

Low Rate Is Insufficient for Local Testability

Eli Ben-Sasson* Michael Viderman
 Computer Science Department
 Technion — Israel Institute of Technology
 Haifa, 32000, Israel.
 {eli, viderman}@cs.technion.ac.il

May 4, 2010

Abstract

Locally testable codes are error-correcting codes for which membership of a given word in the code can be tested probabilistically by examining it in very few locations. A linear code $C \subseteq \mathbb{F}_2^m$ is called sparse if $\dim(C) = O(\log(n))$. We say that a code $C \subseteq \mathbb{F}_2^m$ is ϵ -biased if all nonzero codewords of C have relative weight in the range $(\frac{1}{2} - \epsilon, \frac{1}{2} + \epsilon)$, where ϵ may be a function of n .

Kaufman and Sudan [11] proved that all sparse linear codes with relative distance $\frac{1}{2} - n^{-\Omega(1)}$ are locally testable. Moreover, they showed that all sparse $n^{-\Omega(1)}$ -biased linear codes are locally decodable. In particular sparse random codes are locally testable and are locally decodable with probability $1 - o(1)$.

Kopparty and Saraf [12] conjectured that all sparse linear codes (even with a bad distance) are locally testable.

In this paper we refute this conjecture by showing that for every $d(n)$ ranging from $\omega(1)$ to $\Omega(n)$ there exists a family of codes $\{C^{(n)} \subset \mathbb{F}_2^n\}_{n \in \mathbb{Z}}$ with linear distance and $\dim(C^{(n)}) = \Theta(d(n))$ which are not locally testable (decodable). Furthermore, we show that there exists a family of codes $\{C^{(n)} \subset \mathbb{F}_2^n\}_{n \in \mathbb{Z}}$ with bias $n^{-o(1)}$ and $\dim(C^{(n)}) = \log(n)$ which are not locally testable (decodable). This also shows that the results of Kaufman and Sudan [11] were surprisingly tight.

1 Introduction

Locally testable codes (LTCs) are error correcting codes for which distinguishing, when given oracle access to a purported word w , between the case that w is a codeword and the case that it is very far from all codewords, can be accomplished by a randomized algorithm, called a *tester*, which reads a constant amount of information from w .

On the other hand, *locally decodable codes (LDCs)* allow to recover each message entry with high probability by reading only a few entries of the codeword even if a constant fraction of it is adversely corrupted. Both families of error correcting codes are explicitly studied, see e.g. the survey [15].

Given a linear code $C \subseteq \mathbb{F}_2^n$, the dimension of C , denoted by $\dim(C)$, is its dimension as a vector space and its distance, denoted by $\Delta(C)$, is the minimal Hamming distance between two closest different

*The research leading to these results has received funding from the European Community's Seventh Framework Programme (FP7/2007-2013) under grant agreement number 240258. Research of both authors supported by grant number 2006104 by the US-Israel Binational Science Foundation and by grant number 679/06 by the Israeli Science Foundation.

codewords. A linear code $C \subseteq \mathbb{F}_2^n$ is called *sparse* if its dimension is $O(\log(n))$. We say that a code C is ϵ -biased if all nonzero codewords of C have relative weight in the range $(\frac{1}{2} - \epsilon, \frac{1}{2} + \epsilon)$, where ϵ may be a function of n . Notice that the ϵ -bias property implies that the relative distance is at least $\frac{1}{2} - \epsilon$.

Kaufman and Sudan [11] showed that local testability and decodability exists in random sparse linear codes. They showed that for any constant $\gamma > 0$ all sparse linear codes with relative distance $\frac{1}{2} - n^{-\gamma}$ are (strongly) locally testable. They also showed that for any constant $\gamma > 0$ all sparse $n^{-\gamma}$ -biased linear codes are locally decodable¹. In particular, sparse random linear codes have low-bias and hence are locally testable and decodable. This result was later generalized by Kopparty and Saraf [13] to the problem known as “local list-decoding and testing in the high error regime” (see [13] for the definition and discussion of the problem), i.e., they proved that all sparse $n^{-\Omega(1)}$ -biased linear codes are locally testable and locally list-decodable even in the high error regime.

Kopparty and Saraf [12] conjectured that all sparse linear codes (even with a bad distance) are locally testable. In particular, this conjecture says that the result of Kaufman and Sudan could be extended to all sparse linear codes and the $n^{-\Omega(1)}$ -bias requirement is not necessary.

We refute this conjecture in Theorem 3.1 by showing that for any $d(n)$ ranging from $\omega(1)$ to $\Omega(1)$ there exists a family of linear codes $\{C^{(n)} \subset \mathbb{F}_2^n\}_{n \in \mathbb{Z}}$ with linear distance and $\dim(C^{(n)}) = \Theta(d(n))$ which are not locally testable (decodable). Furthermore, in Theorem 3.4 we show that there exists a family of sparse $n^{-o(1)}$ -biased linear codes which are not locally testable (decodable), i.e., for any computable function $h(n) = o(1)$ we can construct a family of sparse $n^{-h(n)}$ -biased linear codes which are not locally testable (decodable). In particular, this theorem shows that the $n^{-\Omega(1)}$ -bias requirement in the work of Kaufman and Sudan [11] is necessary, and has a surprising tightness.

This paper demonstrates an interesting phenomena regarding sparse codes, i.e., while all sparse $n^{-\Omega(1)}$ -biased codes have been shown by [13] to be LTCs and LDCs in the high-error regime, we show that certain sparse $n^{-o(1)}$ -biased codes are not LTCs and LDCs even in the easier-to-attain low-error regime.

Organization of the paper. In the following section we provide the standard definitions regarding locally testable and locally decodable codes. In Section 3 we state our main results (Theorems 3.1, 3.4). Section 4 contains some useful propositions and in Section 5 we prove Theorems 3.1 and 3.4.

2 Preliminaries

Let \mathbb{F} be a finite field and $[n]$ be the set $\{1, \dots, n\}$. In this work, we consider only linear codes. We start with a few definitions.

