0(x), \dots, v_l(x))\}| \leq \delta|\mathcal{D}|.$$

Theorem 6.2. *Let $u_0, u_1, \dots, u_{l+1} : \mathcal{D} \rightarrow \mathbb{F}_q$, let $m \geq 3$, define $\delta_0(\rho, m) := 1 - \sqrt{\rho} - \frac{\sqrt{\rho}}{2m}$, and let $\delta \leq \delta_0(\rho, m)$. Define*

$$S = \{z \in \mathbb{F}_q : \Delta(u_0 + zu_1 + \dots + z^lu_l, V) \leq \delta\}$$

and suppose

$$|S| > \frac{(1 + \frac{1}{2m})^7 m^7}{3\rho^{3/2}} n^2 l. \tag{6.1}$$

Then u_0, \dots, u_l are simultaneously δ -close to V , i.e. $\exists v_0, \dots, v_l \in V$ such that

$$|\{x \in \mathcal{D} : \forall 0 \leq i \leq l, u_i(x) = v_i(x)\}| \geq (1 - \delta)|\mathcal{D}|.$$

These generalizations do not greatly affect the proofs, which were presented in the previous sections in the special case $l = 1$ only for the purposes of simplicity. We will thus not repeat the arguments in full, but only detail the required changes in subsections 6.1 and 6.2.

Another generalization of interest is to correlated agreement in the entire affine space, stated as Theorem 1.7, restated here for completeness:

Theorem 1.7 (Correlated agreement over affine spaces). *Let V, q, n, k and ρ be as defined in Theorem 1.2. For $u_0, u_1, \dots, u_l \in \mathbb{F}_q^{\mathcal{D}}$ let $U = u_0 + \text{span}\{u_1, \dots, u_l\} \subset \mathbb{F}_q^{\mathcal{D}}$ be an affine subspace. If $\delta \in (0, 1 - \sqrt{\rho})$ and*

$$\Pr_{u \in U} [\Delta(u, V) \leq \delta] > \epsilon,$$

where ϵ is as defined in Theorem 1.2, then there exists $\mathcal{D}' \subset \mathcal{D}$ and $v_0, \dots, v_l \in V$ satisfying

- **Density:** $|\mathcal{D}'|/|\mathcal{D}| \geq 1 - \delta$, and

- **Agreement:** for all $i \in \{0, \dots, l\}$, the functions u_i and v_i agree on all of \mathcal{D}' .

Furthermore, in the unique decoding regime $\delta \in \left(0, \frac{1-\rho}{2}\right]$, there exists a unique maximal \mathcal{D}' satisfying the above, with unique v_i .

To prove this theorem we will make use of the following lemma:

Lemma 6.3. *Let V, q, n, k and ρ be as defined in Theorem 1.2. For $u_0, u_1, \dots, u_l \in \mathbb{F}_q^{\mathcal{D}}$ let $U = u_0 + \text{span}\{u_1, \dots, u_l\} \subset \mathbb{F}_q^{\mathcal{D}}$ be an affine subspace, and let $U' = \text{span}\{u_1, \dots, u_l\}$ be the corresponding linear subspace (so that $U = u_0 + U'$ and $U' = U - U$). If $\delta \in (0, 1 - \sqrt{\rho})$ and*

$$\Pr_{u \in U} [\Delta(u, V) \leq \delta] > \epsilon,$$

where ϵ is as defined in Theorem 1.2, then we have $\Delta(u', V) \leq \delta$ for every $u' \in U'$.

Note that Lemma 6.3 is very similar to Theorem 1.2, the difference being only that the consequent $\Delta(u', V) \leq \delta$ is stated for elements of the linear space U' instead of elements of the affine space U . In particular, when U itself is linear and $U' = U$, they are equivalent. Moreover, Theorem 1.2 can be proven from Lemma 6.3, and in fact we will pass through such a proof in the course of proving Theorem 1.7. Both Lemma 6.3 and Theorem 1.7 are proven in Section 6.3.

6.1 Proof of Theorem 6.1

Following the proof in Section 4, the first change is in the analysis of the matrix M , where the entries in the columns corresponding to the coefficients of A are of degree $\leq l$ in Z instead of 1 (and the columns corresponding to the coefficients of B remain of degree 0). The determinants of minors $R(Z)$ are then of degree at most $l(e+1)$, which is less than $|S|$, so are still identically 0, and the matrix has rank $< k + 2e + 2$.

Similarly, the solution A, B of the equation system will have Z -degrees $\deg_Z(A) \leq le$, $\deg_Z(B) \leq l(e+1)$. To apply the Polishchuk–Spielman lemma, the necessary inequalities follow from bounds on the X -degrees which are unchanged, as well as the bounds

$$\frac{\deg_Z(A)}{|S|} < \frac{e}{n}, \quad \frac{\deg_Z(B)}{|S|} < \frac{e+1}{n}$$

which are still valid; in both fractions, the bounds on both numerator and denominator are simply multiplied by l .

The ratio $Q = \frac{B}{A}$ and its coefficients will have Z -degree at most $l(e+1)$, and the set $S' = \{z \in S : A(X, z) \neq 0\}$ will be of size at least $|S'| \geq |S| - \deg_Z(A) > l(n - e) \geq l(e+1)$, from which it follows that $\deg_X(Q) \leq k$.

The definition of \mathcal{D}' , the bound on its size, and the choice of $\{x_0, \dots, x_k\}$ remain the same. We again define v_0, \dots, v_l as the minimal interpolation of u_0, \dots, u_l at the x_i , and we find that the $\mathbb{K}[X]$ polynomials $Q(X, Z)$ and $\sum_{i=0}^l v_i(X)Z^i$ agree on $k+1$ points and thus must be identical. It then follows that $u_i(x) = v_i(x)$ for all $0 \leq i \leq l$ and $x \in \mathcal{D}'$, with $|\mathcal{D} \setminus \mathcal{D}'| \leq \delta|\mathcal{D}|$, as claimed. \square

6.2 Proof of Theorem 6.2

As in $l = 1$ case, we apply the Guruswami–Sudan decoder to the $\mathbb{K}^{\mathcal{D}}$ word $w(x, Z) = \sum_{i=0}^l u_i(x)Z^i$. We use exactly the same parameters, and the degrees D_X, D_Y are unaffected. In the entries of the equation system, every appearance of (a power of) the previously linear $w(x, Z)$ is replaced by (a

power of) a polynomial of degree l , thus the bounds on D_C, D_R, D, D_{YZ} are all multiplied by l , and specifically $D_{YZ} < \frac{(m+1/2)^3}{6\sqrt{\rho}} n l$.

Also note that instead of the Y, Z degrees being graded in Q such that the coefficient $Q_{ji}(Z)$ of $X^i Y^j$ is of degree at most $D - j$ in Z , it is of degree at most $D - l \cdot j$, i.e. when Y and Z are assigned weights l and 1 (instead of 1 and 1), the total weight of Q , and thus also of R and H , is at most D . Thus all bounds on Z -degrees and weights are henceforth simply multiplied by l . The leading coefficient W of H has degree at most $D - l \cdot d$, and the variable $T = W(Z)Y$ is assigned the l times larger weight $D - l(d - 1)$, which makes the weight of the monic polynomial \tilde{H} dominated by its leading monomial T^d , thus the weight never increases on reductions modulo \tilde{H} . The bounds on weights of ξ, β_t are multiplied by l . Since for $\mathbb{F}_q[Z]$ polynomials the weight corresponds to degree, lemma A.1 remains valid; in all its applications, the upper bounds on weights of the regular elements are multiplied by l ; and so are the lower bounds on the sizes of the sets of vanishing substitutions, i.e. the sets $S, S_{x_0, R, H}, S', S'_{x_j}$. Thus we can still deduce that γ is of degree k , and then by interpolating u_0, \dots, u_l by polynomials v_0, \dots, v_l at $k + 1$ appropriately chosen points, to deduce that $\gamma = \sum_{i=0}^l v_i(X)Z^i$ identically.

For the final argument we again note that $\sum_{i=0}^l u_i(x)z^i = \sum_{i=0}^l v_i(x)z^i$ for all $x \in \mathcal{D}$ and $z \in S'_x$. Since for a fixed x both sides of the equation are degree l polynomials in z , if $|S'_x| > l$ then we must have $u_i(x) = v_i(x)$ for all $0 \leq i \leq l$. We have $|S'| > 2l \cdot n^2$, or equivalently $\frac{l}{|S'|} < \frac{2}{n^2}$, and thus the number of $x \in \mathcal{D}$ for which $|S'_x| \leq l$ is at most $\frac{e|S'|}{|S'| - l} < e + 1$, and again we find that u_i and v_i all agree on the set $\mathcal{D}' = \{x \in \mathcal{D} : |S'_x| > l\}$ which is of size at least $(1 - \delta)n$, as claimed. \square

6.3 Proofs of Lemma 6.3 and Theorem 1.7

Proof of Lemma 6.3. Let $u' \in U'$ be an arbitrary element. If $u' = 0$, then clearly $\Delta(u', V) = 0 \leq \delta$. Otherwise, consider the partition of U into affine lines parallel to u' ; formally, write $U = \tilde{U} \oplus \text{span}\{u'\}$, where $\tilde{U} \subset U$ is some direct complement of u' in U' shifted by u_0 . Then

$$\mathbb{E}_{\tilde{u} \in \tilde{U}} \Pr_{z \in \mathbb{F}_q} [\Delta(\tilde{u} + z \cdot u', V) \leq \delta] = \Pr_{u \in U} [\Delta(u, V) \leq \delta] > \epsilon,$$

and in particular there exists some $\tilde{u} \in \tilde{U}$ for which $\Pr_{z \in \mathbb{F}_q} [\Delta(\tilde{u} + z \cdot u', V) \leq \delta] > \epsilon$. We can thus apply Theorem 1.4 to the line $\{\tilde{u} + z \cdot u' : z \in \mathbb{F}_q\}$, which in particular implies $\Delta(u', V) \leq \delta$, as claimed. \square

Proof of Theorem 1.7. Let $\delta^* = D(U, V)$, where D is the divergence as defined in subsection 1.2, and let $u^* \in U$ be an element with $\Delta(u^*, V) = \delta^*$.

