Local List-Decoding with a Constant Number of Queries

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***Abstract*—Efremenko showed locally-decodable codes of sub-exponential length that can handle close to $\frac{1}{6}$ fraction of errors. In this paper we show that the same codes can be locally unique-decoded from error rate $\frac{1}{2} - \alpha$ for any $\alpha > 0$ and locally list-decoded from error rate $1 - \alpha$ for any $\alpha > 0$, with only a constant number of queries and a constant alphabet size. This gives the first sub-exponential length codes that can be locally list-decoded with a constant number of queries.**

***Keywords*-Locally-decodable codes; List-decoding;**

I. Introduction

Locally Decodable Codes (LDCs) are codes that allow retrieving any symbol of a message by reading only a constant number of symbols from its codeword, even if a large fraction of the codeword is corrupted. Formally, a code $\mathcal{C}$ is said to be locally decodable with parameters (α, q, ϵ) if it is possible to recover any symbol x_i of a message x by making at most q queries to $\mathcal{C}(x)$, such that even if up to a $1 - \alpha$ fraction of $\mathcal{C}(x)$ is corrupted, the decoding algorithm returns the correct answer with probability at least $1 - \epsilon$.

The first formal definition of Locally Decodable Codes was given by Katz and Trevisan in [1]. The Hadamard code is the best-known 2-query Locally Decodable Code, and its length is 2^n (where n is the message length). For 2-query LDCs tight lower bounds on the code length of $2^{\theta(n)}$ were given in [2] for linear codes and in [3] for general codes. For an arbitrary constant number of queries q, there are weak polynomial bounds, see [1], [3], [4].

The first sub-exponential LDCs (with a constant number of queries) were obtained by Yekhanin in [5]. Yekhanin obtained 3-query LDCs with sub-exponential length under a highly believable number theoretic conjecture. Later, Efremenko [6], building on Yekhanin [5] and Raghavendra [7], gave an unconditional construction of sub-exponential length LDCs. This construction also allowed a tradeoff between the number of queries and the codeword length. Unfortunately, these constructions could handle only $\frac{1}{q}$ fraction of errors (where q is the number of queries) over a large alphabet and $\frac{1}{2q}$ over the binary alphabet. In [8], Woodruff showed how to increase the handled error rate to $\frac{1}{q}$ over binary alphabets. Dvir, Gopalan and Yekhanin [9], showed how to handle $\frac{1}{4}$ fraction of errors.

Locally Decodable Codes have many applications in cryptography and complexity theory, see surveys [10], [11]. Many of these applications require LDCs that can handle high error rates. Therefore, the question of local decoding from a high error rate attracted much attention.

The goal of this paper is to construct LDCs that can handle $1 - \alpha$ fraction of errors. Clearly, when the error rate of a code is above half its distance, it is impossible to find a unique answer. Thus, we have to consider *list-decoding*. A code $\mathcal{C}$ is said to be $(1 - \alpha, L)$-list-decodable if for every word, the number of codewords within relative distance $1-\alpha$ from that word is at most L. The notion of list-decoding dates back to works by Elias [12] and Wozencraft [13] in the 50s. Roughly speaking, a code $\mathcal{C}$ is *Locally* List-Decodable if it is $(1 - \alpha, L)$-list-decodable, and given a corrupted word w, an index $k \in [L]$ and a target bit j, the decoder returns the j'th message bit of the k'th codeword that is close to w. As expected, there are some subtleties in the definition. The main issue is guaranteeing that for a fixed k, all answers for inputs (k, j) correspond to the same codeword. More formally, a local list-decoding algorithm generates L machines $\{M_k\}$, such that the machine M_k locally decodes one codeword that is close to w, and the machines $\{M_k\}$ together cover all the codewords that are close to w (for a formal definition, see Section II).

The notion of local list-decoding is a central one in the theory of computer science. It first (implicitly) appeared in the celebrated Goldreich-Levin result [14], that can be seen as a local list-decoding algorithm for the Hadamard code. Later on, many local list-decoding algorithms were studied. Most of the currently known Locally List-Decodable Codes can be divided into three categories: Reed-Muler codes [15], [16], [17], [18], direct product and XOR codes [19], [20], [21] and low-rate random codes [22]. Many of these results play an important role in Complexity Theory.

Our Results: In this paper we show how to locally list-decode the codes given in [6] (and which have sub-exponential length) with only a constant number of queries. We also show that one can uniquely decode this code up to

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radius close to $\frac{1}{2}$. The code we work with is a linear code over a finite field $\mathbb{F}$ of constant-size, i.e., $|\mathbb{F}| = f(k, \alpha) = \Theta(1)$, where f is some function. For unique local decoding we show:

Theorem 1 (Unique decoding): For every $k \geq 2, \alpha > 0$, there exists a $(\frac{1}{2} + \alpha, q, \epsilon)$ LDC of dimension n over $\mathbb{F}$ of length

$$\exp(\exp(O(\sqrt[k]{\log n (\log \log n)^{k-1}}))),$$

with $q = \Theta\left(\frac{k^k \log\left(\frac{1}{\epsilon}\right)}{\alpha^{2+k}}\right) = \Theta(1)$ queries.

Independent of our work, Dvir, Gopalan and Yekhanin in [9] show a restricted version of this theorem for $\alpha \geq \frac{1}{4}$.

For local list-decoding we show:

Theorem 2 (List-Decoding): For every $k \geq 2, \alpha > 0$, there exists a code of dimension n over $\mathbb{F}$ of length

$$\exp(\exp(O(\sqrt[k]{\log n (\log \log n)^{k-1}}))),$$

which is (α, L, q, ϵ) Locally List-Decodable Code with probabilistic reconstruction. The number of queries is $q = O(\frac{k^k \log\left(\frac{1}{\epsilon}\right)}{\alpha^{2k+1}}) = \Theta(1)$ and the list size is $L = |\mathbb{F}|^{O(\frac{\log \frac{n}{\epsilon}}{\alpha})} = \text{poly}(n)$.