Let $C \subseteq \mathbb{F}^n$ be a linear code over \mathbb{F} . For $w \in \mathbb{F}^n$, let $\text{supp}(w) = \{i \in [n] \mid w_i \neq 0\}$ and $|w| = |\text{supp}(w)|$. We define the relative distance between two words $x, y \in \mathbb{F}^n$ to be $\delta(x, y) = \frac{\Delta(x, y)}{n}$ and let $\Delta(x, y) = \delta(x, y) \cdot n$. The distance of a code is denoted by $\Delta(C)$ and defined to be the minimal value of $\Delta(x, y)$ for two distinct codewords $x, y \in C$. Similarly, the relative distance of the code is denoted $\delta(C) = \frac{\Delta(C)}{n}$. For $x \in \mathbb{F}^n$ and $C \subseteq \mathbb{F}^n$, let $\delta(x, C) = \delta_C(x) = \min_{y \in C} \{\delta(x, y)\}$ denote the relative distance of x from the code C . We note that $\Delta(C) = \min_{c \in C \setminus \{0\}} \{\text{wt}(c)\}$. For two linear codes $C_1, C_2 \subseteq \mathbb{F}^n$ we let $\delta(C_1, C_2) = \min_{c_1 \in C_1 \setminus \{0\}} \{\delta(c_1, C_2)\}$.

¹In fact, Kaufman and Sudan [11] proved a stronger result. They showed that sparse “low-bias” linear codes are locally self-correctable and thus are locally decodable.

If $\delta(x, C) \geq \epsilon$, we say that x is ϵ -far from C and otherwise x is ϵ -close to C . Let $\dim(C)$ be the dimension of C . For $u = (u_1, u_2, \dots, u_n), v = (v_1, v_2, \dots, v_n) \in \mathbb{F}^n$ let $\langle u, v \rangle$ denote the bilinear function from $\mathbb{F}^n \times \mathbb{F}^n$ to \mathbb{F} defined by $\langle u, v \rangle = \sum_{i=1}^n u_i v_i$. The dual code C^\perp is defined as $C^\perp = \{u \in \mathbb{F}^n \mid \forall c \in C : \langle u, c \rangle = 0\}$. In a similar way we define $C_{\leq t}^\perp = \{u \in C^\perp \mid |u| \leq t\}$ and $C_t^\perp = \{u \in C^\perp \mid |u| = t\}$. For $w \in \mathbb{F}^n$ and $S = \{j_1, j_2, \dots, j_m\} \subseteq [n]$ we let $w|_S = (w_{j_1}, w_{j_2}, \dots, w_{j_m})$, where $j_1 < j_2 < \dots < j_m$, be the restriction of w to the subset S . Similarly, we let $C|_S = \{c|_S \mid c \in C\}$ denote the projection of the code C onto S . We say that a code C has a q -characterization if $\text{span}(C_{\leq q}^\perp) = C^\perp$.

For $w \in \mathbb{F}^n$ and $i \in \mathbb{N}$ let $w^{(i)} \in \mathbb{F}^{ni}$ be the concatenation of w to itself i times.

2.1 Background on LTCs and LDCs

We formally define locally testable codes, using the definition provided in [5] (see that paper for a justification of the definition).

Note that given a code $C \subseteq \mathbb{F}^n$, the subset $I \subseteq [n]$ uniquely defines $C|_I$. The linearity of C implies that $C|_I$ is a linear subspace of \mathbb{F}^I .

Definition 2.1 (LTCs and Testers). Let $C \subseteq \mathbb{F}^n$ be a linear code. A (q, ϵ, δ) -tester T for C is a distribution \mathbf{D} over subsets $I \subseteq [n]$ such that $|I| \leq q$ and the following holds.

- For all $w \in \mathbb{F}^n$ such that $\delta(w, C) \geq \delta$ we have $\Pr_{I \sim \mathbf{D}} [w|_I \notin C|_I] \geq \epsilon$.

A code $C \subseteq \mathbb{F}^n$ is a (q, ϵ, δ) -LTC if it has a (q, ϵ, δ) -tester.

The tester outputs **accept** on the given word w whenever it selects a subset $I \subseteq [n]$ such that $w|_I \in C|_I$ and otherwise output **reject**. Notice that a tester for a linear code is non-adaptive [2].

We say that a family of codes $\{C^{(n)} \mid n \in \mathbb{Z}\}$ is locally testable if there exist constants $q, \epsilon, \delta > 0$ such that for infinitely many n it holds that $C^{(n)} \subseteq \mathbb{F}^n$ is a (q, ϵ, δ) -LTC.

Remark 2.2. Usually we assume that $\delta \leq 1/3$ (see the discussion in [1] regarding this issue). Hence, we say that the family of codes is **not** locally testable if for every constants $q, \epsilon > 0$, large enough n and distribution \mathbf{D}_n over subsets $I \subseteq [n]$, such that $|I| \leq q$ there is a word w with $\delta(w, C^{(n)}) \geq 1/3$ and $\Pr_{I \sim \mathbf{D}_n} [w|_I \notin C|_I] < \epsilon$.

Next we define locally decodable codes.

Definition 2.3 (LDCs and Decoders). Let $C \subseteq \mathbb{F}_2^n$ and $E_C : \mathbb{F}_2^k \rightarrow \mathbb{F}_2^n$ be its encoding function, i.e., $C = \{E_C(x) \mid x \in \mathbb{F}_2^k\}$. Then C is a (q, ϵ, δ) -LDC if there exists a randomized decoder (\mathbb{D}) such that:

- In every invocation, \mathbb{D} makes at most q queries.
- For all $x \in \mathbb{F}_2^k$, $i \in [k]$ and $\hat{c} \in \mathbb{F}_2^n$ such that $\Delta(E_C(x), \hat{c}) \leq \delta n$ we have $\Pr \left[\mathbb{D}^{\hat{c}}[i] = x_i \right] \geq \frac{1}{2} + \epsilon$, i.e., with probability at least $\frac{1}{2} + \epsilon$ entry x_i will be recovered correctly.

Note that the definition implies that $\delta < \delta(C)/2$ and $\epsilon \leq \frac{1}{2}$. We can assume without loss of generality that the decoder for a linear code is non-adaptive [6]. We say that a family of codes $\{C^{(n)} \mid n \in \mathbb{Z}\}$ is

locally decodable if there exist constants $q, \epsilon, \delta > 0$ such that for infinitely many n it holds that $C^{(n)} \subseteq \mathbb{F}^n$ is a (q, ϵ, δ) -LDC.

It is well-known that q -query locally self-correctable codes are q -query locally decodable, so if we show that a code C is not q -query LDC then it is not q -query locally self-correctable.

Remark 2.4. Notice that a message space for LDC can be a linear subspace $M \subset \mathbb{F}^k$, i.e., the messages are of the length k but not every word in \mathbb{F}^k is a (legal) message. In this case the dimension of the code is $\dim(M)$. Note that because of the linearity of C the message space M must be a linear subspace.

3 Main Results

In this section we present our main results, i.e., Theorem 3.1 and Theorem 3.4. Theorem 3.1 refutes the conjecture of Kopparty and Saraf [12], which says that all sparse linear codes are locally testable. It shows that there exists a family of sparse linear codes with relative distance ≥ 0.49 which are not locally testable (decodable).

