

Circuit Lower Bounds, Help Functions, and the Remote Point Problem

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Abstract

We investigate the power of Algebraic Branching Programs (ABPs) augmented with help polynomials, and constant-depth Boolean circuits augmented with help functions. We relate the problem of proving explicit lower bounds in both these models to the Remote Point Problem (introduced in [3]). More precisely, proving lower bounds for ABPs with help polynomials is related to the Remote Point Problem w.r.t. the rank metric, and for constant-depth circuits with help functions it is related to the Remote Point Problem w.r.t. the Hamming metric. For algebraic branching programs with help polynomials with some degree restrictions we show exponential size lower bounds for explicit polynomials.

1 Introduction

The goal of circuit complexity, which is central to computational complexity, is proving lower bounds for explicit functions. The area has made several advances in the last three decades mainly for restricted circuit models. Some of the major results relating to circuit size lower bounds are the following: Exponential size lower bounds for constant-depth Boolean circuits [7, 12, 6] and for monotone Boolean circuits [2, 11] computing certain explicit Boolean functions; in the arithmetic circuit complexity setting, exponential size lower bounds for monotone arithmetic circuits [8] computing certain explicit polynomials, and exponential size lower bounds for explicit polynomials in the case of *noncommutative* algebraic branching programs [9]. More recently, [10] has shown superpolynomial lower bounds for multilinear arithmetic circuits. We can say that these restricted models of computation have been sufficiently well understood to show the nontrivial explicit lower bounds.

However, most of the central problems in the area continue to remain open. For example, we do not know how to prove superlinear size lower bounds for logarithmic depth Boolean circuits. We do not have superpolynomial size lower bounds for depth-3 arithmetic circuits over rationals.

The aim of this paper is to explore circuit complexity by augmenting the power of some of these restricted models by allowing *help functions* (in the arithmetic circuit case, *help polynomials*). In this paper we consider two specific problems.

1. Proving size lower bounds for constant depth Boolean circuits augmented with help functions. More precisely, given any set $\{h_1, h_2, \dots, h_m\}$ of help Boolean functions where

$$h_i : \{0, 1\}^n \longrightarrow \{0, 1\},$$

and m is (quasi)polynomial in n , we want to find an explicit Boolean function $f : \{0, 1\}^n \longrightarrow \{0, 1\}$ that requires superpolynomial size constant depth circuits C that takes as input $x_1, \dots, x_n, h_1, \dots, h_m$. The function f should be explicit in the sense that it is computable in $2^{n^{O(1)}}$ time.

2. Proving size lower bounds for *noncommutative* algebraic branching programs augmented with help polynomials. More precisely, given any set $\{h_1, h_2, \dots, h_m\}$ of help polynomials in the noncommuting variables $\{x_1, x_2, \dots, x_n\}$ over a field \mathbb{F} , we consider algebraic branching programs whose edges are labeled by \mathbb{F} -linear combinations of the h_i . The problem is to prove superpolynomial lower bounds for some explicit polynomial in x_1, \dots, x_n over \mathbb{F} .

We formally define explicit Boolean functions and explicit polynomials.

We say that a family of Boolean functions $\{f_n\}_{n>0}$, where $f_n : \{0, 1\}^n \longrightarrow \{0, 1\}$ for each n , is *explicit* if there is a uniform $2^{n^{O(1)}}$ time algorithm that takes $x \in \{0, 1\}^n$ as input and computes $f_n(x)$.

We say that a family of *multilinear* polynomials $\{P_n\}_{n>0}$ where $P_n(\bar{x}) \in \mathbb{F}[x_1, \dots, x_n]$ is *explicit* if there is a uniform $2^{n^{O(1)}}$ time algorithm that takes as input $(m, 0^n)$ for a multilinear monomial m (on indeterminates x_1, x_2, \dots, x_n) and outputs the coefficient of m in the polynomial P_n .

Contributions of this paper

For constant-depth circuits and noncommutative ABPs, augmented with help functions/polynomials respectively, proving lower bounds appears to be nontrivial.

1. We show that both the above lower bound problems are related to the *Remote Point Problem* studied by Alon et al [3]. For constant-depth circuits we show a connection to the *Remote Point Problem* in the Hamming metric studied in [3]. For noncommutative ABPs the problem is connected to the *Remote Point Problem* in the *rank metric* which is defined as the rank distance between matrices.
2. We also study the Remote Point Problem in the Rank metric, and we build on ideas from Alon et al's work (for the Hamming metric version) in [3] to give a deterministic

polynomial-time algorithm for certain parameters. However, these parameters are not sufficient to prove lower bounds for ABPs augmented with help polynomials. Similarly, the parameters achieved by the algorithm in [3] for the Hamming metric are not sufficient to prove explicit lower bounds for constant-depth circuits with help functions.

3. On the positive side, when the degrees of the help polynomials are somewhat restricted, using our solution to the Remote Point Problem w.r.t. the rank metric, we show exponential size lower bounds for noncommutative ABPs computing certain explicit polynomials (e.g. Theorem 14).

2 Constant Depth Circuits with Help Functions

In this section, we address the problem of proving lower bounds for constant depth circuits of polynomial size that have access to help functions $\{h_1, h_2, \dots, h_m\}$ at the input level. Our goal is to show how the problem is related to the Remote Point Problem w.r.t. the Hamming metric.

Notice that we can consider the circuit inputs x_1, x_2, \dots, x_n to be included in the set of help functions. Thus, we can assume that we consider constant depth circuits with input h_1, h_2, \dots, h_m and our goal is to prove superpolynomial lower bounds for such circuits. Notice that we cannot predetermine a hard Boolean function as the hard function chosen will depend on h_1, h_2, \dots, h_m .

It is well known that constant depth circuits can be well approximated by polylogarithmic degree polynomials, for different notions of approximation. We state the results of Tarui [13] (also see [4]) in the form that we require. In what follows, the field we work in will be \mathbb{F}_2 , but our results can be stated over any constant sized field, and over the rationals.

A polynomial $p(x_1, x_2, \dots, x_n, r_1, \dots, r_k)$ is called a *probabilistic polynomial* if it has as input the standard input bits x_1, x_2, \dots, x_n and, in addition, random input bits r_1, r_2, \dots, r_k . We say that the polynomial p represents a Boolean function $f : \{0, 1\}^n \rightarrow \{0, 1\}$ with error ϵ if

$$\begin{aligned} \text{Prob}[p(x_1, \dots, x_n, r_1, \dots, r_k) = f(x_1, \dots, x_n)] \\ \geq 1 - \epsilon, \end{aligned}$$

where the probability is over random choices of bits r_j .

Theorem 1. [13, 4] *There is a probabilistic polynomial $p(x_1, x_2, \dots, x_n, r_1, \dots, r_k)$ of degree $O(\log(1/\epsilon) \log^2 n)$ with $O(\log(1/\epsilon) \log^2 n)$ random bits that represents $OR(x_1, \dots, x_n)$ with error ϵ . Furthermore, $AND(x_1, \dots, x_n)$ can be similarly represented.*

Building on the above, the following well-known theorem is shown in [13, 4].

Theorem 2. [13, 4] *Every function f computed by a boolean circuit of depth d and size s is represented by a probabilistic polynomial $p(x_1, x_2, \dots, x_n, r_1, \dots, r_k)$ of degree $O(\log(1/\epsilon) \log^2 n)^d$ that represents $f(x_1, \dots, x_n)$ with error se .[†]*

[†]Tarui's construction yields a probabilistic polynomial q with integer coefficients. We can obtain the desired polynomial p over \mathbb{F}_2 from q by reducing the coefficients modulo 2.

Now, consider Boolean functions computed by constant-depth circuits with help functions. More precisely, let $H = \{h_1, h_2, \dots, h_m\}$ denote a set of Boolean help functions $h_i : \{0, 1\}^n \rightarrow \{0, 1\}$. For $s, d \in \mathbb{N}$, we define $\text{SizeDepth}_H(s, d)$ to be the set of Boolean functions $f : \{0, 1\}^n \rightarrow \{0, 1\}$ such that there is a depth d circuit C of size at most s such that

$$f(\bar{x}) = C(h_1(\bar{x}), h_2(\bar{x}), \dots, h_m(\bar{x})),$$

where \bar{x} denotes the n -tuple (x_1, x_2, \dots, x_n) . The lower bound problem is to construct, for each fixed d , and for any given set of help functions H and $s \in \mathbb{N}$, an explicit Boolean function g such that g is *not* in $\text{SizeDepth}_H(s, d)$.

We do not have a solution to this problem. However, we show that this lower bound problem is connected to the Remote Point Problem (RPP) introduced by Alon et al [3]. An interesting deterministic algorithm for RPP is presented in [3]. A deterministic algorithm with somewhat stronger parameters would solve our lower bound question. We now explain this connection.

The Remote Point Problem (RPP) [3]. Given a k -dimensional subspace $V \subseteq \mathbb{F}_2^N$ the problem is to find a vector $v \in \mathbb{F}_2^N$ such that the Hamming distance $d(u, v) \geq r$ for every $u \in V$ if it exists. We will call an efficient algorithm that does this an (N, k, r) -solution to the problem.

The challenge is to give an efficient deterministic algorithm for RPP. A randomized algorithm that simply picks v at random would be a good solution with high probability (for most parameters k and r of interest). Alon et al in [3] give an (N, k, r) solution for $r = O\left(\frac{N \log k}{k}\right)$, where their deterministic algorithm runs in time polynomial in N . We now state and prove the connection between RPP and our lower bound question.