We first show that $\delta^* \leq \delta$. Let $\bar{U} = \text{span}(U)$ be the linear space spanned by U , which either equals U when $0 \in U$, or otherwise is the disjoint union of $U' = U - U$ and $\{z \cdot U : z \in \mathbb{F}_q \setminus \{0\}\}$. By Lemma 6.3 we have $\Delta(u', V) \leq \delta$ for all $u' \in U'$, and we also have $\Delta(z \cdot u, V) = \Delta(u, V)$ for all $z \neq 0$ and $u \in U$. Thus in both cases, whether $\bar{U} = U$ or $\bar{U} = U' \cup \bigcup_{z \neq 0} (z \cdot U)$, we have

$$\Pr_{\bar{u} \in \bar{U}} [\Delta(\bar{u}, V) \leq \delta] \geq \Pr_{u \in U} [\Delta(u, V) \leq \delta] > \epsilon.$$

We can therefore apply Lemma 6.3 to \bar{U} , and in particular for $u^* \in U \subset \bar{U}$ we get $\delta^* = \Delta(u^*, V) \leq \delta$, as claimed. As previously noted, this is in fact the content of Theorem 1.2, which is now proved.

Let $\{v_1^*, \dots, v_L^*\} \subset V$ be all possible codewords at distance at most δ^* from u^* . Note that they are all in fact at distance *exactly* δ^* , not less, since otherwise we would have $\Delta(u^*, V) < \delta^*$. For each $1 \leq i \leq L$ define the agreement set $\mathcal{D}_i^* = \{x \in \mathcal{D} : u^*(x) = v_i^*(x)\}$ which has size exactly

$(1 - \delta^*)|\mathcal{D}|$. Note that in the unique decoding regime $\delta^* \leq \frac{1-\rho}{2}$ we must have $L = 1$, and in the general Johnson/Guruswami–Sudan regime $\delta^* < 1 - \sqrt{\rho} - \eta$ we have $L < q$: indeed, from $\epsilon = \epsilon_J < 1$ and $\eta < \frac{\sqrt{\rho}}{16}$ we must have $q > (2\eta)^{-7} > \frac{2^{17}}{\eta}, \frac{2^{21}}{\sqrt{\rho}}$, and from the analysis of the Guruswami–Sudan algorithm, using (5.6) and $m = \left\lceil \frac{\sqrt{\rho}}{2\eta} \right\rceil < \frac{\sqrt{\rho}}{2\eta} + 1$ we find

$$L \leq D_Y < \frac{2m+1}{2\sqrt{\rho}} < \frac{1}{2\eta} + \frac{3}{2\sqrt{\rho}} < 2^{-17}q < q.$$

For each $i \leq L$, let $U_i = \{u \in U : u|_{\mathcal{D}'_i} \in V|_{\mathcal{D}'_i}\}$ be the set of all functions in U which agree with some codeword in V on all of \mathcal{D}'_i . Note that this condition is linear, therefore each U_i is an affine subspace of U . We claim that $U \subset \bigcup_{i=0}^L U_i$, i.e. every element of U belongs to at least one U_i . This is obvious for u^* , which belongs to all U_i . Consider any $u \in U \setminus \{u^*\}$, and the affine line $\{u^* + z \cdot (u - u^*) : z \in \mathbb{F}_q\} \subset U$ containing u^* and u . By definition of δ^* , we have

$$\Pr_{z \in \mathbb{F}_q} [\Delta(u^* + z \cdot (u - u^*), V) \leq \delta^*] = 1 > \epsilon.$$

Since $\delta^* \leq \delta$, by Theorem 1.4 we have *correlated* agreement in the line, i.e. there exists $\mathcal{D}' \subset \mathcal{D}$ with $|\mathcal{D}'| \geq (1 - \delta^*)|\mathcal{D}|$, and codewords $v^*, v \in V$ which respectively agree with u^*, u on \mathcal{D}' . In particular $\Delta(u^*, v^*) \leq \delta^*$, thus v^* must be one of the L decodings $\{v_1^*, \dots, v_L^*\}$, as the list was exhaustive, and $\mathcal{D}' = \mathcal{D}'_i$. This exactly implies $u \in U_i \subset \bigcup_{i=0}^L U_i$, as claimed.

Comparing sizes, we find that the largest U_i must then satisfy $|U_i| \geq \frac{|U|}{L} > \frac{|U|}{q}$. But U_i is a subspace of U , and it thus follows that $U_i = U$. Now setting $\mathcal{D}' = \mathcal{D}'_i$, we have in particular that the restrictions of u_0, u_1, \dots, u_l to \mathcal{D}' are codewords, since they are elements or differences of elements in U_i . Then setting $v_0, v_1, \dots, v_l \in V$ to be the unique extensions of these codewords from \mathcal{D}' to \mathcal{D} , we find that $\mathcal{D}', v_0, \dots, v_l$ satisfy both conditions. Again note that in the unique decoding regime, $L = 1$ and this \mathcal{D}' is therefore uniquely defined. \square

7 Applications to Verifiable Secret Sharing and FRI soundness

In this section we give some more details on applications that were briefly described in Section 3.

7.1 Verifiable Secret Sharing

We start with an application of distributed proximity testing of Interleaved Reed–Solomon Codes (Theorem 3.1) to *verifiable secret sharing* (VSS) [CGMA85]. VSS serves as a building block both for useful special tasks, such as simultaneous broadcast and fair coin-flipping, and for general protocols for secure multiparty computation with an honest majority [BGW88, CCD88, RB89].

A VSS scheme with security threshold t is a two-phase protocol involving a dealer, n servers, and an honest output client. The *sharing phase* starts with the dealer distributing the secret s among the servers by sending a share s^i to server i and is followed by a verification protocol. In the *reconstruction phase*, each server sends its share s^i to the output client, who reconstructs s from the shares. Both phases of the protocol are attacked by a malicious adversary who may corrupt at most t servers and possibly also the dealer. Communication proceeds in synchronous rounds and may use secure point-to-point channels. We also assume the availability of a common broadcast medium and a source of unpredictable public coins. These assumptions, which can be eliminated at a small amortized cost, make the protocol simpler. Finally, while the communication is synchronous, the

adversary has a *rushing* capability, in the sense that it can wait to receive messages from uncorrupted parties before sending its own messages.

Here we consider a simplified variant of VSS that allows a denial-of-service attack only in the sharing phase, but not in the reconstruction phase. The latter “guaranteed output delivery” requirement makes the protocol suitable for applications that rely on independence, such as simultaneous broadcast and fair coin-tossing, as they prevent the adversary from making the protocol selectively fail based on information obtained in the final round via rushing.

A VSS protocol as above should satisfy the following properties:

- **Completeness:** if the dealer is honest and the sharing phase succeeds, the output client outputs s . Moreover, the sharing phase succeeds if the adversary does not attack it.
- **Secrecy:** if the dealer is honest, the adversary learns nothing about s .
- **Unique reconstruction:** even if the dealer is dishonest, if the sharing phase succeeds then the messages sent in the sharing phase define a unique s^* such that (except with small failure probability) the output client will output s^* in the end of the reconstruction phase.

Implementing VSS efficiently is a challenging task. Here we consider the following simple approach for simultaneously sharing l secrets s_1, \dots, s_l . This approach underlies the scalable MPC protocols from [DI06, IPS09]. A centralized variant of this protocol (in a relaxed setting that allows reconstruction to fail) is a building block in efficient two-party protocols for zero knowledge [IKOS09, AHIV17] and secure computation [IPS08, HIMV19].

We will assume here that each secret s_j is a single field element and all secrets originate from the same dealer. However, the protocol is even more attractive when the secrets originate from different dealers and when each secret s_j is a vector of $\ell < n$ field elements that are simultaneously shared via so-called “packed secret sharing” [FY92]. Moreover, while the communication complexity of the protocol beats all competing approaches we are aware of in the *amortized* case when l is large, it is potentially useful (and nontrivial to analyze) even with $l = 1$. In fact, if we count public coins and broadcast as normal communication, the communication complexity of the protocol is nontrivial to match even in this case.

The two phases of the protocol proceed as follows.

Sharing. In the sharing phase, the dealer uses Shamir’s secret sharing scheme for sharing each secret, with secrecy threshold t . Viewed abstractly, each s_j is randomly mapped to a codeword u_j of an $[n, t + 1, d]$ RS code V over \mathbb{F}_q , with $d = n - t$. The dealer distributes u_j between the servers, together with an additional random codeword $u_0 \in_R V$ that is used for blinding. Following a public random challenge $r \in \mathbb{F}_q^l$, each server i broadcasts its view of $u' = r^T \mathbf{u}$, namely $u'(i) = u_0(i) + \sum_{j=1}^l r(j)u_j(i)$. Let $u' \in \mathbb{F}_q^n$ be the resulting vector. (As discussed in Section 3, this challenge can be compressed using either cryptography or simple derandomization techniques.) The sharing phase succeeds if $u' \in V$.