One could conjecture that all sparse linear codes with small characterization are locally testable or decodable. So, Theorem 3.1 also shows that there are sparse linear codes which have a small characterization and linear distance, but are not locally testable (decodable).

Finally we show that there are sparse linear codes with dual distance > 2 which are not locally testable (decodable). We believe that this case is interesting since it shows that non-redundant² sparse code can be not locally testable (decodable). Equivalently, picking a large fraction of columns from the Hadamard generating matrix ($G \in \mathbb{F}_2^{[n] \times [2^n]}$) may result in a generating matrix of a code which is *not* locally testable (decodable). This contrasts with the result of [11] which says that taking a large *random* fraction of columns from the Hadamard generating matrix gives, with high probability, a generating matrix of a locally testable (decodable) code.

Theorem 3.1 (Low Rate does not imply LTC or LDC). *For every $q, \epsilon > 0$, function $w(1) \leq d(n) \leq O(n)$ and infinitely many $n > 0$*

1. *There exists $C_\epsilon \subset \mathbb{F}_2^n$ such that $\delta(C) \geq \frac{1}{2} - \epsilon$, $\dim(C_\epsilon) = \Theta(d(n))$ and C_ϵ is not locally testable with $o(d(n))$ queries and is not locally decodable with q queries.*
2. *There exists $C \subset \mathbb{F}_2^n$ such that $\dim(C) = \Theta(d(n))$, $\Delta(C) \geq \Omega(n)$, $\text{span}(C_{\leq 3}^\perp) = C^\perp$ and C is not locally testable with $o(d(n))$ queries and is not locally decodable with q queries.*
3. *There exists $C \subset \mathbb{F}_2^n$ such that $\dim(C) = 1.1 \log(n)$, $\Delta(C) \geq \Omega(n)$, $\Delta(C^\perp) > 2$ and C is not locally testable with $o(\log(n))$ queries and is not locally decodable with q queries.*

Remark 3.2. Folklore claim 7.2 states that every linear code C is testable by $\dim(C) + 1$ queries. Intuitively, C is non-trivially testable if it can be testable with $o(\dim(C))$ queries. So, Theorem 3.1 shows the families of linear codes that can not be non-trivially testable.

Remark 3.3. The construction in the bullet three of Theorem 3.1 can not achieve dimension lower than $\log(n)$, since every linear code $C \subseteq \mathbb{F}_2^n$ such that $\dim(C) < \log(n)$ has $\Delta(C^\perp) \leq 2$.

²The term “non-redundant” here means the dual distance of the code is at least 3. Note that if the dual distance of a code is 1 then some of its bits are identically 0, and if the dual distance is 2 then some pairs of bits are equal to each other and hence one of each pair is redundant.

The next theorem shows a surprising tightness of the results in [11]. Recall that Kaufman and Sudan showed that sparse linear codes with bias $n^{-\Omega(1)}$ are locally testable (decodable). We show that for every $h(n) = o(1)$ there exists a family of sparse linear codes with bias $n^{-h(n)}$ which are not locally testable (decodable).

Theorem 3.4 (Tightness of [11]). *For every constant $q > 0$, computable function $h(n) = o(1)$ and infinitely many $n > 0$ there exists $C \subset \mathbb{F}_2^n$ such that $\dim(C) = \log(n)$, C is $n^{-h(n)}$ -biased and C is not locally testable (decodable) with q queries.*

4 Repetition does not affect the LTCs and LDCs

In this section we show that the “repetition” does not affect the testability and the decodability of the code. To do this we first define modulo- m testers and decoders and then prove our main propositions, i.e., Proposition 4.2 and Proposition 4.3.

4.1 Testers and Decoders modulo m

Definition 4.1 (Repetition Code). Let $R \subseteq \mathbb{F}_2^m$ be a linear code and $i > 0$. We say that $C \subseteq \mathbb{F}_2^{(im)}$ is the i -repetition code of R if it holds that $c \in C$ if and only if $c = r^{(i)}$, where $r \in R$, i.e., every codeword of C is a codeword of R repeated i times.

Notice that the linearity of R implies the linearity of its repetition code.

Testers and Decoders modulo- m . Let $R \subset \mathbb{F}_2^m$ be a linear code and $C \subset \mathbb{F}_2^{(mi)}$ be its (i) -repetition code. For $I = \{i_1, i_2, \dots, i_q\} \subseteq [n]$ let $I \bmod m = \{i_1 \bmod m, i_2 \bmod m, \dots, i_q \bmod m\} \subseteq [m]$.

Given a tester \mathbf{D}_C for C we define a (modulo- m)-tester \mathbf{D}_R for R as follows.

- Pick a test $I \subseteq [mi]$ according to the tester \mathbf{D}_C .
- Choose test $I \bmod m$.

The modulo- m decoder is defined in the similar way, i.e., invokes the decoder of C and takes modulo m over all queried indices.

Note that if $w \in \mathbb{F}_2^m$ then the tester (decoder) for C will get the same “queried entries” on $w^{(i)}$ as modulo- m tester (decoder) for R on w . In this way, if tester for C rejects then the tester for R rejects and if the decoder for C recovers correctly the message bit then the decoder for R recovers correctly the same message bit.

4.2 Main Propositions

We prove the simple (but important) propositions, which say that the repetition does not affect the testability and the decodability of the code. These propositions (Proposition 4.2 and Proposition 4.3) are shown for binary codes but can be easily extended for any field.

Proposition 4.2. *Let $R \subset \mathbb{F}_2^m$ be a linear code and $i > 0$ be an integer. Let $C \subset \mathbb{F}_2^{(mi)}$ be an (i) -repetition code of R . Then,*

- if R is a (q, ϵ, δ) -LTC then C is a $(q, \min\{\epsilon/2, \delta/2\}, 2\delta)$ -LTC

- if C is a (q, ϵ, δ) -LTC then R is a (q, ϵ, δ) -LTC.

Proof. Note that $C|_{[m]} = R$.

Proof of the first bullet. Assume that R is a (q, ϵ, δ) -LTC and let \mathbf{D}_R be a (q, ϵ, δ) -tester for R . We define a tester \mathbf{D}_C for C as following.

- flip a coin
- **If** “heads,”
 - pick $j \in [m]$ and $l_1 \in [i-1]$ independently at random,
 - pick $I = \{j, j + m \cdot l_1\}$ (note that $I \subseteq [mi]$);
- **Else** pick $I \sim \mathbf{D}_R$ (note that $I \subseteq [m]$).