Theorem 3. *Let $N = 2^n$. For any constant $d \in \mathbb{N}$, and any constants $c_0 > c_1 > c_2 > 0$ such that $c_0 > (c_1 + 2c_2)d + c_2$, if the Remote Point Problem with parameters (N, k, r) – for $k = 2^{(\log n)^{c_0}}$ and $r = \frac{N}{2^{(\log n)^{c_1}}}$ – can be solved in time $2^{n^{O(1)}}$, then, for any given set of help functions H such that $|H| = 2^{(\log n)^{c_2}}$ and $s = cn^c$, there is an explicit Boolean function that does not belong to $\text{SizeDepth}_H(s, d)$ for large enough n (depending on c).*

Proof. The proof is an easy application of Theorem 2. Let $H = \{h_1, h_2, \dots, h_m\}$. Consider a circuit C corresponding to the class $\text{SizeDepth}_H(s, d)$. To wit, the function it computes is $C(h_1(\bar{x}), h_2(\bar{x}), \dots, h_m(\bar{x}))$, where C is depth- d , unbounded fanin and of size cm^c . Now, for \bar{x} picked uniformly at random from $\{0, 1\}^n$ suppose the probability distribution of $(h_1(\bar{x}), h_2(\bar{x}), \dots, h_m(\bar{x}))$ on the set $\{0, 1\}^m$ is μ . By Theorem 2 there is a probabilistic polynomial $p(y_1, y_2, \dots, y_m, r_1, r_2, \dots, r_t)$ of degree $O(\log(1/\epsilon) \log^2 m)^d$ that represents $C(y_1, y_2, \dots, y_m)$ with error $cm^c \epsilon$. By a standard averaging argument it follows that we can fix the random bits r_1, r_2, \dots, r_t to get

$$\begin{aligned} \text{Prob}_{\mu}[p(y_1, y_2, \dots, y_m, r_1, r_2, \dots, r_t) = \\ C(y_1, y_2, \dots, y_m)] \geq 1 - cm^c \epsilon, \end{aligned}$$

where (y_1, y_2, \dots, y_m) is picked according to distribution μ . But that is equivalent to

$$\begin{aligned} \text{Prob}[p(h_1(\bar{x}), \dots, h_m(\bar{x}), r_1, r_2, \dots, r_t) = \\ C(h_1(\bar{x}), h_2(\bar{x}), \dots, h_m(\bar{x}))] \geq 1 - cm^c \epsilon, \end{aligned} \tag{1}$$

where \bar{x} is picked uniformly at random from $\{0, 1\}^n$.

Choose $c'_0 < c_0 - c_2$ and $c'_1 > c_1 (> c_2)$ such that $c'_0 = (c'_1 + 2c_2)d$. Let $\epsilon = \frac{1}{2^{(\log n)^{c'_1}}}$. Then the degree of p above is $O(\log n)^{c'_0}$. We will consider Boolean functions on n bits as vectors in \mathbb{F}_2^N . Let V be the subspace in \mathbb{F}_2^N spanned by all monomials (i.e, products of help functions) of degree at most $O(\log n)^{c'_0}$. Then the dimension k of V is $m^{O(\log n)^{c'_0}} < 2^{(\log n)^{c_0}}$. By Inequality (1), it follows that finding a vector $v \in \mathbb{F}_2^N$ that is r -far from V for $r = \frac{N}{2^{(\log n)^{c_1}}} > cNm^c \epsilon$ in time $2^{n^{O(1)}}$ would give us an explicit Boolean function that is not in $\text{SizeDepth}_H(s, d)$. \square

Remark 4. We recall a nice related result of Jin-Yi Cai: He has shown in [5] an exponential lower bound for the size of constant-depth circuits that computes m specific parities in the presence of (any) $m - 1$ help functions, where $m \leq n^{1/5}$. His proof is essentially based on Smolensky's dimension argument [12]. However, in our setting where we allow for polynomially many help functions Smolensky's argument [12] does not work.

We now state an interesting connection between explicit lower bounds against $\text{SizeDepth}_H(n^c, d)$ and lower bounds against the polynomial time many-one closure of AC^0 . The proof proceeds by a simple diagonalization argument. For any complexity class \mathcal{C} , let $\mathcal{R}_m^p(\mathcal{C})$ denote the polynomial-time many-one closure of \mathcal{C} , i.e, the class of languages that can be reduced in polynomial time to a language in \mathcal{C} .

Theorem 5. Suppose, for every fixed $d \in \mathbb{N}$, there is a $2^{n^{O(1)}}$ time algorithm \mathcal{A} that takes as input a set of help functions $H = \{h_i : \{0, 1\}^n \rightarrow \{0, 1\} \mid i \in [m]\}$ where $m \leq n^{\log n}$ ‡ (where each h_i is given by its truth-table), and \mathcal{A} outputs the truth-table of a Boolean function $g : \{0, 1\}^n \rightarrow \{0, 1\}$ such that for any $c > 0$, $g \notin \text{SizeDepth}_H(n^c, d)$ for almost all n . Then $\text{EXP} \not\subseteq \mathcal{R}_m^p(\text{AC}^0)$.

Proof. For any $d \in \mathbb{N}$, let AC_d^0 denote the class of languages that are accepted by polynomial-sized circuit families of polynomial size and depth d .

Note that to prove that $\text{EXP} \not\subseteq \mathcal{R}_m^p(\text{AC}^0)$, it suffices to prove that $\text{EXP} \not\subseteq \mathcal{R}_m^p(\text{AC}_d^0)$ for each fixed $d \in \mathbb{N}$, since EXP contains problems that are complete for it under polynomial-time many-one reductions. We will now describe, for any fixed $d \in \mathbb{N}$, an EXP machine that accepts a language $L_d \notin \mathcal{R}_m^p(\text{AC}_d^0)$.

We proceed by diagonalization. Let R_1, R_2, R_3, \dots be any standard enumeration of all polynomial-time many-one reductions such that each reduction appears infinitely often in

‡Here, $\log n$ can be replaced by any function $f : \mathbb{N} \rightarrow \mathbb{N}$ such that $f(n)$ is $2^{n^{O(1)}}$ -time computable, $f(n) = \omega(1)$, and $f(n) \leq n^{O(1)}$.

the list. Fix $n \in \mathbb{N}$ and let $m = \max_{y \in \{0,1\}^n} |R_n(y)|$. On an input $x \in \{0,1\}^n$, the EXP machine does the following: for each $y \in \{0,1\}^n$, it runs R_n for $n^{\log n}$ time and computes $R_n(y)$ (if R_n does not halt in time $n^{\log n}$, the machine outputs 0 and halts). It can thus produce the truth tables of functions $h_i : \{0,1\}^n \rightarrow \{0,1\}$ ($i \in [m]$) such that for each $y \in \{0,1\}^n$, $h_i(y)$ is the i th bit of $R_n(y)$ if $|R_n(y)| \geq i$ and 0 otherwise. Now, by assumption, in time $2^{n^{O(1)}}$, the EXP machine can compute the truth table of a function $g_n : \{0,1\}^n \rightarrow \{0,1\}$ such that, for any $c > 0$, $g_n \notin \text{SizeDepth}_{\{h_1, \dots, h_m\}}(n^c, d)$ for large enough n . Having computed g_n , the EXP machine just outputs $g_n(x)$.

It is clear, by a standard argument, that L_d cannot be polynomial-time many-one reduced to any language in AC_d^0 . \square

3 Noncommutative Algebraic Branching Programs

Let $X = \{x_1, x_2, \dots, x_n\}$ be a set of n noncommuting variables, and $\mathbb{F}\langle X \rangle$ denote the noncommutative ring of polynomials over X with coefficients from the field \mathbb{F} . For $f \in \mathbb{F}\langle X \rangle$, let $d(f)$ denote the degree of f . Let $\text{Mon}_d(X)$ be the set of degree d monomials over X . For a polynomial f and a monomial m over X , let $f(m)$ denote the coefficient of m in f . A nonempty subset $H \subseteq \mathbb{F}\langle X \rangle$ is *homogeneous* if there is a $d \in \mathbb{N}$ such that all the polynomials in H are homogeneous of degree d .

Let $G = (V, E)$ be a directed acyclic graph. For $u, v \in V$, let $\mathcal{P}_{u,v}$ be the set of paths from u to v , where a path in $\mathcal{P}_{u,v}$ is a tuple of the form $((u_0, u_1), (u_1, u_2), \dots, (u_{l-1}, u_l))$ where $u_0 = u$ and $u_l = v$.

Definition 6. Let $X = \{x_1, x_2, \dots, x_n\}$ and $Y = \{y_1, y_2, \dots, y_m\}$ be disjoint variable sets. Let $H = \{h_1, h_2, \dots, h_m\} \subseteq \mathbb{F}\langle X \rangle$. An Algebraic Branching Program (ABP) with help polynomials H is a layered directed acyclic graph A with a source s and a sink t . Every edge e of A is labeled by a linear form $L(e)$ in variables $X \cup Y$. If $L(e) = \sum_i \alpha_i x_i + \sum_j \beta_j y_j$, the polynomial $L'(e)$ associated with edge e is obtained by substituting h_j for y_j , $1 \leq j \leq m$, in $L(e)$. I.e. $L'(e) = \sum_i \alpha_i x_i + \sum_j \beta_j h_j$. The size of A is the number of vertices in A .

Given a path $\gamma = (e_1, e_2, \dots, e_l)$ in A , define the polynomial $f_\gamma = L'(e_1) \cdot L'(e_2) \cdot \dots \cdot L'(e_l)$ (note that the order of multiplication is important). For vertices u and v of A , we define the polynomial $f_{u,v} = \sum_{\gamma \in \mathcal{P}_{u,v}} f_\gamma$. The ABP A computes the polynomial $f_{s,t}$.