It can be shown that at distance $1 - \epsilon$ there is only a constant number of codewords. The only reason that our algorithm outputs a list of non-constant size is that it requires a logarithmic-size advice.

In comparison, Reed-Muler codes are also locally list-decodable [17]. However, they are either of large length or require a non-constant number of queries:

- There are Reed-Muler codes of length $\exp(n^{\zeta})$, for any constant $\zeta > 0$, which are locally list-decodable codes with a constant number of queries.
- There are Reed-Muler codes of polynomial length which are locally list-decodable codes with a poly-logarithmic number of queries.

As we said before, the above code (from Theorems 1 and 2) is a linear code over a finite field $\mathbb{F}$ of constant-size, i.e., $|\mathbb{F}| = f(k, \alpha) = \Theta(1)$, where f is some function. We can get a Locally List-Decodable *binary* Code, by concatenating the code of Theorem 2 with a good binary code, namely,

Theorem 3: For every $k \geq 2, \alpha > 0$, there exists a *binary* code of dimension at least n and length

$$\exp(\exp(O(\sqrt[k]{\log n (\log \log n)^{k-1}}))) \cdot |\mathbb{F}|,$$

which is (α, L, q, ϵ) locally list-decodable with probabilistic reconstruction. The number of queries is $q = O(\frac{k^k \log\left(\frac{1}{\epsilon}\right)}{\alpha^{3(2k+1)}} \cdot \text{poly}(\frac{\log |\mathbb{F}|}{\alpha})) = \Theta(1)$ and the list size is $L = |\mathbb{F}|^{O(\frac{\log \frac{n}{\epsilon}}{\alpha^3})} = \text{poly}(n)$. Furthermore, if the field $\mathbb{F}$ is of characteristic two, the binary code is *linear*.

We remark that as in [6], a field $\mathbb{F}$ of characteristic 2 and of size $f(k, \alpha) \leq 2^m$ where $m = (k/\alpha)^{O(k)}$, can be used. With this field $\mathbb{F}$ the resulting binary code is *linear*. Alternatively, using the Prime Number Theorem for arithmetic progressions it can be shown that we can use a field $\mathbb{F}$ of prime order $f(k, \alpha) \approx m \log m$ (for the above m), which results in a shorter code, fewer queries and shorter output lists, but produces a *non-linear* binary code.

The rest of the paper is organized as follows: In Section II we give the necessary preliminaries. Section III gives the formal definitions of locally decodable and list-decodable codes. In Section IV we recall the construction of the code and analyze its local structure. Sections V and VI contain the proofs of Theorems 1 and 2, respectively. The proof of Theorem 3 will appear in the full version of this paper.

Related work: Gopalan [23] observes that in the Locally Decodable Codes of [6], restrictions of codewords to (multiplicative) lines are polynomials whose exponents come from a small set S. Dvir, Gopalan and Yekhanin [9] and independently we observe that for the specific set S used by the code (that originates from the work of Grolmusz [24]), the polynomials whose exponents lie in S, do not have many roots. Using this observation Dvir, Gopalan and Yekhanin [9] handle $\frac{1}{4}$ fraction of errors and we get an optimal unique decoding and local list-decoding from any constant fraction of agreement. While the results of [9] and our results were obtained independently, the results of [9] were published before ours.

II. Preliminaries

We use the following standard mathematical notation:

- $[s] = \{1, \ldots s\}$;
- $\mathbb{F}$ is a finite field;
- $\mathbb{F}^*$ is the multiplicative group of the field;
- $\mathbb{Z}_m = \mathbb{Z}/m\mathbb{Z}$, the integers modulo m;
- $\Delta(x, y)$ denotes the relative Hamming distance between vectors $x, y \in \Sigma^n$, i.e. $\Delta(w, w') = \Pr_{i \in [\bar{n}]}[w_i \neq w'_i]$;
- $\text{Ag}(w, w') \triangleq 1 - \Delta(w, w')$, i.e. $\text{Ag}(w, w') = \Pr_{i \in [\bar{n}]}[w_i = w'_i]$;
- A^B denotes the set of functions from B to A, i.e., $A^B = \{f : B \to A\}$. We identify $A^{[m]}$ with A^m.

A code is a function $\mathcal{C} : \Sigma^n \to \Sigma^{\bar{n}}$. We identify a code $\mathcal{C}$ with its image $\mathcal{C} = \{\mathcal{C}(\lambda) \mid \lambda \in \Sigma^n\}$. The distance d of the code is the minimum distance between two codewords in $\mathcal{C}$ and the *relative* distance is $\delta = d/n$. The Hamming balls of radius $\frac{d-1}{2}$ around codewords are disjoint, and therefore one can uniquely correct up to so many errors. If we allow more than $d/2$ errors several decodings are possible. In many cases one can allow much more than $d/2$ errors and still get only *few* possible decodings.

For $w \in \Sigma^{\bar{n}}$ and $\mu > 0$, define

$$\mathcal{L}_{\mathcal{C}}(w, \mu) = \{z \in \mathcal{C} : \Delta(w, z) \leq \mu\}.$$

When the code $\mathcal{C}$ is implicit from the text we abbreviate $\mathcal{L}_{\mathcal{C}}(\cdot)$ to $\mathcal{L}(\cdot)$.

Definition 1: We say that a code $\mathcal{C}$ is (μ, L) *list-decodable* if for every $w \in \Sigma^{\bar{n}}$ there are at most L codewords within distance μ from w, i.e. $|\mathcal{L}(w,\mu)| \le L$.

Fact 4 (The Johnson Bound): Let $\mathcal{C}$ be a code with relative distance δ. Then, for every $\alpha > \sqrt{1-\delta}$, the code $\mathcal{C}$ is $(1-\alpha, \frac{\alpha-(1-\delta)}{\alpha^2-(1-\delta)})$ list-decodable.