We argue that \mathbf{D}_C is a $(q, \min\{\epsilon/2, \delta/2\}, 2\delta)$ -tester for C . Let $w \in \mathbb{F}^{(mi)}$ be a word such that $\delta(w, C) \geq 2\delta$. If $\delta(w|_{[m]}, C|_{[m]}) = \delta(w|_{[m]}, R) \geq \delta$ we are done, since

$$\Pr_{I \sim \mathbf{D}_C} [w|_I \notin C|_I] \geq \frac{1}{2} \cdot \Pr_{I \sim \mathbf{D}_R} [(w|_{[m]})|_I \notin R|_I] \geq \frac{\epsilon}{2}.$$

Otherwise we have $\delta(w|_{[m]}, C|_{[m]}) = \delta(w|_{[m]}, R) < \delta$.
But $\delta(w, C) \geq 2\delta$ implies that

$$\mathbf{E}_{j \in [i-1]} [\delta(w|_{\{jm+1, \dots, (j+1)m\}}, w|_{[m]})] \geq 2\delta - \delta = \delta.$$

Hence $\Pr_{I \sim \mathbf{D}_C} [w|_I \notin C|_I] \geq \frac{1}{2} \Pr_{j \in [m], l_1 \in [i-1]} [w|_{\{j, j+m \cdot l_1\}} \notin C|_{\{j, j+m \cdot l_1\}}] \geq \delta/2$. □

Proof of the second bullet. Assume that C is a (q, ϵ, δ) -LTC and let \mathbf{D}_C be its (q, ϵ, δ) -tester. Let \mathbf{D}_R be a modulo- m tester of C . Note that \mathbf{D}_R is a distribution over subsets $I \subseteq [m]$ such that $|I| \leq q$.

We argue that \mathbf{D}_R is a (q, ϵ, δ) -tester for R . Let $w \in F_2^m$ be a word such that $\delta(w, R) \geq \delta$. Assume by way of contradiction that $\Pr_{I \sim \mathbf{D}_R} [w|_I \notin R|_I] < \epsilon$. Notice that $\delta(w^{(im)}, C) = \delta(w, R) \geq \delta$. We have

$\Pr_{I \sim \mathbf{D}_C} [w^{(im)}|_I \in C|_I] < \epsilon$ since if for $I \subset [im]$ it holds that $w^{(im)}|_I \notin C|_I$ then $w|_{I \bmod m} \notin R$. We conclude that \mathbf{D}_C is not a (q, ϵ, δ) -distribution for C . Contradiction. □

□

Proposition 4.3. Let $R \subset \mathbb{F}_2^m$ be a code such that the first $\dim(R)$ bits of R are message bits. Let $i > 0$ and C be an (i) -repetition of R . Then,

- If R is not (q, ϵ, δ) -LDC then C is not (q, ϵ, δ) -LDC.
- If R is a (q, ϵ, δ) -LDC then C is a $(q, \epsilon/2, \frac{\epsilon\delta}{2})$ -LDC.

Proof. Let $k = \dim(R) = \dim(C)$. Note that the first k bits of C are message bits.

Proof of the first bullet. If R is not (q, ϵ, δ) -LDC then for every q -query decoder there exist a word $w \in \mathbb{F}_2^m$ such that $\delta(w, R) \leq \delta$, but there exists $i \in [k]$ such that the probability that the decoder recovers correctly the message bit i is less than $\frac{1}{2} + \frac{\epsilon}{2}$.

Assume by way of contradiction that C is a (q, ϵ, δ) -LDC and has the decoder \mathbb{D}_C . Let \mathbb{D}_R be a modulo- m decoder of \mathbb{D}_C . But then there exist $w \in \mathbb{F}^m, \delta(w, R) \leq \delta$ such that the decoder \mathbb{D}_R recovers some message bit i with probability less than $\frac{1}{2} + \frac{\epsilon}{2}$. But \mathbb{D}_C will get always the same information on $w^{(n/m)}$ as \mathbb{D}_R on w , and moreover, $\delta(w^{(n/m)}, C) = \delta(w, R) \leq \delta$. Thus \mathbb{D}_C is not a (q, ϵ, δ) -decoder for C . Contradiction. \square

Proof of the second bullet. Assume R is (q, ϵ, δ) -LDC and let D_R be its (q, ϵ, δ) -decoder. For $j \in [i]$ and $I = \{i_1, i_2, \dots, i_q\} \subseteq [m]$ let $jm + I = \{jm + i_1, jm + i_2, \dots, jm + i_q\}$. Note that $jm + I \subseteq \{m \cdot j + 1, \dots, m(j + 1)\}$. The decoder D_C for C recovers message bit l from the given word $w \in \mathbb{F}^{im}$ as following.

- pick $j \in [i - 1]$ and select $I = jm + [m]$,
- return $D_R^{(w|_I)}[l]$, i.e., recover the l th message bit as the decoder D_R on $w|_I$.

We argue that if $\delta(w, C) \leq \frac{\epsilon\delta}{2}$ then D_C recovers correctly the l th message bit of C with probability at least $\frac{1}{2} + \frac{\epsilon}{2}$. Let $r^{(i)} \in C$ be a closest codeword of C to w , i.e., $\delta(w, r^{(i)}) = \delta(w, C) \leq \frac{\epsilon\delta}{2}$.

For $j \in [i]$ we say that $w|_{(jm+[m])}$ is a j -block of w . We say that j -block is corrupted if $\delta(w|_{jm+[m]}, r) > \delta$. The fraction of corrupted blocks is bounded by $\frac{\epsilon}{2}$, because otherwise we have $\delta(w, r^{(i)}) > \frac{\epsilon\delta}{2}$. Recall that the decoder D_C for C picks random $j \in [i - 1]$ and invokes the decoder for R on the j -block of w . The probability that D_R will be invoked on the non-corrupted block and will recover a message bit correctly is at least $(\frac{1}{2} + \epsilon) \cdot (1 - \frac{\epsilon}{2}) \geq \frac{1}{2} + \epsilon - \frac{\epsilon}{2} = \frac{1}{2} + \frac{\epsilon}{2}$, where the inequality follows since $0 \leq \epsilon \leq \frac{1}{2}$. \square

\square

From the previous propositions we conclude the next corollary.

Corollary 4.4. *Let $R \subset \mathbb{F}_2^m$ be a linear code and $i > 0$ be an integer. Let $C \subset \mathbb{F}_2^{(mi)}$ be an (i) -repetition code of R . Then,*

- if R is not (q, ϵ, δ) -LTC then C is not (q, ϵ, δ) -LTC.
- if R is not (q, ϵ, δ) -LDC then C is not (q, ϵ, δ) -LDC.

4.3 Repetition preserves Redundancy of Testers

In this section we discuss the tightness of some of the results that appeared in [1]. We start by defining the *redundancy* of a tester.