Suppose $L(e) = \sum_i \alpha_i x_i + \sum_j \beta_j y_j$. We say that the edge e is *homogeneously labeled* if all the polynomials in the set $\{x_i \mid \alpha_i \neq 0\} \cup \{h_j \mid \beta_j \neq 0\}$ are homogeneous and of the same degree $d(e)$. If the above set is empty, we let $d(e) = 0$. Now, suppose all edges of an ABP A are homogeneously labeled; then, for a path $\gamma = (e_1, e_2, \dots, e_t)$ in A let $d(\gamma) = \sum_{i=1}^t d(e_i)$. The ABP A with help polynomials H is *homogeneous* if:

- all the edges in A are homogeneously labeled,
- For all u, v in A and $\gamma_1, \gamma_2 \in \mathcal{P}_{u,v}$, $d(\gamma_1) = d(\gamma_2)$.

For a homogeneous ABP A with help polynomials and any pair of vertices u, v in A , the polynomial computed from u to v is homogeneous.

In the absence of help polynomials, this gives the standard Algebraic Branching Programs as defined in, e.g. Nisan [9]. Nisan [9] has shown explicit lower bounds, e.g. for the Permanent and Determinant, for this model of computation. Our aim is to prove lower bounds for ABPs with help polynomials.

We show that any ABP with arbitrary help polynomials computing a homogeneous polynomial can be transformed into an equivalent *homogeneous* ABP with homogeneous help polynomials with only a small increase in size. Thus, it suffices to prove lower bounds against homogeneous ABPs with help polynomials. Fix the help polynomial set $H \subseteq \mathbb{F}\langle X \rangle$. Let $m = |H|$ and $d(H) = \max_{h \in H} d(h)$. Also, fix some $d \in \mathbb{N}$.

Given $f \in \mathbb{F}\langle X \rangle$ and $i \in \mathbb{N}$, let $f^{(i)}$ denote the i th homogeneous part of h . For $2 \leq i \leq d$, let $\tilde{H}_i = \{h^{(i)} \in \mathbb{F}\langle X \rangle \mid h \in H\}$; let $\tilde{H} = \bigcup_{2 \leq i \leq d} \tilde{H}_i$. Let \tilde{m}_i denote $|\tilde{H}_i|$ for each i , and let \tilde{m} denote $|\tilde{H}| = \sum_i \tilde{m}_i$. We show the following homogenization theorem.

Theorem 7. *Given any ABP A using the help polynomials H computing a homogeneous polynomial f of degree $d \geq 1$, there is a homogeneous ABP \tilde{A} using the help polynomials \tilde{H} that computes the same polynomial as A , where the size of \tilde{A} is at most $S(d+1)$, where S denotes the size of A .*

Proof. The following construction is fairly standard. Let s and t be the designated source and sink, respectively, of the ABP. We will use the notation of Section 3.

We now define \tilde{A} . \tilde{A} will use the variables $X \cup \tilde{Y}$, where $\tilde{Y} = \{y_i^{(j)} \mid 1 \leq i \leq m, 2 \leq j \leq d(h_i)\}$. The vertices of \tilde{A} are tuples (u, i) , where u is a vertex of A and $i \in \mathbb{N}$ is a number between 0 and d . The source of \tilde{A} will be $(s, 0)$ and the sink (t, d) . We will define the set of edges of \tilde{A} in two stages. We will first construct an ABP on the set of vertices of \tilde{A} which will include edges with weights from \mathbb{F} (i.e, edges e such that $L(e)$ is a non-zero degree 0 polynomial), and we will then show how to remove these edges from the ABP. Consider any edge e in the ABP A ; let the label $L(e)$ of e be $\sum_{i=1}^n \alpha_i x_i + \sum_{j=1}^m \beta_j y_j$ and $0 \leq k \leq d$, define the linear form $L(e)_k$ – which captures the k th homogeneous part of $L'(e)$, the polynomial computed by edge e – as follows:

- If $k = 0$, define $L(e)_k$ to be the field element $\sum_{j=1}^m \beta_j h_j^{(0)}$.
- If $k = 1$, define $L(e)_k$ to be $\sum_{i=1}^n \alpha_i x_i + \sum_{j=1}^m \beta_j h_j^{(1)}$.
- If $k > 1$, define $L(e)_k$ to be $\sum_{j=1}^m \beta_j y_j^{(k)}$

Fix any vertex (v, k) of \tilde{A} . Let $\{u_1, u_2, \dots, u_l\}$ be the predecessors of v in A and let e_i denote the edge (u_i, v) . Then, it is easy to see that

$$f_{s,v}^{(k)} = \sum_{i=1}^l \sum_{j=0}^k f_{s,u_i}^{(j)} L'(e_i)_{(k-j)}$$

Hence, we define edges $e_{i,j}$ in \tilde{A} from vertices (u_i, j) to (v, k) with label $L(e_{i,j}) = L(e_i)_{k-j}$. (Note that the label $L(e_{i,k})$ is just a field element. We will change this presently.) This concludes the first stage. Note that, since we only add edges from (u, i) to (v, j) when (u, v) is an edge in A , the graph of \tilde{A} is acyclic. Also note that an edge e is labeled by a field element if and only if it connects vertices of the form (u, k) and (v, k) , for some u, v , and k . Finally, it is easily seen from the definition of \tilde{A} that the polynomial computed from $(s, 0)$ to (u, i) is the polynomial $f_{s,u}^{(i)}$ for any s, u , and i .

In the second stage, we will get rid of those edges in \tilde{A} such that $L(e) \in \mathbb{F}$. We do this in two passes. Fix some topological ordering of the vertices of \tilde{A} , and order the edges (\tilde{u}, \tilde{v}) of \tilde{A} lexicographically. As long as there is an edge $e = (\tilde{u}, \tilde{v})$ of \tilde{A} such that \tilde{v} is *not* the designated sink (t, d) and $L(e) \in \mathbb{F}$, we let e be the least such edge and do the following: we remove the edge e , and for each edge $e' = (\tilde{v}, \tilde{w})$ of \tilde{A} going out of \tilde{v} , we change the label of the edge $e'' = (\tilde{u}, \tilde{w})$ to $L(e'') + L(e) \cdot L(e')$ (if no such edge e'' exists, we add this edge to the ABP and give it the label $L(e) \cdot L(e')$). It should be clear that the homogeneity of the ABP is preserved. After at most $O((sd)^2)$ many such modifications, all edges in \tilde{A} that are labeled by field elements are of the form $(\tilde{u}, (t, d))$. Moreover, by the above construction, it is clear that $\tilde{u} = (u, d)$ for some vertex $u \neq t$ of A . Since $d \geq 1$, we know that $\tilde{u} \neq (s, 0)$, the designated source node. We also know that there are no edges into \tilde{u} which are labeled by a field element. We now do the following: for each edge $e = (\tilde{u}, (t, d))$ labeled by a field element, we remove the vertex \tilde{u} and for each edge $e' = (\tilde{v}, \tilde{u})$, we remove e' and change the label of $e'' = (\tilde{v}, (t, d))$ to $L(e'') + L(e') \cdot L(e)$ (if no such e'' exists, we add such an edge e'' and set its label to $L(e') \cdot L(e)$). This concludes the construction.

It is easy to prove inductively that after every modification of \tilde{A} , the polynomial computed from $(s, 0)$ to (t, d) remains $f_{s,t}^{(d)}$. Hence, the ABP \tilde{A} computes exactly the polynomial f computed by A . Also, by construction, the edges of \tilde{A} are all homogeneously labeled; finally, it can also be seen that given a path γ from vertex (u, i) to vertex (v, j) in \tilde{A} , $d(\gamma) = j - i$: hence, the ABP is indeed homogeneous, and we are done. \square

4 Decomposition of Communication Matrices

We now generalize the key lemma of Nisan [9] that connects the size of noncommutative ABPs for an $f \in \mathbb{F}\langle X \rangle$ to the ranks of certain communication matrices $M_k(f)$. The generalization is for noncommutative ABPs with help polynomials, and it gives a more complicated connection between the size of ABPs to the ranks of certain matrices. For usual noncommutative ABPs considered in [9], Nisan's lemma directly yields the lower bounds. In our case, this generalization allows us to formulate the lower bound problem as a Remote Point Problem for the rank metric.

We will assume that the explicit polynomial for which we will be proving lower bounds is homogeneous. Thus, by Theorem 7 we can assume that each help polynomial in $H = \{h_1, h_2, \dots, h_m\}$ is homogeneous and of degree at least 2.

We first fix some notation. Let $d \in \mathbb{N}$ be an even number. Let $d(H) = \max_{h \in H} d(h)$. Also,

for $2 \leq i \leq d(H)$, let $H_i = \{h \in H \mid d(h) = i\}$.

Suppose $f \in \mathbb{F}\langle X \rangle$ is homogeneous of even degree $d \geq 2$, and $k \in \mathbb{N}$ such that $0 \leq k \leq d$. We define the $n^k \times n^{d-k}$ matrix $M_k(f)$ (as in [9]): Each row is labeled by a distinct monomial in $\text{Mon}_k(X)$ and each column by a distinct monomial in $\text{Mon}_{d-k}(X)$. Given monomials $m_1 \in \text{Mon}_k(X)$ and $m_2 \in \text{Mon}_{d-k}(X)$, the (m_1, m_2) th entry of $M_k(f)$ is the coefficient of the monomial $m_1 m_2$ in f and is denoted by $M_k(f)(m_1, m_2)$.