Reconstruction. In the reconstruction phase, each server sends its shares to output client, who recovers the secrets s_j by error-correcting the (potentially) corrupted codewords u_j^* .

We assume in the following that $n > 3t$, implying that the minimal distance of the underlying RS code satisfies $d = n - t > 2t$. We start by considering the case of a *static* adversary, who decides in advance which t servers to corrupt. Completeness follows from the fact that $t < d/2$ and each u_j^*

is t -close to V . Secrecy follows from the secrecy property of Shamir’s scheme and from the fact that u_0 blinds the information exchanged during the verification process. For the unique reconstruction property, consider the shares of the $n - t > 2t$ uncorrupted servers. If they are not consistent with V , then (by a simple analysis) the sharing phase will fail except with $1/q$ probability. If they are consistent with V , then (as before) u_j^* is within the unique decoding radius and the outputs will be correct.

What goes wrong when the adversary can be adaptive? In this case, the dealer could potentially distribute badly formed vectors \mathbf{u} that have the following devious property: there is u_j which is very far from the code, and yet u' is with high probability (say, $1/2$) t -close to the code. Now, the adaptive adversary can corrupt only those servers in $T = \Gamma(u', V)$ and send on their behalf field elements that make u' consistent with V . This in turn makes the sharing phase succeed. But since u_j is not within the unique decoding distance from V , we lose the unique decoding guarantee. Theorem 1.2 rules out the existence of such a devious \mathbf{u} . But this is not enough. We actually need to ensure that \mathbf{u} is consistent with V when restricted to the $n - t$ servers that are not corrupted during reconstruction phase.

This is ensured by the stronger guarantee of Theorem 3.1. The analysis proceeds as follows. If \mathbf{u} is t -far from the interleaved code \mathbf{V} , then u' will be t -far from V except with $\leq n/q$ probability. In this event, even adaptive corruption cannot make u' look consistent with V , and the sharing phase fails. If \mathbf{u} is within distance $t' \leq t$ from \mathbf{V} , then except with $\leq n/q$ probability we have that $\Gamma(u', V) = \Gamma(\mathbf{u}, \mathbf{V})$. Denoting this set by T' , the adversary must corrupt the entire set T' for the sharing phase to succeed, and may additionally corrupt $t - t'$ more servers. In this case, \mathbf{u} restricted to the $n - t$ uncorrupted servers is fully consistent with V , guaranteeing unique reconstruction.

We finally note that security against an adaptive adversary can be reduced to security against a static adversary via a generic union bound argument, taking the union over all $\binom{n}{t}$ sets of servers that can be eventually be corrupted [CDD⁺04]. (Alternatively, this follows from the simple derivation of proximity gaps over exponentially large fields.) However, this would require the field size q to satisfy $q = 2^{\Omega(n)}$, which would make communication grow quadratically (rather than linearly) with the number of servers n . There are VSS protocols that meet this quadratic bound and achieve perfect (rather than statistical) security against an adaptive adversary [BGW88].

7.2 Fast RS IOP of Proximity (FRI) Soundness

In this section we use Theorem 1.6 to improve the soundness of the FRI protocol [BBHR18b]. We start by briefly recalling the essential components of the protocol, needed to state our theorem (see [BBHR18b] for a more thorough explanation of the protocol).

The FRI protocol As explained in Section 3.2, the purpose of FRI is to verify, in the IOP model, the proximity of a received word $f^{(0)} : \mathcal{D}^{(0)} \rightarrow \mathbb{F}$ to an RS code $V^{(0)} := \text{RS}[\mathbb{F}, \mathcal{D}^{(0)}, k^{(0)}]$. FRI requires $\mathcal{D}^{(0)}$ to be a “2-smooth” group, meaning $|\mathcal{D}^{(0)}| = 2^s$, for integer s . Henceforth we assume the group $\mathcal{D}^{(0)}$ is multiplicative⁵. The protocol has two phases, called COMMIT and QUERY. During the COMMIT phase, a sequence of functions $f^{(1)} : \mathcal{D}^{(1)} \rightarrow \mathbb{F}, f^{(2)} : \mathcal{D}^{(2)} \rightarrow \mathbb{F}, \dots, f^{(r)} : \mathcal{D}^{(r)} \rightarrow \mathbb{F}$ is generated over a finite number r of interactive rounds. At the beginning of the i -th round, the prover message $f^{(i)} : \mathcal{D}^{(i)} \rightarrow \mathbb{F}$ has already been created (and the verifier has oracle access to it). The verifier now sends a uniformly random $z^{(i)} \in \mathbb{F}$ and the prover replies with a new function $f^{(i+1)} : \mathcal{D}^{(i+1)} \rightarrow \mathbb{F}$ where $\mathcal{D}^{(i+1)}$ is a (2-smooth) strict subgroup of $\mathcal{D}^{(i)}$.

⁵ Generalizing the results here to (i) the additive case, (ii) t -smooth groups for larger constant t and (iii) cosets of such groups, is straightforward and omitted.

$\mathcal{D}^{(i+1)}$ partitions $\mathcal{D}^{(i)}$ into cosets of size $l^{(i)} := |\mathcal{D}^{(i)}|/|\mathcal{D}^{(i+1)}|$. Let C_g denote the coset corresponding to $g \in \mathcal{D}^{(i+1)}$, namely $C_g := \{g' \in \mathcal{D}^{(i)} \mid (g')^{l^{(i)}} = g\}$. For each coset C_g , there exists a linear map $M_g : \mathbb{F}^{C_g} \rightarrow \mathbb{F}^{l^{(i)}}$ over \mathbb{F} that maps $f^{(i)}|_{C_g}$ — the restriction of $f^{(i)}$ to domain $C_g \subset \mathcal{D}^{(i)}$ — to a sequence $u_0(g), \dots, u_{l^{(i)}-1}(g)$ and the following holds (see [BBHR18b, Section 4.1] for a proof).

Claim 7.1. *Suppose that $f^{(i)} \in \text{RS}[\mathbb{F}, \mathcal{D}^{(i)}, k^{(i)}]$ where $k^{(i)} + 1$ is an integral power of 2. Then, for any $z^{(i)} \in \mathbb{F}$, the function $f_{z^{(i)}}^{(i+1)} : \mathcal{D}^{(i+1)} \rightarrow \mathbb{F}$ defined on $g \in \mathcal{D}^{(i+1)}$ by*

$$f_{z^{(i)}}^{(i+1)}(g) := \sum_{j=0}^{l^{(i)}-1} \left(z^{(i)}\right)^j \cdot u_j(g)$$

is a valid codeword of $\text{RS}[\mathbb{F}, \mathcal{D}^{(i+1)}, k^{(i+1)}]$ where $k^{(i+1)} := (k^{(i)} + 1)/l^{(i)} - 1$.

This claim implies that an honest prover starting with a legitimate codeword $f^{(0)} \in V^{(0)}$ may use it to construct $f^{(1)}, \dots, f^{(r)}$. And the QUERY phase of FRI checks whether this is the case, by sampling a random $g^{(1)} \in \mathcal{D}^{(1)}$, checking whether $f^{(1)}(g^{(1)})$ indeed equals what the claim predicts based on the values of $f^{(0)}|_{C_{g^{(1)}}$, and then repeating the process inductively for $g^{(2)}, \dots, g^{(r)}$. (We refer the reader to [BBHR18b, BKS18] for more details.)

Summarizing what is stated above:

1. At the end of the protocol's COMMIT phase the verifier will have oracle access to a sequence of functions $f^{(0)} : \mathcal{D}^{(0)} \rightarrow \mathbb{F}, \dots, f^{(r)} : \mathcal{D}^{(r)} \rightarrow \mathbb{F}$ where $\mathcal{D}^{(0)} \supseteq \dots \supseteq \mathcal{D}^{(r)}$ is a sequence of 2-smooth groups and $f^{(i)}$ depends arbitrarily on $z^{(0)}, \dots, z^{(i)}$ (and $f^{(0)}, \dots, f^{(i-1)}$).
2. There exists a $l^{(i)}$ -to-1 map that can be applied to cosets of $f^{(i)}$ and transform them into a sequence of vectors $\mathbf{u} = \{u_0, \dots, u_{l^{(i)}}\} \subset \mathbb{F}^{\mathcal{D}^{(i+1)}}$. Furthermore, if $f^{(i)}$ is a valid RS codeword over $\mathcal{D}^{(i)}$ of rate ρ , then each vector on the parameterized curve passing through \mathbf{u} is a valid RS codeword over $\mathcal{D}^{(i+1)}$ of the same rate ρ .
3. Each iteration of the QUERY phase checks whether $f^{(i+1)}$ was constructed from $f^{(i)}$ via the process specified by Claim 7.1.

Claim 7.1 implies perfect completeness. Regarding soundness, Theorem 1.6 implies the following statement.