Definition 4.5 (Distributions and Redundancy). Given a distribution \mathcal{D} over \mathbb{F}^n let the support of \mathcal{D} be $\mathcal{D}_S = \{u \in \mathbb{F}^n \mid \mathcal{D}(u) > 0\}$. We say that \mathcal{D} is uniform if for all $u_1, u_2 \in \mathcal{D}_S$ we have $\mathcal{D}(u_1) = \mathcal{D}(u_2)$. The redundancy of the distribution \mathcal{D} is defined to be $\text{Redun}(\mathcal{D}) = |\mathcal{D}_S| - \dim(\mathcal{D}_S)$.

Recall that a q -query tester for a code C is a distribution \mathcal{D} over $C_{\leq q}^\perp$. Hence $\text{Redun}(\mathcal{D})$ is used to estimate the redundancy needed to test the code C (for more information about “redundancy” see [1]).

[1] showed that if $C \subseteq \mathbb{F}^n$ is LTC, $\dim(C) \geq \omega(1)$ and \mathcal{D} its tester then $\text{Redun}(\mathcal{D}) \geq \Omega(\dim(C))$. However, the only known upper bound on $\text{Redun}(\mathcal{D})$ was $\text{Redun}(\mathcal{D}) = O(n)$.

One could conjecture that the tight lower bound should be $\text{Redun}(\mathcal{D}) \geq \Omega(n)$. We refute this in Theorem 4.6 and show that the bound $\text{Redun}(\mathcal{D}) \geq \Omega(\dim(C))$ was nearly tight.

Theorem 4.6. *For every $\omega(1) \leq f(n) \leq n/\text{poly}(\log n)$ there exists a locally testable code C with tester \mathcal{D} such that $\text{Redun}(\mathcal{D}) \leq O(f(n))$, where $\delta(C) \geq \Omega(1)$ and $\dim(C) \geq f(n)/\text{poly}(\log f(n)) \geq \omega(1)$.*

This Theorem proves that the redundancy $\text{Redun}(\mathcal{D})$ of the tester for LTC can be close to its dimension $\dim(C)$ up to a multiplicative $\text{poly}(\log f(n))$ factor for arbitrary low $f(n)$.

We first show in Proposition 4.7 that the repetition of a code preserves the redundancy of the associated tester. Next we use this to prove Theorem 4.6.

Proposition 4.7. *Let $i > 1$ be an integer. Let $R \subseteq \mathbb{F}_2^m$ be a (q, ϵ, δ) -LTC and let \mathcal{D} be its (q, ϵ, δ) -tester. Let $C \subseteq \mathbb{F}_2^{(mi)}$ be an (i) -repetition code of R . Then C has $(q, \min\{\epsilon/2, \delta/2\}, 2\delta)$ -tester \mathcal{D}' such that $\text{Redun}(\mathcal{D}'_S) = \text{Redun}(\mathcal{D}_S)$.*

Proof. Proposition 4.2 (first bullet) defined the $(q, \min\{\epsilon/2, \delta/2\}, 2\delta)$ -tester \mathcal{D}' for C . It can be verified that for all $i \in [mi] \setminus [m]$ there exists a single u such that $\mathcal{D}'(u) > 0$ and $i \in \text{supp}(u)$, moreover this unique u has $\text{supp}(u) = \{i, i \bmod m\}$. I.e., every coordinate $i \in [mi] \setminus [m]$ has its unique test $u \in \mathcal{D}'_S$. Hence the number of tests $u \in \mathcal{D}_S$ such that $\text{supp}(u) \cap [mi] \setminus [m]$ is (exactly) $mi - m$.

Furthermore, the number of tests $u \in \mathcal{D}'_S$ such that $\text{supp}(u) \subseteq [m]$ is (exactly) $|\mathcal{D}_S|$. This is true since for all $u \in \mathbb{F}_2^{(mi)}$ such that $\text{supp}(u) \subseteq [m]$ we have $u \in \mathcal{D}'_S$ if and only if $u|_{[m]} \in \mathcal{D}$, i.e., \mathcal{D}_S and \mathcal{D}'_S contains the same number of tests u with $\text{supp}(u) \subseteq [m]$. Note that for all $u \in \mathcal{D}_S$ we have $\text{supp}(u) \subseteq [m]$. We conclude that $|\mathcal{D}'_S| = |\mathcal{D}_S| + (mi - m)$.

We also know that $\dim(\text{span}(\mathcal{D}'_S)) = \dim(\text{span}(\mathcal{D}_S)) + (mi - m)$ because, as we said, every coordinate $i \in [mi] \setminus [m]$ has its unique test, which is independent on all other tests in \mathcal{D}'_S . Hence, $\text{Redun}(\mathcal{D}') = |\mathcal{D}'_S| - \dim(\text{span}(\mathcal{D}'_S)) = |\mathcal{D}_S| - \dim(\text{span}(\mathcal{D}_S)) = \text{Redun}(\mathcal{D})$. \square

In [4, 7, 14] was shown the construction of a family of LTCs of inverse poly-logarithmic rate with linear distance and only $O(1)$ queries. Then Ben-Sasson et al. [1] showed that a tester for every locally testable “uses” only $\leq 3n$ tests. Hence the following theorem was obtained.

Theorem 4.8. *There exist constants $q, \epsilon, \delta > 0$, polynomial $p(\cdot)$ and a family of (q, ϵ, δ) -LTCs $\{R_m\}_{m \in \mathbb{Z}}$ with testers \mathcal{D}_m , where $R_m \subseteq \mathbb{F}^m$, $\dim(C) \geq m/p(\log m)$ and \mathcal{D}_m is a (q, ϵ, δ) -tester for R_m . Moreover, $\text{Redun}(\mathcal{D}_m) \leq 3m$.*

We are ready to prove Theorem 4.6.

Proof of Theorem 4.6. Let $R \subseteq \mathbb{F}^m$ be a LTC as shown in Theorem 4.8 and \mathcal{D}_m be its tester such that $\text{Redun}(\mathcal{D}_m) \leq 3m$, $\dim(R) \geq m/\text{poly}(\log m)$ and $\Delta(R) = \Theta(m)$. Let $i > 0$ be an integer such that $f(mi) = \Theta(m)$ and let $n = mi$ and $C \subseteq \mathbb{F}^n$ be the (i) -repetition code of R . Proposition 4.7 implies that C is a LTC and has tester \mathcal{D}_n such that $\text{Redun}(\mathcal{D}_n) = \text{Redun}(\mathcal{D}_m) \leq 3m = O(f(mi)) = O(f(n))$. Furthermore, $\dim(C) = \dim(R) \geq m/\text{poly}(\log m) = f(n)/\text{poly}(\log f(n))$. \square

5 Proof of Main Results (Theorems 3.1,3.4)

We first prove Theorem 3.1 in Section 5.1. Then we prove Theorem 3.4 in Section 5.2.