Call M an (l, m) -matrix if M is an $n^l \times n^m$ matrix with entries from \mathbb{F} , where the rows of M are labeled by monomials in $\text{Mon}_l(X)$ and columns by monomials in $\text{Mon}_m(X)$. Suppose $0 \leq l \leq k$ and $0 \leq m \leq d - k$. Let M_1 be an (l, m) -matrix and M_2 a $(k - l, (d - k) - m)$ -matrix. We define the $(k, d - k)$ -matrix $M = M_1 \otimes_{l, m}^k M_2$ as follows: Suppose $m_1 \in \text{Mon}_k(X)$ and $m_2 \in \text{Mon}_{d-k}(X)$ are monomials such that $m_1 = m_{11} m_{12}$ with $m_{11} \in \text{Mon}_{k-l}(X)$ and $m_{12} \in \text{Mon}_l(X)$ and $m_2 = m_{21} m_{22}$ with $m_{21} \in \text{Mon}_m(X)$ and $m_{22} \in \text{Mon}_{(d-k)-m}(X)$. Then the $(m_1, m_2)^{\text{th}}$ entry of M is defined as

$$M(m_1, m_2) = M_1(m_{12}, m_{21}) \cdot M_2(m_{11}, m_{22}).$$

Let A be a homogeneous ABP with help polynomials H computing a polynomial f of degree d . Let u, v and w be vertices in the ABP A , and $\gamma_1 \in \mathcal{P}_{u, v}$ and $\gamma_2 \in \mathcal{P}_{v, w}$ be paths. We denote by $\gamma_1 \circ \gamma_2 \in \mathcal{P}_{u, w}$ the concatenation of γ_1 and γ_2 .

Since A is homogeneous, each of the polynomials $f_{u, v}$ for vertices u, v of A is homogeneous. For $1 \leq k \leq d/2$, define the k -cut of A , $C_k \subseteq V(A) \cup E(A)$, as follows: A vertex $v \in V(A)$ is in C_k iff $d(f_{s, v}) = k$, and an edge $e = (u, v) \in E(A)$ is in C_k iff $d(f_{s, u}) < k$ and $d(f_{s, v}) > k$. For each $x \in C_k$, let \mathcal{P}_x denote the set of s - t paths passing through x . Clearly, the sets $\{\mathcal{P}_x \mid x \in C_k\}$ partition $\mathcal{P}_{s, t}$, the set of all paths from s to t . Thus, we have

$$\begin{aligned} f &= \sum_{x \in C_k} \sum_{\gamma \in \mathcal{P}_x} f_\gamma \\ &= \sum_{v \in C_k \cap V(A)} \sum_{\gamma \in \mathcal{P}_v} f_\gamma + \sum_{e \in C_k \cap E(A)} \sum_{\gamma \in \mathcal{P}_e} f_\gamma. \end{aligned} \tag{2}$$

We now analyze Equation 2. For $v \in C_k \cap V(A)$, $\mathcal{P}_v = \{\gamma_1 \circ \gamma_2 \mid \gamma_1 \in \mathcal{P}_{s, v}, \gamma_2 \in \mathcal{P}_{v, t}\}$. Hence, for any $v \in C_k \cap V(A)$:

$$\begin{aligned} \sum_{\gamma \in \mathcal{P}_v} f_\gamma &= \sum_{\substack{\gamma_1 \in \mathcal{P}_{s, v} \\ \gamma_2 \in \mathcal{P}_{v, t}}} f_{\gamma_1 \circ \gamma_2} = \sum_{\substack{\gamma_1 \in \mathcal{P}_{s, v} \\ \gamma_2 \in \mathcal{P}_{v, t}}} f_{\gamma_1} \cdot f_{\gamma_2} \\ &= f_{s, v} f_{v, t}. \end{aligned} \tag{3}$$

Similarly, for any edge $e = (u, v) \in C_k \cap E(A)$, $\mathcal{P}_e = \{\gamma_1 \circ (e) \circ \gamma_2 \mid \gamma_1 \in \mathcal{P}_{s, u}, \gamma_2 \in \mathcal{P}_{v, t}\}$,

where (e) denotes the path containing just the edge e . Thus,

$$\begin{aligned}
\sum_{\gamma \in \mathcal{P}_e} f_\gamma &= \sum_{\substack{\gamma_1 \in \mathcal{P}_{s,u} \\ \gamma_2 \in \mathcal{P}_{v,t}}} f_{\gamma_1 \circ (e) \circ \gamma_2} \\
&= \sum_{\substack{\gamma_1 \in \mathcal{P}_{s,v} \\ \gamma_2 \in \mathcal{P}_{v,t}}} f_{\gamma_1} \cdot L'(e) \cdot f_{\gamma_2} \\
&= f_{s,u} L'(e) f_{v,t}.
\end{aligned} \tag{4}$$

From Equations 2, 3, and 4, we get

$$\begin{aligned}
f &= \sum_{v \in C_k \cap V(A)} f_{s,v} f_{v,t} + \\
&\quad \sum_{e=(u,v) \in C_k \cap E(A)} f_{s,u} L'(e) f_{v,t}.
\end{aligned}$$

As A is homogeneous of degree d , each polynomial in the sums above is homogeneous of degree d . Hence

$$\begin{aligned}
M_k(f) &= \sum_{v \in C_k \cap V(A)} M_k(f_{s,v} f_{v,t}) + \\
&\quad \sum_{e=(u,v) \in C_k \cap E(A)} M_k(f_{s,u} L'(e) f_{v,t}).
\end{aligned} \tag{6}$$

For any $v \in C_k \cap V(A)$, $f_{s,v}$ and $f_{v,t}$ are homogeneous degree k and $d - k$ polynomials respectively. We denote by M_v the matrix $M_k(f_{s,v} f_{v,t})$. Notice that for $m_1 \in \text{Mon}_k(X)$ and $m_2 \in \text{Mon}_{d-k}(X)$, the $(m_1, m_2)^{th}$ entry of the matrix $M_v = M_k(f_{s,v} f_{v,t})$ is $f_{s,v}(m_1) f_{v,t}(m_2)$. Thus, M_v is an outer product of two column vectors and is hence a matrix of rank at most 1. Therefore, the first summation in Equation 6 is a matrix of rank at most $|C_k \cap V(A)|$.

For $e = (u, v) \in C_k \cap E(A)$, we know that $d(f_{s,u}) < k$ and $d(f_{s,v}) > k$ and thus, $d(e) \geq 2$. Hence, $L'(e) = \sum_{h \in H_{d(e)}} \beta_{e,h} h$, for $\beta_{e,h} \in \mathbb{F}$. Therefore, expanding the second summation in Equation 6, we get

$$\begin{aligned}
\sum_{\substack{e=(u,v) \in \\ C_k \cap E(A)}} M_k(f_{s,u} L'(e) f_{v,t}) &= \\
&\quad \sum_{\substack{e=(u,v) \in \\ C_k \cap E(A)}} \sum_{h \in H_{d(e)}} \beta_{e,h} M_k(f_{s,u} \cdot h \cdot f_{v,t})
\end{aligned} \tag{7}$$

Consider a term of the form $M_k(f_{s,u} h f_{v,t})$. For the rest of the proof let $d(w)$ denote $d(f_{s,w})$, for any vertex w of A . Given monomials $m_{11} \in \text{Mon}_{d(u)}(X)$, $m_{12} \in \text{Mon}_{k-d(u)}(X)$, $m_{21} \in$

$\text{Mon}_{d(h)-(k-d(u))}(X)$, and $m_{22} \in \text{Mon}_{d-d(v)}(X)$, the entry $M_k(f_{s,u}h f_{v,t})(m_{11}m_{12}, m_{21}m_{22}) = h(m_{12}m_{21})f_{s,u}(m_{11})f_{v,t}(m_{22})$, since all polynomials involved are homogeneous. Hence, the matrix $M_k(f_{s,u}h f_{v,t})$ is precisely $M_{k-d(u)}(h) \otimes_{k-d(u), d(h)-(k-d(u))}^k M_e$, where $M_e(m_{11}, m_{22}) = f_{s,u}(m_{11})f_{v,t}(m_{22})$, for $m_{11} \in \text{Mon}_{d(u)}(X)$, $m_{22} \in \text{Mon}_{d-d(v)}(X)$. Clearly, M_e is a matrix of rank at most 1, for any $e \in C_k \cap E(A)$ and $h \in H_{d(e)}$. Continuing with the above calculation, we get

$$\begin{aligned} & \sum_{\substack{e=(u,v) \in \\ C_k \cap E(A)}} M_k(f_{s,u}L'(e)f_{v,t}) \\ &= \sum_{\substack{e=(u,v) \in \\ C_k \cap E(A)}} \sum_{h \in H_{d(e)}} \beta_{e,h} M_{l_e}(h) \otimes_{l_e, m_e}^k M_e \\ &= \sum_{h \in H} \sum_{i=d_1(h)}^{d_2(h)} M_i(h) \otimes_{i, d(h)-i}^k \cdot \sum_{\substack{e=(u,v) \in C_k: \\ d(e)=d(h) \\ d(u)=k-i}} \beta_{e,h} M_e, \end{aligned}$$

where $d_1(h) = \max\{1, d(h) - (d - k)\}$, $d_2(h) = \min\{d(h) - 1, k\}$, $l_e = k - d(u)$, and $m_e = d(h) - (k - d(u))$.