Theorem 7.2 (FRI Soundness). *Consider the code $\text{RS}[\mathbb{F}, \mathcal{D}^{(0)}, k^{(0)}]$ where $\mathcal{D}^{(0)}$ is a coset of a 2-smooth group of size $n^{(0)} = |\mathcal{D}^{(0)}|$, and $\rho = \frac{k^{(0)}+1}{n^{(0)}}$ satisfies $\rho = 2^{-R}$ for positive integer R . Let $n^{(0)} > n^{(1)} > \dots > n^{(r)}$ be the sequence of sizes of functions sent by the prover during the COMMIT phase, and let $l^{(i)} := n^{(i)}/n^{(i+1)}$ for $i < r$. Suppose $f^{(0)} : \mathcal{D}^{(0)} \rightarrow \mathbb{F}$ is at least δ -far from the code, for $\delta \leq 1 - \sqrt{\rho} - \eta$ and $0 < \eta \leq \frac{\sqrt{\rho}}{20}$.*

In this case, the FRI protocol has COMMIT soundness error at most

$$\epsilon_{\text{C}} \leq \frac{\rho^2}{(2\eta)^7 \cdot |\mathbb{F}|} \cdot \sum_{j < r} \left(l^{(j)} - 1\right) \cdot \left(n^{(j+1)}\right)^2 < \frac{(k^{(0)} + 1)^2}{(2\eta)^7 \cdot |\mathbb{F}|}.$$

Thus, for any malicious FRI prover, with probability $\geq 1 - \epsilon_{\text{COMMIT}}$ over the randomly selected $(z^{(0)}, \dots, z^{(r-1)})$, the probability of accepting $f^{(0)}$ erroneously during the QUERY phase, using a single iteration of the QUERY protocol, is at most

$$\epsilon_{\text{Q}} \leq (1 - \delta).$$

Consequently, for $f^{(0)}$ specified above and any malicious prover, the total soundness error ϵ_{FRI} of the FRI protocol with t iterations of the QUERY phase is at most

$$\epsilon_{\text{FRI}} = \epsilon_{\text{C}} + \epsilon_{\text{Q}}^t \leq \frac{(k^{(0)} + 1)^2}{(2\eta)^7 \cdot |\mathbb{F}|} + (1 - \delta)^t.$$

Numerical Example: Suppose $q \geq 2^{256}$, $n = 2^{20}$, $\rho = 2^{-4}$, so $k + 1 = 2^{16}$, an entirely reasonable choice for practical applications (In STARKs, k corresponds to the length of a computation for which a STARK proof is being generated). Set $\eta = 2^{-14}$, so

$$\epsilon_{\text{C}} \leq 2^{32+7 \cdot 13-256} = 2^{-133}.$$

Assume $f^{(0)}$ is δ -far from V for

$$\delta = 1 - \sqrt{\rho} - \eta \approx 0.7499 \approx \frac{3}{4}.$$

Invoking the QUERY protocol for $t = 65$ invocations gives

$$\epsilon_{\text{Q}}^t \approx 0.25006^{65} < 2^{-129.9}$$

So the total FRI error is bounded by

$$\epsilon_{\text{FRI}} \leq \epsilon_{\text{C}} + \epsilon_{\text{Q}}^t < 2^{-128}.$$

Thus we can get 128-bit security for the FRI protocol for this setting of parameters with 65 invocations of the QUERY phase.

Proof sketch of Theorem 7.2. The proof follows directly the proof of [BKS18, Theorem 16], while replacing Theorem 10 there with our improved Theorem 1.6. Notice this leads to a version of Corollary 17 there with proximity and error parameters as in our Theorem 1.6, which directly leads to Theorem 7.2. \square

Discussion and Two Open Questions Theorem 7.2 improves on the previous state of the art, due to [BKS18, BGKS20], in several ways. First, as explained in Section 3.2, the prior state of the art required a proximity parameter that is smaller than the “one-and-a-half Johnson bound”: $\delta < 1 - \sqrt[3]{\rho}$; the current result pushes the proximity parameter (for large fields) up to the Johnson/Guruswami–Sudan bound. Second, the error parameter during the COMMIT phase was worse, and the analysis incurred an additional loss in the proximity parameter during the QUERY phase, which led to worse concrete soundness bounds across all proximity parameter settings. Last, but not least, the prior bounds on δ (above the unique decoding radius) were only valid for the case where the “folding parameters” $l^{(0)}, \dots, l^{(r-1)}$ were all equal to the fixed value $l^* = 2$, and deteriorated swiftly for larger $l^{(i)}$ (they only work up to radius $1 - \sqrt[3]{\rho}$). The current bounds deteriorate much more slowly with $l^{(i)}$, and this is important because large values of $l^{(i)}$ are often preferable in practice.

Ben-Sasson et al. showed in [BBHR18b] that for the FRI protocol to achieve security parameter λ (i.e., $\epsilon_{\text{FRI}} \leq 2^{-\lambda}$), we need to take the repetition parameter $t \geq \lambda \log \frac{1}{\rho}$ and conjectured that this lower bound on t is also sufficient (for sufficiently large fields). As noticed earlier in Section 3.2 our results imply that taking $t \leq 2\lambda \log \frac{1}{\rho}$ suffices to get security parameter λ for quadratically large

fields. Closing the gap between the provable upper and lower bounds on t is left as an interesting open problem. Concretely, the following conjectured improvement to our main correlated agreement theorem (Theorem 1.4) will imply the conjecture of [BBHR18b]. To the best of our knowledge, nothing contradicts setting $c_1 = c_2 = 2$ in the conjecture below. When limiting the scope to prime fields, which have no sub-fields, we are not aware of anything contradicting $c_1 = c_2 = 1$; notice that [BGKS20, Appendix B] shows these smaller exponents cannot hold for fields of characteristic two.

Conjecture 7.3. *There exist constants c_1, c_2 such that the following statements hold for all $\eta > 0$.*

- *Theorems 1.2, 1.4 and 1.7 hold for proximity parameter $\delta \leq 1 - \rho - \eta$ with error*

$$\epsilon \leq \frac{1}{(\eta\rho)^{c_1}} \cdot \frac{n^{c_2}}{q}.$$

- *Theorem 1.6 holds for proximity parameter $\delta \leq 1 - \rho - \eta$ and parameterized curves of degree l with error*

$$\epsilon \leq \frac{1}{(\eta\rho)^{c_1}} \cdot \frac{(l \cdot n)^{c_2}}{q}.$$

DEEP FRI is another RPT protocol that is closely related to FRI (as its name suggests). Introduced in [BGKS20], it is slightly less efficient (in terms of prover and verifier complexity) than FRI because it requires more from the prover, making it harder for a malicious prover to cheat. Prior to this work, the extra complexity of DEEP FRI led to improved soundness, which reaches the Johnson/Guruswami–Sudan bound. But Theorem 7.2 shows that FRI has soundness that also reaches the same bound. Moreover, when nearing the Johnson bound, DEEP FRI requires cubic size fields for the arguments to work, whereas FRI is shown here to require only quadratic size fields. Thus, according to our new understanding, in terms of both complexity and field size, FRI dominates DEEP FRI, even though DEEP FRI demands strictly more from the prover side.⁶ Thus, an interesting second question raised by our work is understanding how the techniques developed here may be combined with the techniques of the DEEP FRI protocol to derive better soundness bounds for DEEP FRI and new, improved RPT protocols.

⁶The one remaining virtue of DEEP FRI is that its soundness is closely connected to *a classically studied problem* (the list-decodability of Reed–Solomon codes), and under a simple, plausible conjecture about that problem, it achieves the optimal bound on the required repetition parameter $t = (1 + o(1))\lambda \log \frac{1}{\rho}$.

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A Algebraic Extensions of $\mathbb{F}_q(Z)$

In this section we develop preliminaries that will be necessary for the proof of Theorem 5.1. In the proof, we will have a trivariate polynomial $Q(X, Y, Z)$, with an irreducible factor $R(X, Y, Z)$. When evaluated at a certain $X = x_0 \in \mathbb{F}_q$, $R(x_0, Y, Z)$ will have an irreducible factor $H(Y, Z)$. For both R and H , irreducible means irreducible in the respective ring of polynomials over \mathbb{F}_q . Our goal will be to show that Q has a factor of the form $Y - P(X, Z)$, where P has low X and Z degree, and in fact R is this factor. Considering Q and R as polynomials in Y over $\mathbb{F}_q[X, Z]$, this is equivalent to finding a rational root $P(X, Z)$ of Q , which is also a root of R , and $P(x_0, Z)$ is thus a root of H . To do so, we will instead start by understanding roots of H , not necessarily of the required form or even lying in $\mathbb{F}_q(Z)$, lifting them to roots of R (and Q), and then investigating these lifts to prove that they are indeed of the required form $P(X, Z)$.

We will therefore need to understand the roots of $H(Y, Z)$, which naturally lie in the field $\mathbb{F}_q(Z)[T]/(H(T, Z))$, which is a finite algebraic extension of $\mathbb{F}_q(Z)$. We will construct this field, in a slightly different way, and also introduce its ring of *regular* elements, in Appendix A.1. In this field, we will need a concept of *weight*, which is an extension of the concept of “degree” from $\mathbb{F}_q(Z)$, and is introduced in Appendix A.2. A major tool that we will use in several ways is of substitution maps from this field to \mathbb{F}_q , extending the concept of substituting $Z = z$ in a rational function in $\mathbb{F}_q(Z)$. These substitutions and a useful lemma regarding them are introduced in Appendix A.3. The final tool we need is the lifting of roots of $H(Y, Z)$ to roots of $R(X, Y, Z)$, which is known as Hensel lifting. We describe this process and state an important lemma on the weights of the field elements appearing in this lift in Appendix A.4.