5.1 Proof of Theorem 3.1

Proof of the first bullet. Let $\epsilon > 0$ be a constant and $m = d(n) \geq w(1)$. Let $R_\epsilon \subset \mathbb{F}_2^m$ be a linear code such that $\delta(R_\epsilon) \geq \frac{1}{2} - \epsilon$ and $\delta(R_\epsilon^\perp) \geq \Omega(1)$ and $\dim(R_\epsilon) = \Theta(m)$ (e.g., a random linear code of constant rate will have these properties). Claim 7.1 implies that R_ϵ is not locally testable with $o(m)$ queries. Lower bounds on the blocklength of locally decodable codes from [10] imply that R_ϵ is not q -query locally decodable code.

Let $C_\epsilon \subset \mathbb{F}^n$ be a (n/m) -repetition code of R_ϵ . We have $\dim(C_\epsilon) = \Theta(m) = \Theta(\log(n))$ and $\delta(C_\epsilon) \geq 1/2 - \epsilon$. Furthermore, Corollary 4.4 implies that C_ϵ is not locally testable with $o(m)$ queries and not locally decodable with q queries. \square

One could conjecture that sparse codes, which are characterized by small weight dual words and have linear distance are locally testable. We refute this next.

Proof of the second bullet. Let $m = d(n) \geq w(1)$. Then, for sufficiently large m , Theorem 7.3 implies the existence of a linear code $R \subset \mathbb{F}_2^m$ such that $\dim(R) = \Theta(d(n))$, $\Delta(R) \geq \Omega(m)$, $\text{span}(R_{\leq 3}^\perp) = R^\perp$ and R is not locally testable with $o(m)$ queries and not locally decodable with q queries.

Let $C \subset \mathbb{F}_2^n$ be a (n/m) -repetition code of R . Then Proposition 4.2 implies that C is not locally testable with $o(m) = o(d(n))$ queries. Proposition 4.3 implies that C is not locally decodable with q queries. Notice that $\text{span}(C_{\leq 3}^\perp) = C^\perp$. \square

5.1.1 Proof of the Third Bullet

Given two linear codes $C_1, C_2 \subseteq \mathbb{F}^n$, let $C_1 + C_2 = \{c_1 + c_2 \mid c_1 \in C_1, c_2 \in C_2\}$. Notice that $(C_1 + C_2)$ is a linear code.

We start from a straightforward proposition that will be useful in the next theorem.

Proposition 5.1. *Let $C_1, C_2 \subseteq \mathbb{F}_2^n$ be two linear binary codes. Then $(C_1 + C_2)^\perp = C_1^\perp \cap C_2^\perp$ and so $(C_1 + C_2)^\perp \subseteq C_1^\perp$ and $(C_1 + C_2)^\perp \subseteq C_2^\perp$.*

Proof. We have $u \in (C_1 + C_2)^\perp$ iff $(u \in C_1^\perp$ and $u \in C_2^\perp)$ iff $u \in (C_1^\perp \cap C_2^\perp)$. \square

It follows that if C_1 is a repetition code but C_2 is not then $C_1 + C_2$ is not repetition code. Moreover, if there is a ‘‘small-size’’ intersection between low-weight dual words of C_1 and of C_2 then $C_1 + C_2$ will have a small number of low-weight dual words and hence, intuitively will not be a locally testable code.

Claim 5.2. *Let $C \subseteq \mathbb{F}^n$ be a linear code and a (q, ϵ, δ) -LDC. Assume that $C' \subset C$ is a linear code (subcode of C). Then C' is a (q, ϵ, δ) -LDC.*

Proof. Assume C has the (linear) message space $S \subseteq F^k$ and has the decoder \mathbb{D} . Let $S' \subset S$ be a (linear) message space for C' . We argue that C' has the same decoder \mathbb{D} . Let w be δ -close to C' (δ -close to the encoding of some message $m \in S'$). Then w is δ -close to C and thus for all $i \in [k]$ the decoder \mathbb{D} recovers correctly the message entry (m_i) with probability at least $\frac{1}{2} + \epsilon$.

Notice that the message space S' of the subcode C' will have smaller dimension than S , i.e., $\dim(S') < \dim(S)$. S' is a linear vector space because for every two messages $x_1, x_2 \in S'$ which are encoded to $c_1, c_2 \in C'$, respectively, we have $(x_1 + x_2) \in S'$ and $(x_1 + x_2)$ is encoded to $c_1 + c_2$. For every $\alpha \in \mathbb{F}$ we also have that $\alpha x_1 \in S'$ and is encoded to αc_1 . \square

Proposition 5.3. *Let $C_1, C_2 \subseteq \mathbb{F}_2^n$ be linear codes. If C_1 is not locally decodable with q queries then $C_1 + C_2$ is not locally decodable with q queries.*

Proof. If $C_1 + C_2$ is locally decodable with q queries then C_1 is locally decodable with q queries by Claim 5.2, because C_1 is a subcode of $C_1 + C_2$. \square

Proof of the third bullet. Let $m = \log(n)$ and $R \subset \mathbb{F}_2^m$ be a linear code with $\Delta(R) \geq m/5$, $\Delta(R^\perp) \geq \Theta(m)$ and $\dim(R) = m/10$ (e.g., a random linear code of constant rate will have these properties). Claim 7.1 implies that R is not locally testable with $o(m)$ queries, and in particular, not locally testable with q queries.

Let $w \in \mathbb{F}_2^m$ be a word such that $\delta(w, R) \geq \frac{1.1}{3}$ (a random $w \in \mathbb{F}_2^m$ satisfies this condition with high probability). Notice that for all $u \in R_{\leq o(m)}^\perp$ we have $\langle u, w \rangle = 0$ because R^\perp has no words of weight less than $o(m)$, i.e., $R_{\leq o(m)}^\perp = \emptyset$.

Let $C_1 \subset \mathbb{F}_2^n$ be a (n/m) -repetition code of R . Then $\delta(w^{(n/m)}, C_1) = \delta(w, R) \geq \frac{1.1}{3}$. Notice that by construction for all $u \in C_{1, \leq o(m)}^\perp$ we have $\langle u, w^{(n/m)} \rangle = 0$. We also have $\delta(C_1) \geq 1/5$ and $\dim(C_1) = \log(n)/10$. Let $C_2 \subseteq \mathbb{F}_2^n$ be the Hadamard code³ (assume w.l.o.g. that n is a power of 2). Note that $\Delta(C_2^\perp) > 2$.