Plugging the above observations into Equation 6, we have

$$\begin{aligned} M_k(f) &= \underbrace{\left(\sum_{v \in C_k \cap V(A)} M_v \right)}_{M'} + \\ & \sum_{h \in H} \sum_{i=d_1(h)}^{d_2(h)} M_i(h) \otimes_{i, d(h)-i}^k \underbrace{\left(\sum_{\substack{e=(u,v) \in C_k: \\ d(e)=d(h) \\ d(u)=k-i}} \beta_{e,h} M_e \right)}_{M'_{i,h}} \end{aligned}$$

Notice that M' above has rank at most $|V(A)|$, and $M'_{i,h}$ has rank at most $|E(A)| \leq |V(A)|^2$ for any $h \in H$ and $d_1(h) \leq i \leq d_2(h)$. Hence, we have proved the following result:

Theorem 8. *Let A be a homogeneous ABP of size S computing a (homogeneous) polynomial f of degree d using the help polynomials H . Then, for any $k \in \{0, 1, \dots, d\}$, we can write $M_k(f)$ as:*

$$M_k(f) = M' + \sum_{h \in H} \sum_{i=d_1(h)}^{d_2(h)} M_i(h) \otimes_{i, d(h)-i}^k M'_{i,h},$$

where $d_1(h) = \max\{1, d(h) - (d - k)\}$ and $d_2(h) = \min\{d(h) - 1, k\}$ such that $\text{rank}(M') \leq S$ and $\text{rank}(M'_{i,h}) \leq S^2$ for each $h \in H$, and $i \in \{\max\{1, d(h) - (d - k)\}, \dots, \min\{d(h) - 1, k\}\}$.

5 Remote Point Problem for the rank metric

We now introduce an algorithmic problem that will help us prove lower bounds on the sizes of ABPs computing explicit polynomials using a (given) set of help polynomials H . This problem is actually the *Remote Point Problem for matrices in the rank metric* that we denote RMP. This problem is analogous to the Remote Point Problem (RPP), which we discussed in Section 2.

Given two matrices $P, Q \in \mathbb{F}^{a \times b}$, the *Rank distance* between P and Q is defined to be $\text{rank}(P - Q)$. It is known that this defines a metric, known as the *rank metric* on the set of all $a \times b$ matrices over \mathbb{F} .

The RMP problem. Given as input a set of $N \times N$ matrices P_1, P_2, \dots, P_k over a field \mathbb{F} and $r \in \mathbb{N}$, the problem is to compute an $N \times N$ matrix P such that for any matrix $P' = \sum_{i=1}^k \alpha_i P_i$ in the subspace generated by P_1, P_2, \dots, P_k , the rank distance between P and P' is at least r .

In the problem N is taken as the input size, and k and r are usually functions of N . We say that the RMP problem has an (N, k, r) -solution over \mathbb{F} if there is a *deterministic* algorithm that runs in time polynomial in N and computes a matrix P that is at rank distance at least r from the subspace generated by the P_1, P_2, \dots, P_k .

Remark 9. *How does a solution to RMP give us an explicit noncommutative polynomial f for which we can show lower bounds for the sizes of noncommutative ABPs with help polynomials? We now explain the connection.*

Let A be a homogeneous ABP of size S computing a polynomial f of degree d . Let $d_1(h)$ denote $\max\{1, d(h) - d/2\}$ and $d_2(h)$ denote $\min\{d/2, d(h) - 1\}$. For $a, b, p, q \in \mathbb{N}$ such that $p \in [n^a]$ and $q \in [n^b]$, let $E_{a,b}^{p,q}$ be the $n^a \times n^b$ elementary matrix with 1 as (p, q) th entry, and 0 elsewhere. The matrices $\{E_{a,b}^{p,q} \mid p \in [n^a], q \in [n^b]\}$ span all matrices in $\mathbb{F}^{n^a \times n^b}$. By Theorem 8

$$M_{d/2}(f) = M' + \sum_{h \in H} \sum_{i=d_1(h)}^{d_2(h)} M_i(h) \otimes_{i, d(h)-i}^{d/2} M'_{i,h},$$

where $\text{rank}(M') \leq S$. For $h \in H$ and $i \in \{d_1(h), \dots, d_2(h)\}$, the matrix $M'_{i,h}$ is an $n^{d/2-i} \times n^{d/2-d(h)+i}$ dimension matrix. We can write $M'_{i,h}$ as a linear combination of the elementary matrices in $\{E_{d/2-i, d/2-d(h)+i}^{p,q} \mid p \in [n^{d/2-i}], q \in [n^{d/2-d(h)+i}]\}$.

Let \mathcal{A} be the set of matrices of the form $M_i(h) \otimes_{i, d(h)-i}^{d/2} E_{d/2-i, d/2-d(h)+i}^{p,q}$ where $h \in H$, $i \in \{d_1(h), \dots, d_2(h)\}$, and $p \in [n^{d/2-i}], q \in [n^{d/2-d(h)+i}]$. Each matrix in \mathcal{A} is an $n^{d/2} \times n^{d/2}$ matrix, with its rows and columns labeled by monomials in $\text{Mon}_{d/2}(X)$. Every matrix of the form $M_i(h) \otimes_{i, d(h)-i}^{d/2} M'_{i,h}$ is a linear combination of matrices in \mathcal{A} . Crucially, note that \mathcal{A} depends only on the set of help polynomials and the parameter d , and it does not depend on the ABP A .

By substitution for $M'_{i,h}$ we obtain the following expression for $M_{d/2}(f)$ in terms of linear

combination of matrices in \mathcal{A} .

$$M_{d/2}(f) = M' + \sum_{M \in \mathcal{A}} \alpha_M M,$$

where $\alpha_M \in \mathbb{F}$. Since, M' has rank at most S , it implies that $M_{d/2}(f)$ is at rank distance at most S from the subspace generated by the matrices in \mathcal{A} . Thus, if we can compute a matrix \hat{M} in deterministic time polynomial in n^d that has rank distance $S = 2^{O(n)}$ from the subspace generated by \mathcal{A} we would obtain an explicit homogeneous degree d polynomial f with lower bound $2^{\Omega(n)}$ by setting $\hat{M} = M_{d/2}(f)$. This is the approach that we will take for proving lower bounds.

We present the following simple algorithm, which suffices for our lower bound application.

Theorem 10. *For any k , the RMP has an $(N, k, \lfloor N/k + 1 \rfloor)$ -solution over any field \mathbb{F} such that field operations in \mathbb{F} and Gaussian elimination over \mathbb{F} can be performed in polynomial time.*

Proof. We assume that $k < N$; otherwise the problem is trivial. Let r denote $\lfloor N/k + 1 \rfloor$. Choose the first r column vectors in each of the matrices P_1, P_2, \dots, P_k . Let $v_1, v_2, \dots, v_{rk} \in \mathbb{F}^N$ be these vectors in some order. As $rk \leq N - r$, using Gaussian elimination, we can efficiently choose $v_{rk+1}, v_{rk+2}, \dots, v_{r(k+1)} \in \mathbb{F}^N$ with the following property: for every $i \in [k + 1]$, v_{rk+i} is linearly independent of $v_1, v_2, \dots, v_{r(k+i-1)}$. Let P be any matrix that has $v_{rk+1}, v_{rk+2}, \dots, v_{r(k+1)}$ as its first r columns. It is not too difficult to see that given any matrix P' in the subspace generated by P_1, P_2, \dots, P_k , the first r columns of $P - P'$ remain independent, i.e $\text{rank}(P - P') \geq r$. \square

Remark 11. *The Remote Point Problem is fascinating as an algorithmic question. In [3] Alon et al provide a nontrivial algorithm for RPP in the Hamming metric (over \mathbb{F}_2). We use similar methods to provide an improved solution to RMP for small prime fields. The result is proved in Section 7. Unfortunately, the improvement in parameters over the trivial solution above is not enough to translate into an appreciably better lower bound.*

6 Lower bounds for ABPs with Help Polynomials

In this section, we prove some lower bounds for ABPs computing some explicit polynomials using a set of given help polynomials H . Here, ‘explicit’ means that the coefficients of the polynomial can be written down in time polynomial in the number of coefficients of the input (the help polynomials H) and the output (the hard to compute polynomial).

Throughout this section, \mathbb{F} will be a field over which field operations and Gaussian elimination can be performed efficiently. Let the set of help polynomials be $H = \{h_1, h_2, \dots, h_m\}$; let $d(H) = \max_{h \in H} d(h)$.

We will first consider the case of homogeneous ABPs using the help polynomials H ; H is, in this case, assumed to be a set of homogeneous polynomials. We will then derive a lower bound for general ABPs and a general set of help polynomials using Theorem 7.

6.1 The homogeneous case

Let H be a set of homogeneous polynomials in this section. Our aim is to produce, for any degree $d \in \mathbb{N}$, an explicit homogeneous polynomial F_d of degree d that cannot be computed by homogeneous ABPs. To avoid some trivialities, we will assume that d is even.

We first observe that, to compute homogeneous polynomials of degree d , a homogeneous ABP cannot meaningfully use help polynomials of degree greater than d :

Lemma 12. *Let A be a homogeneous ABP using the help polynomials H to compute a polynomial f of degree d . Then, there is a homogeneous ABP A' , of size at most the size of A , such that A' computes f and furthermore, for every edge $e \in E(A')$, $d(e) \leq d$.*

Proof. Simply take A and throw away all edges $e \in E(A)$ such that $d(e) > d$; call the resulting homogeneous ABP A' . Since A is homogeneous, no path from source to sink in A can contain an edge e that was removed above. Hence, the polynomial computed remains the same. \square

Hence, to prove a lower bound for an explicit homogeneous polynomial of degree d , it suffices to prove a lower bound on the sizes of ABPs computing this polynomial using the help polynomials $H_{\leq d} = \{h \in H \mid d(h) \leq d\}$. As above, let $d(H_{\leq d}) = \max_{h \in H_{\leq d}} d(h)$.