III. Locally Decodable and List-Decodable Codes

As always, one can study the combinatorial properties of a code, or ask for an explicit decoding algorithm. If the decoding algorithm makes only a few queries to the corrupted word, we say it is *local*. We begin with a formal definition of local *unique* decoding:

Definition 2: We say that a probabilistic oracle machine M^w *locally outputs a string* s *with confidence* $1-\epsilon$, if

$$\forall i \;\; \Pr[M^w(i) = s_i] \ge 1-\epsilon,$$

where the probability is taken over the randomness of M.

Definition 3 (Local Unique Decoding): A code $\mathcal{C}$ over a field $\mathbb{F}$, $\mathcal{C} : \mathbb{F}^n \mapsto \mathbb{F}^{\bar{n}}$ is said to be (α, q, ϵ) *locally decodable* if there exists a probabilistic oracle machine M^w (the decoding algorithm) with oracle access to a received codeword w such that:

1) For every message $\lambda = (\lambda_1, \lambda_2, \ldots \lambda_n) \in \mathbb{F}^n$ and for every $w \in \mathbb{F}^{\bar{n}}$ such that $\mathrm{Ag}(\mathcal{C}(\lambda), w) \ge \alpha$[1], it holds that M^w locally outputs λ with confidence $1-\epsilon$.
2) $M^w(i)$ makes at most q queries to w for all $i \in [n]$.

Recall that a code $\mathcal{C}$ is list-decodable if for every codeword w there are few codewords near w. Let $\mathcal{C}(y_1), \mathcal{C}(y_2), \ldots, \mathcal{C}(y_L)$ be the list of codewords near w. Roughly speaking, a code $\mathcal{C}$ is Locally List-Decodable if there exists a machine M that given i, j and an oracle access to the received word w, outputs the jth symbol of y_i. The locality property requires that the machine M makes a few queries to w. To make this formal:

Definition 4: Let $\mathcal{C} : \Sigma^n \to \Sigma^{\bar{n}}$ be a code. We say that a set of probabilistic oracle circuits $M_1 \ldots M_L$ with oracle queries to w, (α, L, q, ϵ) *local list-decodes* $\mathcal{C}$ at the word $w \in \Sigma^{\bar{n}}$, if,

- Every oracle circuit M_j makes at most q queries to the input word w.
- For every codeword $c = \mathcal{C}(\lambda) \in \mathcal{C}$ with $\mathrm{Ag}(c, w) \ge \alpha$, there exists some $k \in [L]$, such that M_k^w locally outputs λ with confidence $1-\epsilon$.

Definition 5: (Locally List-Decodable Codes with deterministic reconstruction) Let $\mathcal{C} : \Sigma^n \to \Sigma^{\bar{n}}$ be (α, L) list-decodable. A deterministic algorithm A (α, L, q, ϵ) *local list-decodes* $\mathcal{C}$, if on input n, A outputs probabilistic oracle circuits $M_1 \ldots M_L$ which (α, L, q, ϵ) local list-decode $\mathcal{C}$ at *every* word $w \in \Sigma^{\bar{n}}$.

The code $\mathcal{C}$ is (α, L) list-decodable and therefore every $w \in \Sigma^{\bar{n}}$ has at most L α-close codewords $c_1, \ldots, c_L$. Each such codeword $c_i = \mathcal{C}(\lambda_i)$ is represented by a probabilistic circuit M_i such that $\forall j \; M_i(j) = (\lambda_i)_j$ (recall that M_i is a probabilistic circuit, and therefore $M_i(j) = (\lambda_i)_j$ means that M_i outputs $(\lambda_i)_j$ with probability at least $1-\epsilon$). The algorithm A outputs L machines that are good for every $w \in \Sigma^{\bar{n}}$. One way to think about it is that $i \in [L]$ is an advice that tells which of the L solutions corresponds to the codeword we are interested in.

For Reed-Muller codes, a variant of the algorithm considered in [17] gives a local list-decoding algorithm with deterministic reconstruction, polynomial list size and poly-logarithmic number of queries. Often, one can reduce the list size by using probabilistic reconstruction defined as follows:

Definition 6: (Locally List-Decodable Codes with probabilistic reconstruction) Let $\mathcal{C} : \Sigma^n \to \Sigma^{\bar{n}}$ be (α, L) list-decodable. A probabilistic algorithm A (α, L, q, ϵ) *local list-decodes* $\mathcal{C}$, if on input n, A outputs probabilistic oracle circuits $M_1 \ldots M_L$ such that for every word $w \in \Sigma^{\bar{n}}$, with probability $2/3$ over the random coins of A, $M_1 \ldots M_L$ local list-decode $\mathcal{C}$ at w, i.e.,

$$\forall w \in F^{\bar{n}} \;\; \Pr_A \Bigg[\forall \lambda \; \Big(\mathrm{Ag}(\mathcal{C}(\lambda), w) \ge \alpha \;\Rightarrow \\ \exists i \; \forall j \Pr[M_i(j) = \lambda_j] \ge 1-\epsilon \Big) \Bigg] \ge 2/3.$$

Notice the order of the quantifiers: for every $w \in \Sigma^{\bar{n}}$ most of the random coins of A are good for w; however, it is not the case that most of the random coins of A are good for every w.

The local list-decoder of [14] uses probabilistic reconstruction to output a *constant* number of machines with constant query complexity, but the code length is exponential.

The local list-decoder of [17] uses a slightly different definition where the reconstruction algorithm A is allowed to access the received word w before outputting the machines that list-decode it. Of course, when such access is allowed, the number of queries to w that A performs should also be bounded (for otherwise A could simply read w and compute all the close codewords). In [17] it is required that both A and $M_1, \ldots, M_\ell$ are efficient (run in poly-logarithmic time) and this, in particular, bounds the number of queries. [17] shows a reconstruction algorithm that outputs a *constant* number of machines. It seems that with a minor modification, the [17] algorithm can work with our definitions, and use probabilistic reconstruction to output a constant number of machines that list decode with a poly-logarithmic number of queries.