A.1 The algebraic extension and its regular elements

Let $d \geq 1$ be integers, and let

$$H(Y, Z) = h_0(Z)Y^d + h_1(Z)Y^{d-1} + \cdots + h_d(Z) \in \mathbb{F}_q[Y, Z]$$

be an irreducible bivariate polynomial, expressed as a polynomial in Y over $\mathbb{F}_q[Z]$. Denote the leading coefficient of H also as $W = h_0$.

We wish to understand the field $\mathbb{F}_q(Z)[T]/(H(T, Z))$, of polynomials in T over $\mathbb{F}_q(Z)$ modulo H . The presence of the leading coefficient W means that T is not an integer in this field, and makes the arithmetic modulo H unpleasant to keep track of, with possible emergence of high powers of W in denominators whenever a reduction modulo H is performed. To avoid this, we first define a “monicized” version of H , denoted \tilde{H} , which is a monic, irreducible polynomial generating the same field as H :

$$\begin{aligned} \tilde{H}(T, Z) &= W(Z)^{d-1}H(T/W(Z), Z) \\ &= T^d + h_1(Z)T^{d-1} + h_2(Z)W(Z)T^{d-2} + \cdots + h_d(Z)W(Z)^{d-1}. \end{aligned}$$

We now denote $\mathbb{L} = \mathbb{F}_q(Z)[T]/(\tilde{H}(T, Z))$, and observe that $Y = \frac{T}{W(Z)}$ is a root of $H(Y, Z)$ in \mathbb{L} . We say that an element of \mathbb{L} is *regular* if it can be expressed as a polynomial in T with coefficients only in the ring $\mathbb{F}_q[Z]$ instead of the field $\mathbb{F}_q(Z)$; equivalently, if this is true for its canonical form as a polynomial in T of degree less than d . We denote the set of regular elements by $\mathcal{O} = \mathbb{F}_q[Z][T]/(\tilde{H}(T, Z))$. The regular elements are in fact a ring, which is a subring of the ring of integers of \mathbb{L} .

A.2 Algebraic weights

Let $D \geq d$ be the total Y, Z degree of H , so that $\deg_Z h_k \leq D + k - d$ for all k . We define a *weight* function Λ on $\mathbb{F}_q[T, Z]$ by assigning $\Lambda(Z) = 1$ and $\Lambda(T) = D + 1 - d$, extended additively to monomials, i.e. $\Lambda(T^a Z^b) = a\Lambda(T) + b\Lambda(Z)$, and the weight of a polynomial is the maximal weight of all monomials appearing in it with non-zero coefficients (with the weight of the 0 polynomial being $-\infty$). Note that Λ is fully additive on $\mathbb{F}_q[T, Z]$, i.e. for any $A, B \in \mathbb{F}_q[T, Z]$, $\Lambda(AB) = \Lambda(A) + \Lambda(B)$. Also note that when restricted to $\mathbb{F}_q[Z]$, $\Lambda = \deg_Z$.

Observe that $\tilde{H}(T, Z)$ has weight $\Lambda(\tilde{H}) = d(D + 1 - d) = dD - d(d - 1)$, with the leading monomial being of this exact weight and every other monomial bounded by it. It follows that any simple modulo $\Lambda(\tilde{H})$ operation of the form

$$T^{d+k} \rightarrow - \sum_{i=1}^d h_i(Z) W(Z)^{i-1} T^{d+k-i}$$

never increases the weight Λ , and so does complete reduction modulo \tilde{H} .

We now define the weight $\Lambda(\alpha)$ of a regular element $\alpha \in \mathcal{O}$ as the weight of the canonical representative of α with degree less than d , which by the above is also the minimal value of Λ over all representatives of α . It also follows that for any $\alpha, \beta \in \mathcal{O}$, if $A(T, Z), B(T, Z)$ are their canonical representatives, and $C(T, Z) = A(T, Z)B(T, Z) \bmod \tilde{H}(T, Z)$ is the canonical representative of $\gamma = \alpha\beta$, then

$$\Lambda(\gamma) = \Lambda(C) \leq \Lambda(AB) = \Lambda(A) + \Lambda(B) = \Lambda(\alpha) + \Lambda(\beta).$$

In other words, Λ is sub-additive in \mathcal{O} .

A.3 Rational substitutions

Let $z \in \mathbb{F}_q$ be such that $\tilde{H}(T, z)$ has a rational root $T = t_z$. In other words, $(t_z, z) \in \mathbb{F}_q^2$ is a root of \tilde{H} , with t_z considered as depending on z (in our applications, it will in fact be given as a function of z). For any such root-pair, we define the *substitution* π_z , which is the homomorphism $\pi_z : \mathcal{O} \rightarrow \mathbb{F}_q$ given by $\pi_z(Z) = z, \pi_z(T) = t_z$. The homomorphism is well defined since $\mathcal{O} = \mathbb{F}_q[T, Z]/(\tilde{H}(T, Z))$ and $\pi_z(\tilde{H}(T, Z)) = \tilde{H}(t_z, z) = 0$. The substitution π_z can be extended naturally to any element of \mathbb{L} for which z is not a pole, i.e. elements of the form $\frac{\beta}{C(Z)}$ with $\beta \in \mathcal{O}$ and z not a root of C , by $\pi_z\left(\frac{\beta}{C(Z)}\right) = \frac{\pi_z(\beta)}{C(z)}$.

The following lemma gives an upper bound on the number of substitutions in which an element $\beta \in \mathcal{O}$ can vanish in terms of its weight. It is analogous to the statement that a polynomial of degree at most d which vanishes for more than d evaluations must be the 0 polynomial.

Lemma A.1. *Let $\beta \in \mathcal{O}$ be regular with weight $\Lambda(\beta)$. Let*

$$S_\beta = \{z \in \mathbb{F}_q : \exists t_z \in \mathbb{F}_q, \tilde{H}(t_z, z) = 0 \text{ and } \pi_z(\beta) = 0\}$$

and suppose $|S_\beta| > d \cdot \Lambda(\beta)$. Then $\beta = 0 \in \mathbb{L}$.

Proof. Let $\beta = \sum_{i=0}^{d-1} \beta_i(Z) T^i \in \mathbb{F}_q[Z][T]$ be the canonical representative, with

$$\deg_Z(\beta_i) = \Lambda_Z(\beta_i) \leq \Lambda(\beta) - i\Lambda(T) = \Lambda(\beta) - i(D + 1 - d).$$

Consider the resultant $R(Z) = \text{Res}_T(\beta, \tilde{H})$. From the Sylvester matrix, or considering weights of roots, we find $\deg_Z R \leq d \cdot \Lambda(\beta) < |S_\beta|$. On the other hand, every $z \in S_\beta$ is a root of R , since

$\tilde{H}(T, z)$ and $\beta(T, z)$ have a common root t_z . It follows that R is identically 0, i.e. β and \tilde{H} are not coprime, but since \tilde{H} is irreducible and $\deg_T(\beta) < \deg_T(\tilde{H})$, we have $\beta = 0$, as claimed. \square

A.4 Hensel lifts

Suppose $H(Y, Z)$ as above is a factor of $R_{x_0}(Y, Z) = R(x_0, Y, Z)$, where $x_0 \in \mathbb{F}_q$, and $R(X, Y, Z)$ is irreducible in $\mathbb{F}_q[X, Y, Z]$. Additionally assume that R and R_{x_0} are both separable in Y , which means that they do not have double roots when considered as polynomials in Y , or equivalently, that they are coprime to their Y -derivatives. Denoting $\alpha_0 = \frac{T}{W} \in \mathbb{L}$, we find $R(x_0, \alpha_0, Z) = 0 \in \mathbb{L}$, or equivalently in $\mathbb{L}[X]$,

$$R(X, \alpha_0, Z) \equiv 0 \pmod{X - x_0}.$$

The separability of R_{x_0} implies that $\zeta = \frac{\partial R}{\partial Y}(x_0, \alpha_0, Z) \neq 0$, or in other words $\zeta \in \mathbb{L}$ is invertible. This fact allows us to iteratively lift the root α_0 of $R(X, Y, Z) \pmod{X - x_0}$ to a root

$$\alpha_0 + \alpha_1(X - x_0) \quad \text{of} \quad R(X, Y, Z) \pmod{(X - x_0)^2},$$

then to a root

$$\alpha_0 + \alpha_1(X - x_0) + \alpha_2(X - x_0)^2 \quad \text{of} \quad R(X, Y, Z) \pmod{(X - x_0)^3},$$

and so on. Moreover, at each step the lifting is unique.

We obtain an infinite sequence $(\alpha_t)_{t=0}^\infty$, such that at each step s the truncated series $\gamma_s = \sum_{t=0}^s \alpha_t(X - x_0)^t \in \mathbb{L}[X]$ satisfies

$$R(X, \gamma_s, Z) \equiv 0 \pmod{(X - x_0)^{s+1}}.$$

Equivalently, for the infinite formal power series $\gamma = \sum_{t=0}^\infty \alpha_t(X - x_0)^t \in \mathbb{L}[[X - x_0]]$, we have $R(X, \gamma, Z) = 0 \in \mathbb{L}[[X - x_0]]$. Here $\mathbb{L}[[X - x_0]]$ is the ring of formal power series in $X - x_0$ with coefficients in \mathbb{L} . This power series γ is the Hensel lift of α_0 , and the process by which it is computed is the Hensel lifting.