We begin with a simple explicit lower bound. Call a homogeneous polynomial $F \in \mathbb{F}\langle X \rangle$ of degree d *d-full-rank* if $\text{rank}(M_{d/2}(F)) = n^{d/2}$. Full-rank polynomials are easily constructed; here is a simple example of one: $F(X) = \sum_{m \in \text{Mon}_{d/2}(X)} m \cdot m$. It follows easily from Nisan's result [9] that, without any help polynomials, homogeneous ABPs computing any d -full-rank polynomial are of size at least $n^{d/2}$.

Theorem 13. *Assume that $d(H_{\leq d}) \leq d(1 - \epsilon)$, for a fixed constant $\epsilon > 0$ and let $F \in \mathbb{F}\langle X \rangle$ be a d -full-rank polynomial. Then, any homogeneous ABP A computing F has size at least $\left(n^{\frac{\epsilon d}{4}} / \sqrt{2md}\right)$.*

Proof. Consider a homogeneous ABP A computing F using the help polynomials H . By the above lemma, we may assume that A uses only the polynomials $H_{\leq d}$. Let S denote the size of A . For any $h \in H_{\leq d}$, let $d_1(h)$ denote $\max\{1, d(h) - d/2\}$ and $d_2(h)$ denote $\min\{d/2, d(h) - 1\}$. By Theorem 8, we know that

$$M_{d/2}(F) = M' + \sum_{h \in H_{\leq d}} \sum_{i=d_1(h)}^{d_2(h)} M_i(h) \otimes_{i, d(h)-i}^{d/2} M'_{i,h}$$

where $\text{rank}(M') \leq S$ and $\text{rank}(M'_{i,h}) \leq S^2$, for each $h \in H_{\leq d}$ and $i \in \{d_1(h), \dots, d_2(h)\}$. For any h and any i such that $0 \leq i \leq d(h)$, $\text{rank}(M_i(h)) \leq \min\{n^i, n^{d(h)-i}\}$, which is at most $n^{d(h)/2} \leq n^{d(H_{\leq d})/2}$. By our assumption on $d(H_{\leq d})$, we see that $\text{rank}(M_i(h)) \leq n^{(1-\epsilon)d/2}$. By the definition of $\otimes_{i, d(h)-i}^{d/2}$, this implies that $\text{rank}(M_i(h) \otimes_{i, d(h)-i}^{d/2} M'_{i,h}) \leq \text{rank}(M_i(h)) \cdot$

$\text{rank}(M'_{i,h})$, which is at most $n^{(1-\epsilon)d/2}S^2$. Thus, we see that

$$\begin{aligned}\text{rank}(M_{d/2}(F)) &\leq S + \sum_{h \in H_{\leq d}} \sum_{i=d_1(h)}^{d_2(h)} n^{(1-\epsilon)d/2} S^2 \\ &\leq S + |H_{\leq d}| d n^{(1-\epsilon)d/2} S^2 \\ &\leq 2mdS^2 n^{(1-\epsilon)d/2}\end{aligned}$$

As F is d -full-rank, this implies that

$$\begin{aligned}2mdS^2 n^{(1-\epsilon)d/2} &\geq n^{d/2} \\ \therefore S &\geq \frac{n^{\frac{\epsilon d}{4}}}{\sqrt{2md}}\end{aligned}$$

□

The above theorem tells us that as long as the help polynomials are not too many in number ($m = n^{o(d)}$ will do), and of degree at most $(1 - \epsilon)d$, then any full rank polynomial remains hard to compute for ABPs with these help polynomials.

We now consider the case when $d(H_{\leq d})$ can be as large as d . In this case, we are unable to come up with an unconditional explicit lower bound. A strong solution to the RMP introduced in Section 5 would give us such a bound. However, with the suboptimal solution of Theorem 10, we are able to come up with explicit lower bounds in a special case. Let $\delta(H)$ denote $\min_{h \in H} d(h)$. By assuming some lower bounds on $\delta(H)$, we are able to compute an explicit hard function.

Theorem 14. *Assume $\delta(H) \geq (\frac{1}{2} + \epsilon)d$, for a fixed constant $\epsilon > 0$. Then, there exists an explicit homogeneous polynomial $F \in \mathbb{F}\langle X \rangle$ of degree d such that any homogeneous ABP A computing F using the help polynomials H has size at least $\lfloor n^{\frac{\epsilon d}{2}} / 2md \rfloor$.*

Proof. Let A be a homogeneous ABP A of size S computing a polynomial f of degree d . Let $d_1(h)$ denote $\max\{1, d(h) - d/2\}$ and $d_2(h)$ denote $\min\{d/2, d(h) - 1\}$. As explained in Remark 9, let $E_{a,b}^{p,q}$ denote the $n^a \times n^b$ -sized elementary matrix with 1 in the (p, q) th entry and 0s elsewhere. The matrices $\{E_{a,b}^{p,q} \mid p \in [n^a], q \in [n^b]\}$ span all $n^a \times n^b$ matrices.

By Theorem 8

$$M_{d/2}(f) = M' + \sum_{h \in H_{\leq d}} \sum_{i=d_1(h)}^{d_2(h)} M_i(h) \otimes_{i, d(h)-i}^{d/2} M'_{i,h}$$

where $\text{rank}(M') \leq S$. As explained in Remark 9, $M'_{i,h}$ is an $n^{d/2-i} \times n^{d/2-d(h)+i}$ dimension matrix and is in the span of $\{E_{d/2-i, d/2-d(h)+i}^{p,q}\}$, where $p \in [n^{d/2-i}]$, $q \in [n^{d/2-d(h)+i}]$.

Let \mathcal{A} denote the set of $n^{d/2} \times n^{d/2}$ matrices of the form $M_i(h) \otimes_{i, d(h)-i}^{d/2} E_{d/2-i, d/2-d(h)+i}^{p,q}$, where $h \in H_{\leq d}$, $i \in \{d_1(h), \dots, d_2(h)\}$, and $p \in [n^{d/2-i}]$, $q \in [n^{d/2-d(h)+i}]$. Then we obtain

$$M_{d/2}(f) = M' + \sum_{M \in \mathcal{A}} \alpha_M M, \quad (8)$$

where $\alpha_M \in \mathbb{F}$. Since M' is a matrix of rank at most S , this implies that M is at rank distance at most S from the subspace generated by the matrices in \mathcal{A} .

Let $k = |\mathcal{A}|$. For each $h \in H$ and $i \in \{d_1(h), \dots, d_2(h)\}$, we have added precisely $n^{d-d(h)}$ many matrices of the form $M_i(h) \otimes_{i, d(h)-i}^{d/2} E$, where E is an elementary matrix of dimension $n^{d/2-i} \times n^{d/2-d(h)+i}$. Since $d(h) \geq d(\frac{1}{2} + \epsilon)$ for each $h \in H_{\leq d} \subseteq H$, this implies that $k \leq mdn^{\frac{d}{2}(1-\epsilon)}$. Let N denote $n^{d/2}$; \mathcal{A} consists of $k \leq mdN^{1-\epsilon}$ $N \times N$ matrices. By Theorem 10, we can, in time $\text{poly}(N)$, come up with an $N \times N$ matrix M_0 that is at rank distance at least $\lfloor \frac{N}{k+1} \rfloor$ from the subspace generated by the matrices in \mathcal{A} . We label the rows and columns of M_0 by monomials from $\text{Mon}_{d/2}(X)$, in the same way as the matrices in \mathcal{A} are labeled. Using M_0 , we define the homogeneous degree d polynomial $F \in \mathbb{F}\langle X \rangle$ to be the unique polynomial such that $M_{d/2}(F) = M_0$; that is, given any monomial $m \in \text{Mon}_d(X)$ such that $m = m_1 \cdot m_2$ for $m_1, m_2 \in \text{Mon}_{d/2}(X)$, $F(m)$ is defined to be $M_0(m_1, m_2)$.

Let A be a homogeneous ABP of size S computing F using the help polynomials H . Then, by Equation 8 we have

$$M_{d/2}(F) = M' + \sum_{M \in \mathcal{A}} \alpha_M M$$

where $\alpha_M \in \mathbb{F}$, and $\text{rank}(M') \leq S$. Since $M_{d/2}(F)$ is M_0 , which is at rank distance at least $\lfloor N/(k+1) \rfloor$ from the subspace generated by \mathcal{A} , we see that $S \geq \text{rank}(M') \geq \lfloor N/(k+1) \rfloor$. This implies that,

$$S \geq \left\lfloor \frac{N}{mdN^{1-\epsilon} + 1} \right\rfloor \geq \left\lfloor \frac{N^\epsilon}{2md} \right\rfloor = \left\lfloor \frac{n^{\frac{\epsilon d}{2}}}{2md} \right\rfloor$$

□

Remark 15. *The rather unnatural condition on $\delta(H)$ above can be removed with better solutions to the RMP problem. Specifically, one can show along the above lines that if the RMP has an $(N, k, N/k^{1/2-\epsilon})$ -solution for $k = N^{2\delta}$, then for any H , there is an explicit polynomial that cannot be computed by any ABP A using H of size at most $n^{\Omega(\epsilon d)}/(md)^{O(1)}$. Here, ϵ and δ are arbitrary constants in $(0, 1)$.*

6.2 The inhomogeneous case

Let \tilde{H} denote the set of all homogeneous parts of degree at least 2 obtained from polynomials in H , i.e. $\tilde{H} = \{h_j^{(i)} \mid j \in [m], 2 \leq i \leq d(h_j)\}$. For $2 \leq i \leq d(H)$, let $\tilde{H}_i = \{h \in \tilde{H} \mid d(h) = i\}$. Note that $\tilde{H} = \bigcup_{2 \leq i \leq d(H)} \tilde{H}_i$.