In summary, [17] have polynomial code length, constant list size but poly-logarithmic number of queries, while our

[1]Note that here α denotes agreement and not distance.

code has sub-exponential length, polynomial list size and constant number of queries.

IV. The Code

In this section we define the code and study its local properties.

A. Definition of the Code

We first review the definition of the code from [6]. Fix a composite number $m = p_1 \cdot p_2 \dots p_k$ which is a product of k distinct primes. The definition of the code will depend only on m.

In order to define the code we need the following definition:

Definition 7: A family of vectors $\{u_i\}_{i=1}^n \subseteq \mathbb{Z}_m^h$ is said to be *S-matching* if the following conditions hold:

1) $\langle u_i, u_i \rangle = 0$ for every $i \in [n]$.
2) $\langle u_i, u_j \rangle \in S$ for every $i \neq j$.

Grolmusz [24] showed how to construct a large set of S-matching vectors $\{u_i\}_{i=1}^n,\ u_i \in \mathbb{Z}_m^h$, for

$$S = \{x \in \mathbb{Z}_m \setminus \{0\} \mid \forall i,\ x \bmod p_i \in \{0,1\}\}.$$

Let $\mathbb{F}$ be a field that contains an element $\gamma \in \mathbb{F}$ of order m, i.e. $\gamma^m = 1$ and $\gamma^i \neq 1$ for $i < m$. We define a code $\mathcal{C} : \mathbb{F}^n \mapsto \mathbb{F}^{m^h}$, where we think of a codeword as a function from $\mathbb{Z}_m^h$ to $\mathbb{F}$. The encoding of the message $\lambda_1, \lambda_2 \dots \lambda_n$ is the function:

$$\mathcal{C}(\lambda_1, \lambda_2, \dots, \lambda_n)(x) \triangleq \sum_{i=1}^n \lambda_i \gamma^{\langle u_i, x \rangle}.$$

Equivalently, we can write

$$\mathcal{C}(\lambda_1, \lambda_2, \dots, \lambda_n) = \sum_{i=1}^n \lambda_i f_i, \tag{1}$$

where $f_i(x) \triangleq \gamma^{<u_i, x>}$. We denote the codeword length by $\bar{n} = m^h$. An asymptotic relation between n and $\bar{n}$ is:

$$\bar{n} = \exp(\exp(O(\sqrt[k]{\log n (\log\log n)^{k-1}}))).$$

Note that the asymptotic rate of the code depends only on k, the number of different primes dividing m.

For simplicity, sometimes we denote $G \triangleq \mathbb{Z}_m^h$.

B. Local Properties of the Code

In this subsection we study local properties of the code. Specifically, we study the restriction of the code to lines.

Definition 8: (line) Let $v, u \in G$. The *line through v in direction u* is the function $\ell = \ell_{v,u} \in G^{[m]}$ defined by $\ell(t) = v + tu$.

Definition 9 (restriction): Let $\ell \in G^{[m]}$ be a line.

- For a function $f \in \mathbb{F}^G$, *the restriction of f to ℓ*, denoted by $f|_\ell \in \mathbb{F}^{[m]}$ is defined by $f|_\ell(t) = f(\ell(t))$.
- For a code $\mathcal{C} : \mathbb{F}^n \to \mathbb{F}^G$, *the restriction of $\mathcal{C}$ to ℓ*, denoted by $\mathcal{C}|_\ell : \mathbb{F}^n \to \mathbb{F}^m$, is the vector space $\{\mathcal{C}(\lambda)|_\ell \mid \lambda \in \mathbb{F}^n\}$.

Now, we analyze the restriction of the code in direction u_j. Observe that

$$\begin{aligned}\mathcal{C}(\lambda_1, \dots, \lambda_n)(v + tu_j) &= \sum_i \lambda_i \gamma^{\langle u_i, v+tu_j \rangle} \\ &= \sum_i \lambda_i \gamma^{\langle u_i, v \rangle} (\gamma^{\langle u_i, u_j \rangle})^t \\ &= \sum_{b \in S \cup \{0\}} \left[\sum_{i:\langle u_i, u_j \rangle = b} \lambda_i \gamma^{\langle u_i, v \rangle} \right] (\gamma^t)^b.\end{aligned}$$

Define $p : \mathbb{F} \to \mathbb{F}$ by $p(x) = \sum_{b \in S \cup \{0\}} a_b x^b$, where $a_b = \sum_{i:\langle u_i, u_j \rangle = b} \lambda_i \gamma^{\langle u_i, v \rangle}$, then $\mathcal{C}|_{\ell_{v,u_j}}(\lambda)(t) = p(\gamma^t)$. Furthermore, $a_0 = \lambda_j \gamma^{\langle u_j, v \rangle}$, and so when $\lambda_j \neq 0$, p is a non-zero polynomial.

The observation that a codeword restricted to a line is a polynomial whose free coefficient encodes λ_j appears in [23]. We now prove (Lemma 5) that this polynomial does not have too many roots and therefore the code restricted to the line has a large distance. This Lemma was also independently found by Dvir, Gopalan and Yekhanin [9].

Lemma 5: Let $\mathcal{C}$ be the code above. For every $v \in G$ and $j \in [n]$, the code $\mathcal{C}|_{\ell_{v,u_j}} : \mathbb{F}^n \to \mathbb{F}^m$ is of dimension at most 2^k and distance $\delta \geq 1 - \sum_{i=1}^k \frac{1}{p_i}$.