We make a slight change to the notation, and henceforth d will be the Y -degree of R , and $d_H \leq d$ will denote the Y -degree of H (previously denoted by d). We will also assume that D is an upper bound not only on the total degree of H but also of R . Note that W , the leading coefficient of H , divides the leading coefficient of R_{x_0} , and has weight $\Lambda(W) \leq D - d_H$.

The following claim describes the coefficients α_t appearing in the Hensel lift, and bounds their denominators and the weights of their numerators:

Claim A.2. α_t is of the form $\frac{\beta_t}{W^{t+1}\xi^{e_t}}$, where

- $\xi = W(Z)^{d-2}\zeta \in \mathcal{O}$ with $\zeta = \frac{\partial R}{\partial Y}(x_0, \frac{T}{W}, Z) \in \mathbb{L}$, and

$$\Lambda(\xi) \leq (D - 1) + (d - 2)\Lambda(W) \leq (d - 1)(D - d_H + 1),$$

- $e_t = \max(0, 2t - 1)$, i.e. $e_0 = 0$ and $e_t = 2t - 1$ for $t \geq 1$,

- $\beta_t \in \mathcal{O}$ with

$$\begin{aligned} \Lambda(\beta_t) &= 1 + (t + 1)\Lambda(W) + e_t\Lambda(\xi) \\ &\leq ((d - 1) \cdot e_t + t + 1)(D - d_H + 1) - t < (2t + 1)dD. \end{aligned}$$

The claim is proven by straight-forwardly expanding $R(X, \gamma, Z)$ as a series in $X - x_0$, comparing each coefficient to 0, and using induction on t . The existence and uniqueness of the Hensel lift γ , and how each new coefficient α_t is computed from the previous ones, will also follow from the proof.

Proof. We prove the statement by induction on t . For $t = 0$, we have simply $\alpha_0 = \frac{T}{W} \bmod \tilde{H}$, i.e. $\beta_0 = T \bmod \tilde{H}$ and indeed $\Lambda(T) = \Lambda(W) + 1$.

We wish to write $R(X, \gamma, Z)$ as a power series in $X - x_0$. A *partition* of t is a sequence of non-negative integers $\lambda = (\lambda_l)_{l \geq 1}$ with $\sum_l l \cdot \lambda_l = t$. Such sequences are non-zero only finitely many times, and we trim any trailing 0s in writing, e.g. $(1) = (1, 0, 0, 0, \dots)$. We also denote $\Sigma\lambda = \sum_{l \geq 1} \lambda_l$. Let $\mathcal{P}(t)$ be the set of partitions of t . For any $t, i_1 \leq t$, and $\lambda \in \mathcal{P}(t - i_1)$, denote

$$\begin{aligned} A_{i_1, \lambda} &= \sum_{\substack{j_0 \\ j = j_0 + \Sigma\lambda}} \binom{j}{j_0, \lambda_1, \dots, \lambda_l, \dots} \alpha_0^{j_0} \sum_{\substack{i_0 \\ i = i_0 + i_1}} \binom{i}{i_0, i_1} R_{ji}(Z) x_0^{i_0} \\ &= \sum_{\substack{i_0, j_0 \\ i = i_0 + i_1 \\ j = j_0 + \Sigma\lambda}} \binom{i}{i_0, i_1} \binom{j}{j_0, \lambda_1, \dots, \lambda_l, \dots} R_{ji}(Z) x_0^{i_0} \alpha_0^{j_0} \\ &= \binom{\Sigma\lambda}{\lambda_1, \dots, \lambda_l, \dots} \sum_{i, j} \binom{i}{i_1} \binom{j}{\Sigma\lambda} R_{ji}(Z) x_0^{i-i_1} \alpha_0^{j-\Sigma\lambda} \\ &= \binom{\Sigma\lambda}{\lambda_1, \dots, \lambda_l, \dots} \left(\Delta_X^{i_1} \Delta_Y^{\Sigma\lambda} R(X, Y, Z) \right) \Big|_{(x_0, \alpha_0, Z)} \end{aligned}$$

where the sums are taken only over non-zero terms, i.e. with $i + kj < D_X$, and Δ_V^k is the k -th Hasse derivative with respect to the variable V . In particular, whenever λ is such that $\Sigma\lambda = 1$, we have

$$A_{0, \lambda} = (\Delta_Y R)(x_0, \alpha_0, Z) = \frac{\partial R}{\partial Y}(x_0, \alpha_0, Z) = \zeta.$$

Since the maximal degree of $\alpha_0 = \frac{T}{W}$ appearing in $A_{i_1, \lambda}$ is $d - \Sigma\lambda$, we can generally write $A_{i_1, \lambda} = \frac{B_{i_1, \lambda}}{W^{d-\Sigma\lambda}}$, where $B_{i_1, \lambda} \in \mathcal{O}$ has weight $\Lambda(B_{i_1, \lambda}) = (D - \Sigma\lambda) + (d - \Sigma\lambda)\Lambda(W)$. In the special case $i_1 = 0$, the coefficient of $\alpha_0^{d-\Sigma\lambda}$ is a multiple of $\sum_i R_{di}(Z) x_0^i$, which is the leading coefficient of R_{x_0} , hence divisible by W , and thus we can save a little and write $A_{0, \lambda} = \frac{B_{0, \lambda}}{W^{d-1-\Sigma\lambda}}$, where $B_{0, \lambda} \in \mathcal{O}$ has weight $(D - \Sigma\lambda) + (d - 1 - \Sigma\lambda)\Lambda(W)$. When $\Sigma\lambda = 1$ we then get $\zeta = A_{0, \lambda} = \frac{B_{0, \lambda}}{W^{d-2}} = \frac{\xi}{W^{d-2}}$ where $\xi \in \mathcal{O}$ has weight $(D - 1) + (d - 2)\Lambda(W)$ as stated. To generalize the two cases, we may write $A_{i_1, \lambda} = \frac{B_{i_1, \lambda}}{W^{d-\delta_{i_1, 0}-\Sigma\lambda}}$, where $\delta_{i_1, 0} = 1$ if $i_1 = 0$ and $\delta_{i_1, 0} = 0$ otherwise.

Now, expanding $R(X, \gamma, Z)$, we get

$$\begin{aligned} 0 &= R(X, \gamma, Z) = R\left(x_0 + (X - x_0), \sum_l \alpha_l (X - x)^l, Z\right) \\ &= \sum_{ij} R_{ji}(Z) (x_0 + (X - x_0))^i \left(\sum_l \alpha_l (X - x)^l\right)^j \\ &= \sum_{ij} R_{ji}(Z) \left(\sum_{i_0+i_1=i} \binom{i}{i_0, i_1} x_0^{i_0} (X - x_0)^{i_1} \right. \\ &\quad \left. \times \sum_{j_0+\Sigma\lambda=j} \binom{j}{j_0, \lambda} \left(\alpha_0^{j_0} \prod_l \alpha_l^{\lambda_l} \right) (X - x_0)^{\sum_l l \cdot \lambda_l} \right) \end{aligned}$$

$$= \sum_{t=0}^{\infty} (X - x_0)^t \sum_{\substack{i_1 \\ \lambda \in \mathcal{P}(t - i_1)}} A_{i_1, \lambda} \prod_{l \geq 1} \alpha_l^{\lambda_l}$$

Note that α_t appears for the first time in the term corresponding to $(X - x_0)^t$, and only with $i_1 = 0$ and $\lambda = \lambda^{(t)}$ defined by $\lambda_t^{(t)} = 1$, $\lambda_l^{(t)} = 0$ for $l \neq t$, with the coefficient $A_{0, \lambda^{(t)}} = \zeta$. All other summands in the coefficient of $(X - x_0)^t$ involve only α_l with $l < t$, so we may apply the induction. Comparing the coefficient to 0, we get

$$\begin{aligned} \alpha_t &= -\frac{1}{\zeta} \sum_{\substack{i_1; \lambda \in \mathcal{P}(t - i_1) \\ \lambda \neq \lambda^{(t)}}} A_{i_1, \lambda} \prod_{l \geq 1} \alpha_l^{\lambda_l} \\ &= -\frac{W^{d-2}}{\xi} \sum_{\substack{i_1; \lambda \in \mathcal{P}(t - i_1) \\ \lambda \neq \lambda^{(t)}}} \frac{B_{i_1, \lambda}}{W^{d - \delta_{i_1, 0} - \Sigma \lambda}} \prod_{l \geq 1} \left(\frac{\beta_l}{W^{l+1} \xi^{e_l}} \right)^{\lambda_l} \\ &= \sum_{\substack{i_1; \lambda \in \mathcal{P}(t - i_1) \\ \lambda \neq \lambda^{(t)}}} \frac{B_{i_1, \lambda} \prod_l \beta_l^{\lambda_l}}{W^{2 - \delta_{i_1, 0} - \Sigma \lambda + \sum_l (l+1) \lambda_l} \xi^{1 + \sum_l (2l-1) \lambda_l}} \\ &= \sum_{\substack{i_1; \lambda \in \mathcal{P}(t - i_1) \\ \lambda \neq \lambda^{(t)}}} \frac{B_{i_1, \lambda} \prod_l \beta_l^{\lambda_l}}{W^{t - i_1 - \delta_{i_1, 0} + 2} \xi^{2t - 2i_1 - \Sigma \lambda + 1}}, \\ &= \frac{1}{W^{t+1} \xi^{2t-1}} \sum_{\substack{i_1; \lambda \in \mathcal{P}(t - i_1) \\ \lambda \neq \lambda^{(t)}}} W^{i_1 + \delta_{i_1, 0} - 1} \xi^{2i_1 + \Sigma \lambda - 2} B_{i_1, \lambda} \prod_l \beta_l^{\lambda_l}, \end{aligned}$$

thus we have

$$\beta_t = \sum_{\substack{i_1; \lambda \in \mathcal{P}(t - i_1) \\ \lambda \neq \lambda^{(t)}}} W^{i_1 + \delta_{i_1, 0} - 1} \xi^{2i_1 + \Sigma \lambda - 2} B_{i_1, \lambda} \prod_l \beta_l^{\lambda_l} \quad (\text{A.1})$$

which is indeed regular, as the W , ξ , the B 's and the β 's are all regular and the exponents are non-negative: for the exponent of W , it is always the case that $i_1 + \delta_{i_1, 0} \geq 1$. For the exponent of ξ , for $i_1 = 0$, every $\lambda \in \mathcal{P}(t)$ with $\lambda \neq \lambda^{(t)}$ indeed has $\Sigma \lambda \geq 2$, and for $i_1 \geq 1$ we even have $2i_1 + \Sigma \lambda - 2 \geq 1$, since $\Sigma \lambda \geq 1$.