As in the previous subsection, we construct explicit hard polynomials for even $d \in \mathbb{N}$. Let $\tilde{H}_{\leq d}$ denote $\bigcup_{2 \leq i \leq d} \tilde{H}_i$ if $d \leq d(H)$, and \tilde{H} otherwise.

Corollary 16. *Assume $d(\tilde{H}_{\leq d}) \leq d(1 - \epsilon)$, for a fixed constant $\epsilon > 0$. Then, there is an explicit homogeneous polynomial F of degree d such that any ABP that computes F using the help polynomials H has size at least $\frac{n^{\frac{\epsilon d}{4}}}{\sqrt{2md(d+1)}}$.*

Proof. Let F be a d -full-rank polynomial, as defined in Section 6.1. Consider any ABP A computing F using H . By Theorem 7, there exists a homogeneous ABP \tilde{A} computing F using \tilde{H} , where the size of \tilde{A} is at most $S(d+1)$. By Lemma 12, we may assume that \tilde{A} uses only the help polynomials in $\tilde{H}_{\leq d}$. Since $|\tilde{H}_{\leq d}| \leq md$, Theorem 13 tells us that $S(d+1) \geq n^{\frac{cd}{4}}/\sqrt{2md^2}$, which implies the result. \square

Corollary 17. *Let $\delta(\tilde{H}) = \min_{h \in \tilde{H}} d(h)$, and assume $\delta(\tilde{H}) \geq (\frac{1}{2} + \epsilon)d$ for a fixed constant $\epsilon > 0$. Then, there exists an explicit homogeneous polynomial $F \in \mathbb{F}\langle X \rangle$ of degree d such that any ABP A computing F using the help polynomials H has size at least $\frac{1}{d+1} \left\lfloor \frac{n^{\frac{cd}{2}}}{2md^2} \right\rfloor$.*

Proof. By Theorem 7, given any ABP A of size S computing a homogeneous polynomial of degree d , there is a homogeneous ABP \tilde{A} of size at most $S(d+1)$ that computes the same polynomial as A using the help polynomials \tilde{H} . By Lemma 12, we may assume that \tilde{A} only uses the help polynomials $\tilde{H}_{\leq d}$. Now, let F be the explicit polynomial from Theorem 14, with $\tilde{H}_{\leq d}$ taking on the role of H in the statement of the theorem; since $|\tilde{H}_{\leq d}| \leq md$, Theorem 14 tells us that $S(d+1) \geq \lfloor n^{\frac{cd}{2}}/2md^2 \rfloor$, which implies the result. \square

7 A better solution to the RMP

Following the approach of Alon et al [3], who provide a nontrivial algorithm for RPP in the Hamming metric (over \mathbb{F}_2), we improve on the parameters of Theorem 10 for the RMP over small prime fields. It is interesting to note that in our solution we get similar parameters as [3]. As mentioned earlier, the improvement in parameters over the simple solution of Theorem 10 is too little to give us a much better lower bound.

Throughout this section, \mathbb{F} will denote a constant-sized field. The main result is stated below.

Theorem 18. *For any fixed constant $c > 0$, the RMP has an $(N, \ell N, r)$ -solution over any constant-sized field \mathbb{F} and for any $\ell, r > 0$ such that $\ell \cdot r < c \log N$.*

In proving the above theorem, we will follow the algorithm of [3]. We need the following lemma, implicit in [3]:

Lemma 19. *Fix any field \mathbb{F} such that Gaussian elimination over \mathbb{F} can be performed in polynomial time. There is a $\text{poly}(M, m, |\mathbb{F}|)$ time algorithm for the following problem: Given subspaces V_1, V_2, \dots, V_m of \mathbb{F}^M such that $\sum_{i=1}^m |V_i| < |\mathbb{F}|^M$, find a point $u \in \mathbb{F}^M$ such that $u \notin \bigcup_i V_i$.*

Proof. The algorithm will fix the coordinates of u one by one. Assuming that the values u_1, u_2, \dots, u_i have been fixed for $0 \leq i \leq n$, let $U_i = \{w \in \mathbb{F}^M \mid w_j = u_j \text{ for } 1 \leq j \leq i\}$. The algorithm will fix the coordinates of u , ensuring that the following is true: For each i such that $1 \leq i \leq M$, $\sum_{j=1}^m |V_j \cap U_i| < |U_i| = |\mathbb{F}|^{M-i}$. Note that, since U_0 is just \mathbb{F}^M , the inequality is satisfied at $i = 0$ by the assumption on the size of the subspaces V_1, V_2, \dots, V_m ; also note that the inequality is satisfied at $i = M$ if and only if $u \notin \bigcup_i V_i$.

Assuming u_1, u_2, \dots, u_i have been fixed for $i < M$, we define, for every $\alpha \in \mathbb{F}$, the set $U_{i,\alpha} = \{w \in U_i \mid w_{i+1} = \alpha\}$. Clearly, the sets $\{U_{i,\alpha}\}_\alpha$ partition U_i . Hence, we see that $\sum_{j=1}^m |V_j \cap U_i| = \sum_{\alpha \in \mathbb{F}} \sum_{j=1}^m |V_j \cap U_{i,\alpha}|$ and thus, there is some $\alpha \in \mathbb{F}$ such that $\sum_{j=1}^m |V_j \cap U_{i,\alpha}| < \frac{|U_i|}{|\mathbb{F}|} = |\mathbb{F}^{M-i-1}|$.

Here is the algorithm:

- While u_1, u_2, \dots, u_i have been determined for $i < M$, do the following:
 - As mentioned above, the following invariant is maintained: $\sum_{j=1}^k |V_j \cap U_i| < |U_i| = |\mathbb{F}^{M-i}$.
 - Find $\alpha \in \mathbb{F}$ such that $\sum_{j=1}^k |V_j \cap U_{i,\alpha}| < \frac{|U_i|}{|\mathbb{F}|} = |\mathbb{F}^{M-i-1}|$. By the reasoning in the paragraph above, such an α exists and surely, it can be found in $\text{poly}(M, k, |\mathbb{F}|)$ time using Gaussian elimination.
 - Set u_{i+1} to α .

The correctness of the algorithm is clear from the reasoning above. □

We now briefly describe the improved algorithm for the RMP. Let P_1, P_2, \dots, P_k be the input matrices. We denote by L the subspace of $\mathbb{F}^{N \times N}$ spanned by these matrices. Also, let B_r denote the matrices of rank at most r . The idea of the algorithm is to “cover” the set $L + B_r$ by a union of subspaces V_1, V_2, \dots, V_m such that $\sum_i |V_i| < |\mathbb{F}|^{N^2}$. We then use the algorithm from Lemma 19 to find a matrix P that is not in $\bigcup_i V_i$; by the way we have picked the subspaces, it is clear that M will then be at rank distance at least r from the subspace L .

What follows is an important definition.

Definition 20. Fix positive integers (d_1, d_2) . Given \mathcal{T} , a collection of subspaces of \mathbb{F}^N , we say that \mathcal{T} is (d_1, d_2) -good if:

- $\dim(U) \leq N - d_1$ for each $U \in \mathcal{T}$.
- Each $A \subseteq \mathbb{F}^N$ of size d_2 is contained in some $U \in \mathcal{T}$.

The following claim illustrates the importance of (d_1, d_2) -good subspaces of \mathbb{F}^N .

Claim 21. There is an algorithm that, when given as input \mathcal{T} , a (d_1, d_2) -good collection of subspaces of \mathbb{F}^N , produces a collection \mathcal{S} of subspaces of $\mathbb{F}^{N \times N}$ of cardinality at most $|\mathcal{T}|$, with the following properties:

- $\dim(V) \leq N^2 - d_1 N$ for each $V \in \mathcal{S}$.
- $B_{d_2} \subseteq \bigcup_{V \in \mathcal{S}} V$

Moreover, the algorithm runs in time $\text{poly}(|\mathcal{T}|, N)$.

Proof. For each $U \in \mathcal{T}$, let $V(U)$ denote the subspace of $\mathbb{F}^{N \times N}$ generated by all vectors of the form uv^T , where $u \in U$ and $v \in \mathbb{F}^N$. The collection \mathcal{S} is the collection of all such vector spaces $V(U)$, for $U \in \mathcal{T}$. Clearly, the cardinality of \mathcal{S} is bounded by $|\mathcal{T}|$.

Note that a basis for $V(U)$ can be constructed by picking only uv^T where u and v range over bases for U and \mathbb{F}^N respectively. This shows that $\dim(V(U)) \leq N^2 - d_1N$ and that $V(U)$ can be constructed efficiently.

Finally, given any matrix Q of rank at most d_2 , it can be written as a sum of matrices $Q_1 + Q_2 + \dots + Q_{d_2}$, where each Q_i is a matrix of rank at most 1 and hence can be written as $u_i v_i^T$, where $u_i, v_i \in \mathbb{F}^N$. Let $A = \{u_1, u_2, \dots, u_{d_2}\}$. Since \mathcal{T} is (d_1, d_2) -good, there is some $U \in \mathcal{T}$ such that $A \subseteq U$. This implies that $u_i v_i^T \in V(U)$ for each $i \in [d_2]$. As $V(U)$ is a subspace, it must contain their sum Q . This concludes the proof. \square

It is easily seen that a random collection of subspaces of \mathbb{F}^N of appropriate dimension is (d_1, d_2) -good for the values of d_1 and d_2 that are of interest to us. We now assert the existence of an explicit collection of subspaces with this property.

Claim 22. *Fix any constant $c \geq 1$. For any $\ell, r \in \mathbb{N}$ such that $\ell \cdot r < c \log N$, there is an algorithm that runs in time $N^{O(c)}$ and produces an (ℓ, r) -good collection of subspaces of \mathbb{F}^N .*

We prove the above claim in the next section. Assuming the claim, we can prove Theorem 18.