Proof: In order to prove the lemma we need to show that the polynomial $p(x) = \sum_{b \in S \cup \{0\}} a_b x^b$ does not have too many roots in the set $H \triangleq \{\gamma^i | 0 \leq i < m\}$. Recall that the set S is

$$S = \{x \in \mathbb{Z}_m \setminus \{0\} \mid \forall i\ x \bmod p_i \in \{0,1\}\}.$$

Notice that p might have a large degree, and therefore might have a large number of roots in $\mathbb{F}$. Nevertheless, we show that the number of roots p has in H is at most $\sum_i \frac{m}{p_i}$. To see that denote $\tilde{p}(x) = p(x^{\sum_i \frac{m}{p_i}})$. We show that $\tilde{p}$ has the same number of roots as p. Let $s = \sum_i \frac{m}{p_i}$. Then,

$$s \pmod{p_i} = \frac{m}{p_i} \pmod{p_i} \neq 0.$$

Therefore, $\gcd(s, m) = 1$, that is, s is invertible in $\mathbb{Z}_m$. This implies the mapping $\psi : H \to H,\ \psi(x) = x^s$ is a bijection.

Thus, in order to show p has few roots in H, it suffices to show that $\tilde{p}$ is a low-degree polynomial. Each monomial of $\tilde{p}$ is of degree $b \cdot s \pmod m$ for some $b \in S \cup \{0\}$. Notice that for every $1 \leq i, j \leq k$,

$$\frac{m}{p_i} \cdot b \bmod p_j \quad = \quad \begin{cases} 0 & j \neq i \\ 0 & (j = i) \wedge \\ & \quad (b \bmod p_i = 0) \\ \frac{m}{p_i} \bmod p_i & (j = i) \wedge \\ & \quad (b \bmod p_i = 1) \end{cases}$$

This implies that for every i,

$$\frac{m}{p_i} \cdot b \bmod m \quad = \quad \begin{cases} 0 & b \bmod p_i = 0 \\ \frac{m}{p_i} & b \bmod p_i = 1 \end{cases}$$

Hence, $b \cdot s \pmod m \leq \sum_i \frac{bm}{p_i} \pmod m \leq \sum_i \frac{m}{p_i}$. We conclude that $\tilde{p}$ has at most $\sum_i \frac{m}{p_i}$ roots in H and therefore so does p.

For a polynomial $p : \mathbb{F} \to \mathbb{F}$ define the vector $\overline{p} \in \mathbb{F}^m$ by $\overline{p}(t) = p(\gamma^t)$. Then, $\mathcal{C}|_{\ell_{v,u_j}}$ is a linear subspace of the vector-space $\mathrm{Span}\left\{\overline{x^b} : b \in S \cup \{0\}\right\}$, which is a dimension 2^k $\mathbb{F}$-subspace. Every non-zero codeword corresponds to a non-zero polynomial that can have at most $\sum_i \frac{m}{p_i}$ roots. As the elements γ^t are distinct for $1 \leq t \leq m$, every codeword has at most that many zeroes. ■

Let $\mathcal{C}$ be the code above. Let $v \in G$ and $j \in [n]$. Then every codeword of $\mathcal{C}|_{\ell_{v,u_j}}$ corresponds to a polynomial with 2^k monomials, where the free coefficient is $\lambda_j \gamma^{\langle u_j, v\rangle}$. Thus, any restricted codeword $z \in \mathcal{C}|_{\ell_{v,u_j}}$ contains information about λ_j.

Definition 10: Using the above notation, we denote $D_{v,j}(z) = \lambda_j$.

In particular,

Corollary 6: Let $\mathcal{C}$ be the code above. Let $v \in G$ and $j \in [n]$. If $z, z' \in \mathcal{C}|_{\ell_{v,u_j}}$ and $D_{v,j}(z) \neq D_{v,j}(z')$ then $z \neq z'$ and therefore $\Delta(z, z') \geq \delta$.

Another corollary is,

Corollary 7: The distance of the code $\mathcal{C}$ is at least δ, where $\delta = 1 - \sum_{i=1}^{k} \frac{1}{p_i}$.

Proof: Look at two different codewords $\mathcal{C}(\lambda)$ and $\mathcal{C}(\tilde{\lambda})$ for some $\lambda \neq \tilde{\lambda}$. Then, there exists some $j \in [n]$ such that $\lambda_j \neq \tilde{\lambda}_j$. We can now partition G to disjoint lines in direction u_j. From Corollary 6 it follows that on each of these lines the restrictions of $\mathcal{C}(\lambda)$ and $\mathcal{C}(\tilde{\lambda})$ are different. From Lemma 5 we know that the distance on each of these lines is at least δ. It follows that the distance between $\mathcal{C}(\lambda)$ and $\mathcal{C}(\tilde{\lambda})$ is at least δ. ■

Remark 8: By taking all p_i's of the same order we get that $\delta = 1 - O(\frac{k}{\sqrt[k]{m}})$. In this paper we assume that m is such a product.

V. Local Unique Decoding

We are given some word $w \in \mathbb{F}^G$ that has agreement $\frac{1}{2}+\alpha$ with some codeword $\mathcal{C} = \mathcal{C}(\lambda)$. We are also given some $j \in [n]$. Our goal is to recover (with a good probability) λ_j. A first attempt at local decoding is restricting the code to a random line ℓ_{v,u_j} in direction u_j. Intuitively, this is a good step because we restrict the global code to a small fragment of constant size m, while still keeping information about λ_j. Specifically, by Lemma 5, $\mathcal{C}|_{\ell_{v,u_j}}$ is a linear code with a large distance, and by Corollary 6, a codeword $z = \mathcal{C}(\lambda)|_{\ell_{v,u_j}} \in \mathcal{C}|_{\ell_{v,u_j}}$ corresponds to a polynomial with 2^k monomials, where the free coefficient is $\lambda_j \gamma^{\langle u_j, v\rangle}$.