Let $\pi : [n] \mapsto [n]$ be a permutation. With some abuse of notation, for $w = (w_1, w_2, \dots, w_n) \in \mathbb{F}_2^n$ let $\pi(w) = (w_{\pi(1)}, w_{\pi(2)}, \dots, w_{\pi(n)})$ be a π -permuted word. Let $\pi(C_2) = \{\pi(c) \mid c \in C_2\}$ be a set of all permuted codewords of C_2 . Note that for every permutation $\pi : [n] \mapsto [n]$ it holds that $\delta(\pi(C_2)) = \delta(C_2)$, $\dim(\pi(C_2)) = \dim(C_2)$ and $\Delta((\pi(C_2))^\perp) = \Delta(C_2^\perp) > 2$.

Recall that $\delta(C_1, C_2) = \min_{c_1 \in C_1 \setminus \{0\}} \{\delta(c_1, C_2)\}$. We say that a permutation $\pi : [n] \mapsto [n]$ is *good* if $\delta(w^{(n/m)}, C_1 + \pi(C_2)) \geq \frac{1}{3}$ and $\delta(C_1, \pi(C_2)) \geq 1/10$.

We argue that a random permutation $\pi : [n] \mapsto [n]$ is good with probability at least $1 - o(1)$. It is sufficient to show that a random permutation $\pi : [n] \mapsto [n]$ is bad with probability at most $o(1)$.

By the Chernoff inequality the probability that for some $c_1 \in (C_1 \setminus \{0\})$ and $c_2 \in (C_2 \setminus \{0\})$ we get $\delta(w^{(n/m)}, c_1 + \pi(c_2)) < 1/3$ with probability at most $\frac{1}{2^{\Omega(n)}}$. Note that if $c_2 = 0$ then $\delta(w^{(n/m)}, c_1 + \pi(c_2)) = \delta(w^{(n/m)}, c_1) \geq \frac{1.1}{3}$ by construction.

Take a union bound over all $c_1 \in (C_1 \setminus \{0\})$ and $c_2 \in C_2$ to get that $\delta(w^{(n/m)}, C_1 + \pi(C_2)) < 1/3$ with probability at most $\frac{O(n^2)}{2^{\Omega(n)}} = o(1)$. Moreover, the probability that for given $c_1 \in (C_1 \setminus \{0\})$ and $c_2 \in C_2$ we have $\delta(c_1, \pi(c_2)) \leq 1/10$ is bounded by $\frac{1}{2^{\Omega(n)}}$. Take a union bound over all $c_1 \in (C_1 \setminus \{0\})$ and $c_2 \in C_2$ to get that the probability that $\delta(C_1, \pi(C_2)) \leq 1/10$ is bounded by $\frac{O(n^2)}{2^{\Omega(n)}} = o(1)$.

We conclude that a random permutation $\pi : [n] \mapsto [n]$ is bad with probability at most $o(1)$. So, let $\pi : [n] \mapsto [n]$ be a good permutation. Then, we have $\delta(C_1 + \pi(C_2)) = \delta(C_1, \pi(C_2)) \geq 1/10$ and $\delta(w^{(n/m)}, C_1 + C_2) \geq 1/3$. Proposition 5.1 implies that $w^{(n/m)}$ satisfies all constraints in $(C_1 + C_2)_{\leq o(m)}^\perp$ and thus $w^{(n/m)}$ will be accepted with probability 1 by any tester for $C_1 + \pi(C_2)$ with query complexity $\leq o(m)$. We conclude that $(C_1 + \pi(C_2))$ is not locally testable with $o(m)$ queries. Notice that $\dim(C_1 + \pi(C_2)) = \dim(C_1) + \dim(C_2) = 1.1 \log(n)$.

The proof for local decodability is almost the same. R is not q -query locally decodable by the lower bound on the blocklength of locally decodable codes of [10]. Corollary 4.4 implies that C_1 is not q -query locally decodable. Proposition 5.3 implies that $C_1 + \pi(C_2)$ is not q -query locally decodable. \square

³Instead of the Hadamard code we could take any binary, sparse code with linear distance and dual distance > 2 .

5.2 Proof of Theorem 3.4

Proof. In this proof we always assume that $n > 2$. Let $h(n) = o(1)$ such that $h(n) > 0$ for all $n > 2$. Without loss of generality we can assume that $h(n) \geq \frac{1}{3 \log(\log(n))}$ (otherwise let $h(n) = \frac{1}{3 \log(\log(n))}$) because if $h(n) < \frac{1}{3 \log(\log(n))}$ then $n^{-h(n)} > n^{-\frac{1}{3 \log(\log(n))}}$, so we will prove the Theorem even for a lower bias than $n^{-h(n)}$. Hence we assume that $h(n) \geq \frac{1}{3 \log(\log(n))}$.

Let $g(n) = 3h(n)$ and then $g(n) \geq \frac{1}{\log(\log(n))}$ for all $n > 2$. Let $f(n) = \frac{1}{g(n)} = \omega(1)$, note that $f(n) \leq \log(\log(n))$. Let $m = n^{g(n)}$ and $R \subseteq \mathbb{F}_2^m$ be a random linear code such that $\dim(R) = \log(m) \cdot f(n) = \log(n)$. The probability that at least one nonzero codeword of R has relative weight less than $\frac{1}{2} - m^{-1/3}$ or more than $\frac{1}{2} + m^{-1/3}$ is bounded by $\frac{2 \cdot 2^{(\log(m) \cdot f(n))}}{2^{\Omega(m^{1/3})}} = o(1)$, and this follows from the Chernoff inequality and the Union bound. Moreover, the probability that $\Delta(R^\perp) < \log(f(n)) = \omega(1)$ is bounded by $\frac{m^{\log(f(n))}}{2^{\dim(R)}} = o(1)$, and this follows from the Union bound. So, let R be a $m^{-1/3}$ -biased code such that $\dim(R) = \log(m) \cdot f(n)$ and $\Delta(R^\perp) \geq \log(f(n)) = \omega(1)$, i.e., $R_{\leq \log(f(n))}^\perp = R_{\leq \omega(1)}^\perp = \emptyset$. Assume without loss of generality that the first $\dim(R)$ bits of R are message bits.

Notice that $\log(m) \cdot f(n) = \log(n)$ and $m^{-1/3} = n^{-(1/3)g(n)} = n^{-h(n)}$. Claim 7.1 implies that R is not locally testable (decodable) with $q = O(1)$ queries.