Proof of Theorem 18. We will describe an algorithm for the problem. Without loss of generality, assume that $c \geq 1$. Let L be the input subspace of dimension at most ℓN . We would like to find a matrix P that is at rank distance at least r from L .

We first use the algorithm referred to in Claim 22 to construct an $(\ell + 1, r)$ -good collection of subspaces \mathcal{T} of \mathbb{F}^N in time $N^{O(c)}$. Clearly, $|\mathcal{T}| = N^{O(c)}$. Then, we use the algorithm of Claim 21 to construct a collection of subspaces \mathcal{S} of $\mathbb{F}^{N \times N}$ of size $N^{O(c)}$ with the following properties:

- $\dim(V) \leq N^2 - (\ell + 1)N$ for each $V \in \mathcal{S}$.
- $B_r \subseteq \bigcup_{V \in \mathcal{S}} V$

Consider the collection of subspaces $\mathcal{S}' = \{L + V \mid V \in \mathcal{S}\}$. Clearly, $L + B_r \subseteq \bigcup_{V \in \mathcal{S}'} V$. Moreover, the dimension of each subspace in \mathcal{S}' is at most $\ell N + N^2 - (\ell + 1)N \leq N^2 - N$. Hence, each subspace in \mathcal{S}' is of cardinality at most $|\mathbb{F}|^{N^2 - N}$. Since $|\mathcal{S}'| = N^{O(c)}$, for large enough N , we have $\sum_{V \in \mathcal{S}'} |V| < |\mathbb{F}|^{N^2}$. Hence, using the algorithm described in Lemma 19, we can, in time $N^{O(c)}$, find a matrix $P \notin \bigcup_{V \in \mathcal{S}'} V$. By construction, this matrix P is at rank distance greater than r from the subspace L . The entire algorithm runs in time $N^{O(c)}$. \square

7.1 Proof of Claim 22

We give two different constructions: one for the case that $\ell \geq r$ and the other for the case that $\ell \leq r$.

The following notation will be useful. For each $i \in [N]$, let $e_i \in \mathbb{F}^N$ denote the vector that has a 1 in coordinate i and is 0 elsewhere. For any vector $x \in \mathbb{F}^N$ and $S \subseteq [N]$, we denote by $x|_S$ the vector in $\mathbb{F}^{|S|}$ that is the projection of x to the coordinates indexed by S .

7.1.1 Case 1: $\ell \geq r$

For each $A \subseteq \mathbb{F}^{2\ell}$ of cardinality r , let V_A be the subspace generated by $\{x \in \mathbb{F}^N \mid x|_{[2\ell]} \in A\}$. It is easily seen that $\dim(V_A) \leq N - 2\ell + r \leq N - \ell$. Moreover, given any $A_1 \subseteq \mathbb{F}^{2\ell}$ of size r , $A_1 \subseteq V_A$ where A is any subset of $\mathbb{F}^{2\ell}$ of size r containing $\{x|_{[2\ell]} \mid x \in A_1\}$. Hence, the collection $\mathcal{T} = \{V_A \mid A \subseteq \mathbb{F}^{2\ell}, |A| = r\}$ is an (ℓ, r) -good collection of subspaces.

The cardinality of \mathcal{T} is $\binom{|\mathbb{F}^{2\ell}|}{r} \leq |\mathbb{F}|^{2\ell r} = N^{O(c)}$. Surely, \mathcal{T} can be constructed in time $N^{O(c)}$.

7.1.2 Case 2: $\ell \leq r$

Given a set $A \subseteq \mathbb{F}^m$ for some $m \in \mathbb{N}$, we denote by $\text{rank}(A)$ the size of any maximal set of linearly independent vectors from A ; we denote by $\text{corank}(A)$ the value $(|A| - \text{rank}(A))$.

Fix a set $A \subseteq \mathbb{F}^m$ for some $m \in \mathbb{N}$. Given $d, d' \in \mathbb{N}$, we say that A is d -wise corank d' if each $B \subseteq A$ such that $|B| = d$ satisfies $\text{corank}(B) \leq d'$; A is said to be d -wise linearly independent if it is d -wise corank 0. Sets that are d -wise linearly independent have been studied before: see [1, Proposition 6.5], where matrices whose columns form a d -wise linearly independent set of vectors are used to construct d -wise independent sample spaces. The following claim follows from this result and from the lower bound on the size of any d -wise independent sample space proved in [1, Proposition 6.4].

Claim 23 (implicit in [1]). *Consider a set $A \subseteq \mathbb{F}^m$ of cardinality t . If A is d -wise linearly independent with $d \leq 2\sqrt{t}$, then $m \geq \frac{d \log_{|\mathbb{F}|} t}{5}$, for large enough d, t .*

Using the above claim, we prove the following lower bound on the size of sets that are d -wise corank d' for suitable d, d' .

Claim 24. *Consider a set $A \subseteq \mathbb{F}^r$ of cardinality t . There is an absolute constant c_0 such that the following holds. Let A be d -wise corank d' for positive integers d, d' with $c_0 d' \leq d \leq 2\sqrt{t}$. Then, $r \geq \frac{d \log_{|\mathbb{F}|} t}{12d'}$ if t, d, d' are large enough.*

Proof. Denote by d'' the value $\lfloor d/2d' \rfloor$. We construct a sequence of sets A_0, A_1, \dots as follows: A_0 is the set A ; for any $i \geq 0$, if A_i has been constructed and is d'' -wise linearly independent, we stop; otherwise, there is a $B \subseteq A_i$ of cardinality d'' that is not linearly independent – in

this case, we set $A_{i+1} = A_i \setminus B$; we stop at $i = d'$. It is easy to see that the cardinality t_i of A_i is $t - id''$. It can also be checked that if A_i is d_i -wise corank d'_i , then A_{i+1} , if constructed, is $(d_i - d'')$ -wise corank $d'_i - 1$; it therefore follows that the set S_i , if constructed, is $(d - id'')$ -wise corank $d' - i$, for any $i \geq 0$ – in particular, $S_{d'}$ is $d/2$ -wise linearly independent.

We base our analysis on when the above process stops. Let i_0 be the largest i so that A_i is constructed. Its size t_{i_0} is at least $t - d'd'' \geq t - d/2 \geq t/2$ for large enough t . If $i_0 = d'$, then A_{i_0} is a set of size at least $t/2$ that is $d/2$ -wise linearly independent – by Claim 23, we get $r \geq \frac{d \log_{|\mathbb{F}|} t}{12}$ for large enough d, t . Otherwise, $i_0 < d'$ and we must have A_{i_0} is d'' -wise linearly independent – in this case, by Claim 23, we get $r \geq \frac{d'' \log_{|\mathbb{F}|} t}{5} \geq \frac{d \log_{|\mathbb{F}|} t}{12d'}$ if c_0 is large enough. Thus, in either case, our claim holds. \square

Now, we apply the above lemma with $t = |\mathbb{F}|^{\lceil \frac{20}{c_0} \sqrt{c \log N} \rceil}$ and $d = c_0 \lceil \sqrt{c \log N} \rceil$. We obtain the following corollary:

Corollary 25. *Let t, d be as defined above. For large enough N , given any $A \subseteq \mathbb{F}^r$ of size t , there is a subset B of A of cardinality d such that $\text{corank}(B) \geq \ell$.*

Proof. Assume that A is d -wise corank d' for some d' . We will show that $d' \geq \ell$. For large enough N , by Claim 24, we have $d' \geq \min\{\frac{d}{c_0}, \frac{d \log_{|\mathbb{F}|} t}{12r}\}$. It remains to be shown that this quantity is at least ℓ .

Note that, since $\ell \leq r$, $\ell^2 \leq \ell r \leq c \log N$. Hence, $\ell \leq \sqrt{c \log N}$. Thus, by the choice of d , we see that $d/c_0 \geq \ell$. Moreover,

$$\frac{d \log_{|\mathbb{F}|} t}{12r} \geq \frac{20c \log N}{12r} > \ell$$

Hence, we see that $d' \geq \ell$. \square

We now define the (ℓ, r) -good collection of subspaces. For each $S \subseteq [t]$ of cardinality d , and each $A \subseteq \mathbb{F}^d$ of size $d - \ell$, let $V_{S,A}$ be the subspace generated by $\{x \in \mathbb{F}^N \mid x|_S = u \text{ for some } u \in A\}$. It can be seen that $\dim(V_{S,A}) \leq N - d + d - \ell = N - \ell$ for each S, A .

Given any $A_1 \subseteq \mathbb{F}^N$ of cardinality r , let $P \in \mathbb{F}^{r \times N}$ be the matrix the rows of which are the elements of A_1 . Let A_2 denote the set of the first t columns of P . By Corollary 25, there is a $B \subseteq A_2$ of size d such that $\text{corank}(B) \geq \ell$. Let $S \subseteq [t]$ index the columns of B in P . It can be seen that $A_1 \subseteq V_{S,A'}$ for any A' of size $d - \ell$ containing a set that spans $\{v|_S \mid v \in A_1\}$ (such an A' exists since $\text{corank}(B) \geq \ell$).

Thus, we can take for our collection \mathcal{T} of (ℓ, r) -good subspaces the collection of all $V_{S,A}$, where $S \subseteq [t]$ with $|S| = d$, and $A \subseteq \mathbb{F}^d$ of size $d - \ell$. The size of \mathcal{T} is bounded by $\binom{t}{d} \binom{|\mathbb{F}|^d}{d-\ell} \leq t^d |\mathbb{F}|^{d^2} = N^{O(c)}$, by our choice of d and t . Clearly, \mathcal{T} can be constructed in time $N^{O(c)}$.

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