As $\mathcal{C}(\lambda)$ has $\frac{1}{2} + \alpha$ agreement with w, when we pick a random line in direction u_j, the expected agreement between $w|_{\ell_{v,u_j}}$ and $\mathcal{C}(\lambda)|_{\ell_{v,u_j}}$ is $\frac{1}{2} + \alpha$. The problem is that it may still happen that with high probability the agreement between $w|_{\ell_{v,u_j}}$ and $\mathcal{C}(\lambda)|_{\ell_{v,u_j}}$ is less then $\frac{1}{2}$ and we will decode a wrong value. In order to overcome this problem we sample $K = O(\frac{\log\left(\frac{1}{\epsilon}\right)}{\alpha^2})$ independent lines. Then with high probability the agreement between w and $\mathcal{C}(\lambda)$ is at least $\frac{1}{2} + \frac{\alpha}{2}$ on the sampled lines. Note that the code $\mathcal{C}(\lambda)$ restricted to the union of independent lines in direction u_j may not have a good distance, as two different codewords may coincide on a restriction to a line. However, for any two codewords $\mathcal{C}(\lambda)$ and $\mathcal{C}(\tilde{\lambda})$, where $\lambda_j \neq \tilde{\lambda}_j$, the distance between the restrictions of these two codewords on *each line* must be large because of Corollary 6.

Let $\alpha \geq 2(1-\delta)$ (where δ is the distance of the code, and by Lemma 5 is at least $1 - \sum_i \frac{1}{p_i}$). The unique decoding algorithm for $\frac{1}{2} + \alpha$ agreement is as follows:

- **Input:**
  - $w \in \mathbb{F}^G$ that has agreement $\frac{1}{2} + \alpha$ with some codeword $\mathcal{C}$,
  - $j \in [n]$,
  - $\epsilon > 0$
- **Randomness:** A set of $K = \Theta(\frac{\log\left(\frac{1}{\epsilon}\right)}{\alpha^2})$ random elements in G, $\overline{v} = (v_1, \ldots, v_K) \in G$.
- **Queries:** For each $k \in [K]$, the algorithm queries all points on the line ℓ_{v_k,u_j}.
- **Algorithm:** For every $k \in [K]$ and for every symbol $\sigma \in \mathbb{F}$, the algorithm computes

$$\mathrm{weight}_k(\sigma) = \max\Big\{\mathrm{Ag}(w, z) \ : \ z \in \mathcal{C}|_{\ell_{v_k,u_j}}, D_{v_k,j}(z) = \sigma\Big\}.$$

  The algorithm then computes $\mathrm{weight}(\sigma) = \frac{1}{K}\sum_{k=1}^{K} \mathrm{weight}_k(\sigma)$. The output of the algorithm is the symbol σ with the highest weight.

Theorem 9: Assume $\alpha \geq 2(1-\delta)$. For every $\lambda \in \mathbb{F}^n, w \in \mathbb{F}^G$ with $\mathrm{Ag}(w, \mathcal{C}(\lambda)) \geq \frac{1}{2} + \alpha$ and every $j \in [n]$,

$$\Pr_{\overline{v}}[\text{The algorithm outputs } \lambda_j] \geq 1 - \epsilon.$$

The algorithm uses $\Theta(\frac{\log\left(\frac{1}{\epsilon}\right)}{\alpha^2} \cdot m)$ queries.

Proof: Suppose that $\mathcal{C}(\lambda)$ is a codeword which has $\frac{1}{2}+\alpha$ agreement with the received word w. Then

$$\mathbb{E}_{v \in G}\left[\mathrm{Ag}(w|_{\ell_{v,u_j}}, \mathcal{C}(\lambda)|_{\ell_{v,u_j}})\right] = \frac{1}{2} + \alpha.$$

We say $\overline{v} = (v_1, \ldots, v_k)$ is *good*, if

$$\frac{1}{K}\sum_{k=1}^{K}\left[\mathrm{Ag}(w|_{\ell_{v_k,u_j}}, \mathcal{C}(\lambda)|_{\ell_{v_k,u_j}})\right] \geq \frac{1+\alpha}{2}.$$

By a standard application of the Chernoff Bound,

$$\Pr_{\overline{v}}[\overline{v} \text{ is not good}] \le 2^{-\Omega(\alpha^2 K)} = \epsilon.$$

We now prove that if $\overline{v}$ is good the algorithm outputs the correct answer.

Denote $ag_k = \mathrm{Ag}(w|_{\ell_{v_k,u_j}}, \mathcal{C}(\lambda)|_{\ell_{v_k,u_j}})$. Then,

- For every k, $\mathrm{weight}_k(\lambda_j) \ge \mathrm{Ag}(w|_{\ell_{v,u_j}}, \mathcal{C}(\lambda)|_{\ell_{v,u_j}}) \ge ag_k$ and so $\mathrm{weight}(\lambda_j) \ge \mathbb{E}_k[ag_k] \ge \frac{1+\alpha}{2}$.
- Fix any $\sigma \ne \lambda_j$ and $k \in [K]$. Let $z \in \mathcal{C}|_{\ell_{v_k,u_j}}$ be such that $D_{v_k,j}(z) = \sigma$. Then, by the triangle inequality,

$$\delta \le \Delta(z, \mathcal{C}(\lambda)|_{\ell_{v_k,u_j}}) \le \Delta(\mathcal{C}(\lambda)|_{\ell_{v_k,u_j}}, w|_{\ell_{v_k,u_j}}) + \Delta(w|_{\ell_{v_k,u_j}}, z).$$

Thus, $\Delta(w|_{\ell_{v_k,u_j}}, z) \ge \delta + ag_k - 1$, and $\mathrm{weight}_k(\sigma) \le 1 - ag_k + 1 - \delta$. In particular,

$$\mathrm{weight}(\sigma) \le 1-\delta+\mathbb{E}_k[1-ag_k] \le \frac{1}{2}+1-\delta-\frac{\alpha}{2} \le \frac{1}{2}.$$

Thus, whenever $\overline{v}$ is good the algorithm outputs λ_j. ■

We are now ready to prove Theorem 1.