Let $C \subseteq \mathbb{F}_2^n$ be the (n/m) -repetition code of R . We have $\dim(C) = \dim(R) = \log(n)$ and C has the same bias as R , i.e., C is $n^{-h(n)}$ -biased. In particular we have $\delta(C) \geq \frac{1}{2} - n^{-h(n)}$. Furthermore, Corollary 4.4 implies that C is not locally testable (decodable) with $q = O(1)$ queries. The Theorem follows. \square

6 Acknowledgement

We thank Tali Kaufman, Swastik Kopparty and Shubhangi Saraf for helpful discussions.

References

- [1] E. Ben-Sasson, V. Guruswami, T. Kaufman, M. Sudan, and M. Viderman, “Locally testable codes require redundant testers,” in *IEEE Conference on Computational Complexity*. IEEE Computer Society, 2009, pp. 52–61. [Online]. Available: <http://doi.ieeecomputersociety.org/10.1109/CCC.2009.6>
- [2] E. Ben-Sasson, P. Harsha, and S. Raskhodnikova, “Some 3CNF properties are hard to test,” *SIAM Journal on Computing*, vol. 35, no. 1, pp. 1–21, 2005. [Online]. Available: http://epubs.siam.org/SICOMP/volume-35/art_44544.html
- [3] E. Ben-Sasson and M. Sudan, “Robust locally testable codes and products of codes,” *Random Struct. Algorithms*, vol. 28, no. 4, pp. 387–402, 2006. [Online]. Available: <http://dx.doi.org/10.1002/rsa.20120>
- [4] E. Ben-Sasson and M. Sudan, “Short PCPs with polylog query complexity,” *SIAM J. Comput.*, vol. 38, no. 2, pp. 551–607, 2008. [Online]. Available: <http://dx.doi.org/10.1137/050646445>
- [5] E. Ben-Sasson and M. Viderman, “Composition of semi-LTCs by two-wise tensor products,” in *APPROX-RANDOM*, ser. Lecture Notes in Computer Science, I. Dinur, K. Jansen, J. Naor,

- and J. D. P. Rolim, Eds., vol. 5687. Springer, 2009, pp. 378–391. [Online]. Available: <http://dx.doi.org/10.1007/978-3-642-03685-9>
- [6] A. Deshpande, R. Jain, T. Kavitha, S. V. Lokam, and J. Radhakrishnan, “Lower bounds for adaptive locally decodable codes,” *Random Struct. Algorithms*, vol. 27, no. 3, pp. 358–378, 2005. [Online]. Available: <http://dx.doi.org/10.1002/rsa.20069>
- [7] I. Dinur, “The PCP theorem by gap amplification,” *Journal of the ACM*, vol. 54, no. 3, pp. 12:1–12:44, Jun. 2007.
- [8] O. Goldreich, “Short locally testable codes and proofs (survey),” *Electronic Colloquium on Computational Complexity (ECCC)*, no. 014, 2005. [Online]. Available: <http://eccc.hpi-web.de/eccc-reports/2005/TR05-014/index.html>
- [9] O. Goldreich and M. Sudan, “Locally testable codes and PCPs of almost-linear length,” *Journal of the ACM*, vol. 53, no. 4, pp. 558–655, Jul. 2006.
- [10] J. Katz and L. Trevisan, “On the efficiency of local decoding procedures for error-correcting codes,” in *STOC*, 2000, pp. 80–86. [Online]. Available: <http://doi.acm.org/10.1145/335305.335315>
- [11] T. Kaufman and M. Sudan, “Sparse random linear codes are locally decodable and testable,” in *FOCS*. IEEE Computer Society, 2007, pp. 590–600. [Online]. Available: <http://doi.ieeecomputersociety.org/10.1109/FOCS.2007.65>
- [12] S. Kopparty and S. Saraf, “Tolerant linearity testing and locally testable codes,” in *APPROX-RANDOM*, ser. Lecture Notes in Computer Science, I. Dinur, K. Jansen, J. Naor, and J. D. P. Rolim, Eds., vol. 5687. Springer, 2009, pp. 601–614. [Online]. Available: <http://dx.doi.org/10.1007/978-3-642-03685-9>
- [13] S. Kopparty and S. Saraf, “Local list-decoding and testing of random linear codes from high-error,” *STOC* 2010.
- [14] O. Meir, “Combinatorial construction of locally testable codes,” in *STOC*. ACM, 2008, pp. 285–294. [Online]. Available: <http://doi.acm.org/10.1145/1374376.1374419>
- [15] L. Trevisan, “Some applications of coding theory in computational complexity,” Sep. 23 2004. [Online]. Available: <http://arxiv.org/abs/cs/0409044>

7 Appendix

Claim 7.1 states that the small dual distance of the linear code $C \subseteq \mathbb{F}^n$ is necessary for its local testing and local decoding. We explain this claim now.

Ben-Sasson et al. [2] showed that a q -query tester for a locally testable code is (w.l.o.g.) a distribution over dual codewords of weight at most q . In particular, if $\Delta(C^\perp) \geq q + 1$ we conclude that C is not locally testable with q queries. Now, assume that the first $\dim(C)$ entries of the code C are message entries and $\Delta(C^\perp) \geq q + 1$. Then any local decoder which makes only $q - 1$ queries always obtains a “local view” that contains no information about the message entries and hence message entries can not be recovered with non-trivial probability.

Claim 7.1 (Folklore). *Let $C \subseteq \mathbb{F}^n$ be a linear code such that $\Delta(C^\perp) \geq (q + 1)$, where $q \geq 1$. Assume that the first $\dim(C)$ entries of C are message entries. Then C is not locally testable with q queries and not locally decodable with $q - 1$ queries.*

The other folklore claim (stated e.g. in [1]) says that every linear code is testable with query complexity equal to its dimension plus one.

Claim 7.2 (Folklore 2). *Every linear code C is testable by $\dim(C) + 1$ queries.*

Let us state the central theorem (which we rephrase) from [2]. Ben-Sasson et al. [2] showed a family of codes $C_m \subset \mathbb{F}_2^m$ which has linear distance, constant rate and was characterized by 3 weight dual words. They proved that this family is not locally testable with $o(m)$ queries. Note that this family of codes is not local decodable with constant number of queries (q) because of the lower bound on the blocklength of locally decodable codes due to Katz and Trevisan [10].

Theorem 7.3. *Let $q > 0$ be a constant integer. For infinitely many $m > 0$ there exists a family of codes $C_m \subset \mathbb{F}_2^m$ which has $\delta(C_m) = \Theta(1)$, $\dim(C_m) = \Theta(m)$ and $\text{span}((C_m)_3^\perp) = (C_m)^\perp$. Moreover, C_m is not locally testable with $o(m)$ queries and not locally decodable with q queries.*