The upper bound on the weight of β can be shown by induction using the recursion (A.1), but an easier way to understand it is by considering the weight of α_t : Since $\gamma = \sum_{t=0}^{\infty} \alpha_t (X - x_0)^t$ is a solution to $R(X, Y, Z) = 0$, γ has the same weight as Y ; since X, x_0 have weight 0, each α_t also has weight $\Lambda(\alpha_t) = \Lambda(Y) = 1$. Thus

$$\Lambda(\beta_t) = \Lambda(\alpha_t W^{t+1} \xi^{e_t}) = 1 + (t+1)\Lambda(W) + e_t \Lambda(\xi)$$

as claimed. □

B The Y, Z -degree of Q

In this section we prove inequality (5.7) of Claim 5.4, which claims that the total Y, Z -degree of $Q(X, Y, Z)$ is bounded by $\frac{(m+1/2)^3}{6\sqrt{\rho}} n$.

The bound on $D_{YZ} = \deg_{Y,Z}(Q)$ comes from minors of the matrix M representing the system of equations defining Q . These equations are all of the form “the (m_X, m_Y) -th derivative of Q vanishes at $(x, w(x))$ ”, for $x \in \mathcal{D}$ and non-negative integers m_X, m_Y with $m_X + m_Y < m$. Computing the derivative⁷ and substituting $(x, w(x))$, we see that the (x, m_X, m_Y) -th equation is

$$\Delta_X^{m_X} \Delta_Y^{m_Y} Q(x, w(x, Z), Z) = \sum_{i+kj < D_X} Q_{ji}(Z) \binom{i}{m_X} \binom{j}{m_Y} x^{i-m_X} (u_0(x) + Zu_1(x))^{j-m_Y} = 0.$$

The coefficient of $Q_{ji}(Z)$ in this equation is therefore

$$\binom{i}{m_X} \binom{j}{m_Y} x^{i-m_X} (u_0(x) + Zu_1(x))^{j-m_Y},$$

and this coefficient appears in the matrix M at row (x, m_X, m_Y) and column ji . Note that as a polynomial in Z , it has degree at most $j - m_Y$, which is the sum of j , which is determined by the column, and $-m_Y$, determined by the row. We thus call j and $-m_Y$ the contributions of the column and the row, respectively, to the Z -degree of the matrix’s entry.

Let r be the rank of M , which is bounded from above by the number of rows $\binom{m+1}{2}n$. To find a non-zero solution of the system, first find an $r \times r$ non-singular submatrix, and add an arbitrary single column. Cramer’s rule then tells us that a solution to the $r \times (r+1)$ subsystem is given by taking each Q_{ji} to be the $r \times r$ minor obtained by removing the ji -th column from the submatrix and taking the determinant, with alternating signs. These Q_{ji} will then be a solution to the original system, since these r rows span the entire row-space of the original matrix. The determinant corresponding to Q_{ji} is a sum over products corresponding to permutations, each containing a single entry from each row and each column of the $r \times r$ submatrix. The degree of each such product is thus bounded by the sum of the degree contributions from all columns and rows of the $r \times r$ submatrix, regardless of the permutation, or of all columns and rows of the $r \times (r+1)$ matrix, minus that of the ji -th column. In other words, it is at most $D - j$, where $D = D_C - D_R$ is the sum of the contributions of all columns and (negative) contributions of all rows from the $r \times (r+1)$ matrix. Thus the total Y, Z degree of the monomial $Q_{ji}X^iY^j$ is at most $(D - j) + j \leq D$, and hence the total Y, Z degree of Q is at most D , i.e. $D_{YZ} \leq D$.

It remains to bound $D = D_C - D_R$ from above. We do this by simply using $D_R \geq 0$ and bounding D_C by the sum of column contributions of the entire matrix. It is possible to improve this bound by finding the worst r for which the sum of the largest $r+1$ column contributions minus the sum of the smallest r row contributions is maximal, and computing these sums. We opt for the simpler bound, since the optimal bound requires a more technical and involved computation, and only ends up improving on the simple bound by a small constant factor, with an unpleasant dependence on ρ .

Write $s = \lfloor \frac{D_X}{k} \rfloor$, $t = \left\{ \frac{D_X}{k} \right\} \in [0, 1)$ with $D_X = k(s + t)$. Then

$$D_{YZ} = \deg_{Y,Z}(Q) \leq D = D_C - D_R \leq D_C \leq \sum_{i+kj < D_X} j$$

⁷ We use the Hasse derivatives $\Delta_X^m(X^i) := \binom{i}{m}X^{i-m}$ instead of the regular derivatives to avoid complications due to the characteristic of the field.

$$\begin{aligned}
&= \sum_{j < \frac{D_X}{k}} j(D_X - j \cdot k) = D_X \sum_{j=0}^s j - k \sum_{j=0}^s j^2 = k(s+t) \frac{s(s+1)}{2} - k \frac{s(s+1)(2s+1)}{6} \\
&= \frac{ks(s+1)(s+3t-1)}{6} = \frac{k}{6}(s^3 + 3ts^2 + (3t-1)s) = \frac{k}{6}((s+t)^3 - (1-3t+3t^2)s - t^3) \\
&= \frac{k}{6}((s+t)^3 - (1-t)^3s - (s+1)t^3) < \frac{k}{6}(s+t)^3 = \frac{D_X^3}{6k^2} \leq \frac{(m+\frac{1}{2})^3}{6\sqrt{\rho}} n,
\end{aligned}$$

as claimed in (5.7). \square

C The Inseparable Factor Case in the List Decoding Regime

Recall that in Section 5.2.3, we had assumed that in the decomposition

$$Q(X, Y, Z) = C(X, Z) \prod_i R_i(X, Y^{p^{f_i}}, Z)^{e_i},$$

the factors $R_i(X, Y^{p^{f_i}}, Z)$ were all separable, i.e. $f_i = 0$. This assumption was in fact necessary only for the factor R on which we focused in Section 5.2.4. We now consider the case where $f = f_i > 0$, and our relevant factor is of the form $R(X, Y^{\mathfrak{p}}, Z)$, where $\mathfrak{p} = p^f$ and $R(X, \tilde{Y}, Z)$ is separable and irreducible in \tilde{Y} . Note that we still have that $R(x_0, \tilde{Y}, Z)$ is separable. The elements of $S_{x_0, R, H}$ now satisfy $Y - P_z(X) \mid R(X, Y^{\mathfrak{p}}, Z)$, and equivalently $\tilde{Y} = P_z(X)^{\mathfrak{p}}$ is a root of $R(X, \tilde{Y}, Z)$ and $Y^{\mathfrak{p}} - P_z(X)^{\mathfrak{p}} \mid R(X, Y^{\mathfrak{p}}, Z)$. Similarly, $\tilde{Y} = P_z(x_0)^{\mathfrak{p}}$ is a root of the irreducible factor $H(\tilde{Y}, Z)$ of $R(x_0, \tilde{Y}, Z)$ — and a simple root of both, since they are separable.