Proof of Theorem 1: The code $\mathcal{C}$ has distance at least $\delta = 1 - O(\frac{k}{m^{1/k}})$ and the code length is

$$\exp(\exp(O(\sqrt[k]{\log n(\log\log n)^{k-1}}))).$$

We take m to be a product of m almost equal primes. From Theorem 9, for every $\alpha \ge 2(1-\delta) = O(\frac{k}{m^{1/k}})$, the code is $(\frac{1}{2}+\alpha, q, \epsilon)$ locally decodable with $q = \Theta(\frac{\log\left(\frac{1}{\epsilon}\right)}{\alpha^2}\cdot m)$ queries. We think of k as a constant, and m as depending on α, growing to accommodate the required error rate. Thus $\alpha = 2(1-\delta) \approx \frac{2k}{m^{1/k}}$, or equivalently, $m \approx (\frac{2k}{\alpha})^k$. Thus, the number of queries is $\Theta(\frac{m\log\left(\frac{1}{\epsilon}\right)}{\alpha^2}) = \Theta(k^k \cdot \alpha^{-(k+2)} \cdot \log\left(\frac{1}{\epsilon}\right))$. For $k = 2$ the number of queries is $\Theta(\alpha^{-4} \cdot \log\left(\frac{1}{\epsilon}\right))$. ■

VI. Local list-decoding with probabilistic reconstruction

We first remind the reader of the setting. A probabilistic algorithm A has to produce L probabilistic circuits $M_1, \ldots, M_L$ that (α, L, q, ϵ) local list-decode $\mathcal{C}$. A uses its internal random coins to sample a random subset $\Lambda \subseteq G$ of cardinality $\Theta(\frac{\log\frac{n}{\epsilon}}{\alpha})$. Notice that $|\Lambda|$ is super-constant. The list size L is $\left|\mathbb{F}^{\Lambda}\right|$ and corresponds to all possible values a codeword may take on Λ. We identify an index of a machine $i \in [L]$ with a function $\mathsf{ad} : \Lambda \mapsto \mathbb{F}$ of values of a codeword on the set Λ. The machine M^w_{ad} locally outputs a message λ such that $\mathcal{C}(\lambda)$ has α agreement with w and $\mathsf{ad} = \mathcal{C}(\lambda)|_{\Lambda}$, if such a λ exists.

Given a corrupted word $w \in \mathbb{F}^G$ and a value $j \in [n]$, M_{ad}'s goal is to find (the hopefully unique) codeword $c \in \mathcal{C}$ that is α close to w, and that is consistent with the given advice $\mathsf{ad} \in \mathbb{F}^{\Lambda}$. To do so, M_{ad} does the following: M_{ad} picks K (and K will turn out to be constant even though $|\Lambda|$ is not a constant) random lines in direction u_j that pass through some point in Λ. For each such line, M_{ad} queries w on the line, and finds all the restricted codewords that are close to the given w (on the line). We say that a line is good if among all those codewords, *exactly* one matches the value ad gives on the point from Λ. For each good line, M_{ad} extracts from this unique codeword the value λ_j and adds it to the candidate list. The output of M_{ad} is the most common value in the candidate list. More formally, the algorithm M_{ad} is defined as follows:

- **A's random coins:** A random subset Λ of cardinality $\Theta(\frac{\log\frac{n}{\epsilon}}{\alpha})$.
- **Advice:** Values of some codeword c on Λ.
- **Input:** $w \in \mathbb{F}^G$, $j \in [n]$.
- **M's randomness:** A random subset $\{s_1, \ldots, s_K\}$ of Λ of cardinality $K = \Theta(\frac{\log\left(\frac{1}{\epsilon}\right)}{\alpha})$.
- **Queries:** For each $k \in [K]$, M queries the values of w on the K lines ℓ_{s_k,u_j}.
- **Algorithm:** For every k, the algorithm goes over all codewords of $\mathcal{C}' = \mathcal{C}|_{\ell_{s_k,u_j}}$. For every such k, if there exists *exactly* one codeword z of $\mathcal{C}'$ with:
  - $\mathrm{Ag}(z, w|_{\ell_{s_k,u_j}}) \ge \frac{\alpha}{2}$, and,
  - $z(s_k) = \mathsf{ad}(s_k)$

  then the algorithm adds the value $D_{v_k,j}(z)$ to the candidates list.
- **Output:** The most common value in the candidates list.

Theorem 10: For any $\alpha \ge 8\sqrt{1-\delta}$, $\epsilon > 0$ and $L = \left|\mathbb{F}^{\Lambda}\right| = q^{O(\frac{\log\frac{n}{\epsilon}}{\alpha})}$, $q = Km = O(\frac{m\log\left(\frac{1}{\epsilon}\right)}{\alpha}) = O(\log\left(\frac{1}{\epsilon}\right))$. The above algorithm is a probabilistic polynomial-time (α, L, q, ϵ) local list-decoding algorithm.

Theorem 2 follows immediately from Theorem 10.

Proof of Theorem 2: We take m a product of k distinct almost equal primes. From Theorem 10 we know that for any $\alpha > 8\sqrt{1-\delta} = O(\frac{\sqrt{k}}{\sqrt[2k]{m}})$ the code is (α, L, q, ϵ) local list-decodable with $q = O(\frac{m\log\left(\frac{1}{\epsilon}\right)}{\alpha})$. Therefore, $m = O(\frac{k^k}{\alpha^{2k}})$ and $q = O(k^k \cdot \alpha^{-(2k+1)} \cdot \log\left(\frac{1}{\epsilon}\right))$ with a codeword length:

$$\exp(\exp(O(\sqrt[k]{\log n(\log\log n)^{k-1}})))$$

■

We are left to prove Theorem 10.