Note that $\mathfrak{p} \leq \deg_Y(R(X, Y^{\mathfrak{p}}, Z)) \leq \deg_Y(Q(X, Y, Z)) = D_Y$ and therefore

$$\deg_X(P_z(X)^{\mathfrak{p}}) \leq p^f k \leq k D_Y < D_X.$$

We construct the field \mathbb{L} for the polynomial $H(\tilde{Y}, Z)$ exactly as before, but noting that $\tilde{Y} = Y^{\mathfrak{p}}$ will have weight \mathfrak{p} , which correspondingly affects the weight of $T = W(Z)\tilde{Y} = W(Z)Y^{\mathfrak{p}}$, now defined as $D - (d_H - 1)\mathfrak{p}$. Note that the upper bound on d_H, d (the \tilde{Y} degrees of R, H) is also changed, and is now $\frac{D_Y}{\mathfrak{p}}$ instead of simply D_Y . With these adaptations in mind, we perform the Hensel lift, lifting the root $\alpha_0 = \frac{T}{W} \in \mathbb{L}$ of $R(x_0, \tilde{Y}, Z)$ to the power series root $\gamma \in \mathbb{L}[[X - x_0]]$ of $R(X, \tilde{Y}, Z)$. Claim A.2 still holds and gives us

$$\begin{aligned}
\Lambda(\xi) &\leq (D - \mathfrak{p}) + (d - 2)\Lambda(W) \leq (d - 1)(D - (d_H - 1)\mathfrak{p}), \\
\Lambda(\beta_t) &= \mathfrak{p} + (t + 1)\Lambda(W) + e_t \Lambda(\xi) \\
&\leq ((d - 1) \cdot e_t + t + 1)(D - (d_H - 1)\mathfrak{p}) - \mathfrak{p}t < (2t + 1)dD.
\end{aligned}$$

As in Section 5.2.6, the substitution $\pi_z(\gamma)$ for $z \in S'$ is a root of $R(X, \tilde{Y}, z)$ which is the lift of the simple root $\tilde{Y} = P_z(x_0)^{\mathfrak{p}}$ of $R(x_0, \tilde{Y}, Z)$. Since $P_z(X)^{\mathfrak{p}}$ is also of this form, by the uniqueness of the lifting we get $\pi_z(\gamma) = P_z(X)^{\mathfrak{p}}$, and in particular $\pi_z(\alpha_t) = 0$ and $\pi_z(\beta_t) = 0$ for all t except for $0 \leq t \leq \mathfrak{p}k < D_X$ which are divisible by \mathfrak{p} . As before, we have $|S'| \geq d_H \Lambda(\beta_t)$ for all $t < D_X$: the right hand side is bounded from above by $\frac{D_Y^2 D_{YZ}(2D_X - 1)}{\mathfrak{p}^2}$, which only decreases as \mathfrak{p} increases. It thus follows that $\beta_t = 0$ and $\alpha_t = 0$ for all $t < D_X$ except for those which are at most $\mathfrak{p}k$ and divisible by \mathfrak{p} . In other words we have

$$\gamma_{D_X - 1} = \sum_{t=0}^k \alpha_{\mathfrak{p}t} (X - x_0)^{\mathfrak{p}t}.$$

Our next goal is to show that γ_{D_X-1} is a \mathfrak{p} -th power of a polynomial of degree k (and later that $\gamma = \gamma_{D_X-1}$, and that this polynomial is in fact in $\mathbb{F}_q[Z][X - x_0]$, with coefficients linear in Z). This part did not appear in Section 5, as it is trivial for $\mathfrak{p} = 1$. This polynomial should naturally be the \mathfrak{p} -th root of γ_{D_X-1} — but in order to construct such roots, we will need some more preliminaries about the field in which they live.

Let $\sigma : \mathbb{F}_q \rightarrow \mathbb{F}_q$ be the automorphism mapping each element a to its unique \mathfrak{p} -th root $\sigma(a) = a^{1/\mathfrak{p}}$. Let $\widehat{\mathbb{L}}$ be the inseparable algebraic field extension of \mathbb{L} with elements \widehat{T}, \widehat{Z} satisfying $\widehat{T}^{\mathfrak{p}} - T = \widehat{Z}^{\mathfrak{p}} - Z = 0$; equivalently, $\widehat{T} = T^{1/\mathfrak{p}}$ and $\widehat{Z} = Z^{1/\mathfrak{p}}$. Note that $\widehat{\mathbb{L}}$ can also be defined directly as the field $\mathbb{F}_q(\widehat{Z})[\widehat{T}]/(\widehat{H}(\widehat{T}, \widehat{Z}))$, where $\widehat{H} = \sigma(\widehat{H})$ is the (irreducible) polynomial obtained by applying σ to the coefficients of \widehat{H} , which satisfies

$$\widehat{H}(\widehat{T}, \widehat{Z})^{\mathfrak{p}} = \widetilde{H}(\widehat{T}^{\mathfrak{p}}, \widehat{Z}^{\mathfrak{p}}) = \widetilde{H}(T, Z).$$

Since the monomials of \widehat{H} have the same \widehat{T} - and \widehat{Z} -degrees as the T - and Z -degrees of \widetilde{H} , a weight $\widehat{\Lambda}$ can be defined for regular elements in $\widehat{\mathbb{L}}$ in exactly the same way as in \mathbb{L} . Additionally, σ can be extended to a map $\widehat{\sigma} : \mathbb{L} \rightarrow \widehat{\mathbb{L}}$ satisfying $\widehat{\sigma}(\alpha)^{\mathfrak{p}} = \alpha$ for all $\alpha \in \mathbb{L}$ by defining $\widehat{\sigma}(T) = \widehat{T}$ and $\widehat{\sigma}(Z) = \widehat{Z}$. Note that $\Lambda(\beta) = \widehat{\Lambda}(\widehat{\sigma}(\beta))$ for any $\beta \in \mathcal{O}$, since $\widehat{\sigma}$ preserves degrees. The substitution maps π_z for $z \in S'$ can also be extended to $\widehat{\mathbb{L}}$ by setting $\pi_z(\widehat{Z}) = \sigma(z) = z^{1/\mathfrak{p}}$ and $\pi_z(\widehat{T}) = \sigma(t_z) = t_z^{1/\mathfrak{p}} = \widehat{W}(z^{1/\mathfrak{p}})P_z(x_0)$, where $\widehat{W} = \sigma(W)$.

Define

$$\widehat{\gamma} = \sum_{t=0}^k \widehat{\sigma}(\alpha_{\mathfrak{p}t})(X - x_0)^t \in \widehat{\mathbb{L}}[X - x_0]$$

which indeed satisfies $\gamma_{D_X-1} = \widehat{\gamma}^{\mathfrak{p}}$, and therefore $R(X, \widehat{\gamma}^{\mathfrak{p}}, Z) \equiv 0 \pmod{(X - x_0)^{D_X}}$. Since $R(X, Y^{\mathfrak{p}}, Z)$ is a divisor of $Q(X, Y, Z)$, it follows as before that $\deg_X R(X, \widehat{\gamma}^{\mathfrak{p}}, Z) < D_X$, and therefore $R(X, \widehat{\gamma}^{\mathfrak{p}}, Z) = 0$ identically, and $\gamma = \gamma_{D_X-1} = \widehat{\gamma}^{\mathfrak{p}}$ by the uniqueness of the lifting. Furthermore, for every $z \in S'$, since $\pi_z(\widehat{\gamma})^{\mathfrak{p}} = \pi_z(\widehat{\gamma}^{\mathfrak{p}}) = \pi_z(\gamma) = P_z(X)^{\mathfrak{p}}$, we also have $\pi_z(\widehat{\gamma}) = P_z(X)$.

As in Section 5.2.7, our next goal is now to show that the coefficients $\widehat{\alpha}_t = \widehat{\sigma}(\alpha_{\mathfrak{p}t})$ of $\widehat{\gamma}$ are linear polynomials in $\mathbb{F}_q[Z]$, rather than general elements of $\widehat{\mathbb{L}}$. This is done in exactly the same way, by comparing the values of $\pi_z(\widehat{\gamma}(x))$ and $w(x, z)$ at every $x = x_j \in \mathcal{D}_{\text{top}}$ and $z \in S'_x$, deducing that $\widehat{\gamma}(x)$ and $w(x, Z)$ must be equal in $\widehat{\mathbb{L}}$ since

$$\begin{aligned} d_H \widehat{\Lambda}(\widehat{\beta}_k) &= d_H \widehat{\Lambda}(\widehat{\sigma}(\beta_{\mathfrak{p}k})) = d_H \Lambda(\beta_{\mathfrak{p}k}) \leq d_H(2\mathfrak{p}k + 1)dD < \frac{(2k + 1)D_Y^2 D_{YZ}}{\mathfrak{p}} \\ &< (2k + 1)D_Y^2 D_{YZ} < |S'_x|, \end{aligned}$$

Finally, having shown that $\widehat{\gamma} \in \mathbb{F}_q[X, Z]$, the arguments of Section 5.2.8 can be applied without any further changes, only with $\widehat{\gamma}$ in the role γ , concluding the proof.

D Proof of Lemma 4.4

The difference between what we have to prove and the original version in [Spi95, Lemma 4.2.18] is very small. The conclusion is the same, but the hypotheses are slightly different. The three hypotheses in the original version are

- (a) $\deg_X(A) \leq \deg_X(B)$,
- (b) $\deg_Z(A) \leq \deg_Z(B)$, and

$$(c) \quad \frac{\deg_X(B)}{n_X} + \frac{\deg_Z(B)}{n_Z} < 1.$$

Item (c) is the same as Item 3 in the hypothesis of Lemma 4.4. We will use hypotheses 1 and 2 of Lemma 4.4 to derive (b) and (a), respectively.

Let $A_0(Z), B_0(Z)$ be the leading coefficients of A, B when considered as polynomials in X with coefficients in $\mathbb{F}_q[Z]$. Since $n_Z > \deg_Z(A) + \deg_Z(B) \geq \deg(A_0) + \deg(B_0)$, there must be some $z \in \mathbb{F}_q$ for which $A(X, z)$ divides $B(X, z)$ but z is not a root of either A_0 nor B_0 . From $B_0(z) \neq 0$ we get that $B(X, z) \neq 0$, and then since $A_0(z) \neq 0$ and $A(X, z)$ divides $B(X, z)$ we find

$$\deg_X(A) = \deg(A(X, z)) \leq \deg(B(X, z)) \leq \deg_X(B)$$

as claimed.

Similarly $\deg_Z(A) \leq \deg_Z(B)$ follows from $\deg_X(A) + \deg_X(B) < n_X$.

Then [Spi95, Lemma 4.2.18] gives us the desired conclusion. □