A. Proof of correctness

We need to show that for every received word w, with high probability over the choice of the set Λ, for every codeword $c = \mathcal{C}(\lambda)$ that has α agreement with w, when the advice is $\mathsf{ad} = c|_{\Lambda}$, it holds that for every $j \in [n]$, $\Pr[M^w_{\mathsf{ad}}(j) = \lambda_j] \ge 1 - \epsilon$, where the probability is over the randomness of M.

Fix $w \in \mathbb{F}^G$, a codeword $c = \mathcal{C}(\lambda)$ and $j \in [n]$. For $v \in G$ the machine M_{ad} considers the set $U_j(v) =$

$\left\{z \in \mathcal{C}|_{\ell_{v,u_j}} : (\mathrm{Ag}(z, w|_{\ell_{v,u_j}}) \geq \alpha/2) \wedge (z(v) = \mathsf{ad}(v))\right\}$. In the kth iteration, if $U_j(s_k) = \left\{c|_{\ell_{s_k,u_j}}\right\}$ then M_{ad} adds λ_j to the candidates list.

We say that v is *useful* if $c|_{\ell_{v,u_j}} \in U_j(b)$. Notice that $c|_{\ell_{v,u_j}}(v) = \mathsf{ad}(v)$, hence for v to be useful we only need a high agreement between v and w on the line ℓ_{v,u_j}. We say that v *filters* if $U_j(v) \subseteq \left\{c|_{\ell_{v,u_j}}\right\}$, i.e., for any codeword in the restricted code $z \in \mathcal{C}|_{\ell_{v,u_j}}$ such that $z \neq c$ it holds that $z \notin U_j(v)$.

Lemma 11: For any $\alpha \geq 8\sqrt{1-\delta}$ it holds that

- $\Pr_{v \sim G}[v \text{ is useful}] \geq \frac{\alpha}{2}$
- $\Pr_{v \sim G}[v \text{ does not filter}] \leq \frac{4}{\alpha} \cdot (1-\delta) \leq \frac{\alpha}{16}$.

Proof: Since

$$\mathbb{E}_v[\mathrm{Ag}(w|_{\ell_{v,u_j}}, c|_{\ell_{v,u_j}})] = \alpha,$$

an averaging argument implies that the probability $v \in G$ is useful is at least $\alpha/2$.

We turn to the second item. A point v does not filler if there is a restricted codeword $z \in \mathcal{C}|_{\ell_{v,u_j}}$ such that $z \neq c|_{\ell_{v,u_j}}$ and $z \in U_j(v)$. A restricted codeword z is in $U_j(v)$ if it is in the list $\mathcal{L}(w|_{\ell_{v,u_j}}, \alpha/2)$ and $z(v) = c(v)$. One way to choose v uniformly from G is by first choosing a random line ℓ in direction u_j, and then choosing a random point v on the line. For any line ℓ in direction u_j, $\mathcal{C}' = \mathcal{C}|_\ell$ has distance δ. Therefore, for any $z \neq c$ the probability that $z(v) = c(v)$ is at most $1-\delta$. By the Johnson bound (see Fact 4), the number of codewords with $\alpha/2$ agreement with $w|_\ell$ satisfies

$$\left|\mathcal{L}(w|_{\ell_{v,u_j}}, \alpha/2)\right| \leq \frac{\alpha/2 - (1-\delta)}{\alpha^2/4 - (1-\delta)} < \frac{4}{\alpha},$$

when $\alpha \geq 2\sqrt{2(1-\delta)}$. The probability that such a codeword z agrees with c at v is at most $1-\delta$. The lemma follows from the union bound. ∎

Definition 11: For $w \in \mathbb{F}^G$, a set $\Lambda \subseteq G$ is *good for* w, if for every $c \in \mathcal{L}(w, \alpha/2)$ and every $j \in [n]$,

- $\Pr_{v \in \Lambda}[v \text{ is useful and filters for } (w, c, j)] \geq \frac{\alpha}{4}$.
- $\Pr_{v \in \Lambda}[v \text{ does not filter } (w, c, j)] \leq \frac{\alpha}{8}$.

Lemma 12: Fix $w \in \mathbb{F}^G$. Pick a set Λ uniformly at random from G. The probability Λ is not good for w is at most $\frac{n}{\alpha} \cdot 2^{-\Omega(\alpha|\Lambda|)}$.

Proof: For any w, j and $c \in \mathcal{L}(w, \alpha)$, the probability that a single v is useful and filters, by Lemma 11, is at least $\frac{\alpha}{3}$. By the Chernoff bound, the probability we do not have $\frac{\alpha}{4}$ fraction of good vectors in the sample set Λ is at most $2^{-\Omega(\alpha|\Lambda|)}$.

Similarly, by Lemma 11, for any w, j and $c \in \mathcal{L}(w, \alpha)$, the probability a single v does not filter (w, c, j), is at most $\frac{\alpha}{16}$. By the Chernoff Bound, the probability that we have more than $\frac{\alpha}{8}$ fraction of vectors that do not filter (w, c, j) in the sample Λ is at most $2^{-\Omega(\alpha|\Lambda|)}$.

The lemma follows from a union bound over j and $c \in \mathcal{L}(w, \alpha)$. ∎

Assume Λ is good for w. The probability that at the ith iteration, M_{ad} adds the correct value λ_j to the candidates list is at least the probability that v is useful and filters. By Definition 11 this probability is at least $\frac{\alpha}{4}$. The probability that M_{ad} adds a wrong value to the candidates list is bounded by the probability that v does not filter, which is at most $\frac{\alpha}{8}$. Therefore, by the Chernoff bound, it follows that after $\Theta(\frac{\log(\frac{1}{\epsilon})}{\alpha})$ iterations the probability that λ_j is the most common value in the candidates list is at least $1-\epsilon$. Theorem 10 follows from the above lemma, since for every w, Λ is good for w with probability at least ϵ (by the choice of the cardinality of Λ).

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