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# CS 761 : DERANDOMIZATION AND PSEUDORANDOMNESS

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## Contents

<b>1</b>	<b>Introduction</b>	<b>3</b>
1.1	Lecture 1: Matrix multiplication . . . . .	3
1.2	Lectures 3–4: Pairwise independence . . . . .	4
1.2.1	Lecture 3 . . . . .	4
1.2.2	Lecture 4 . . . . .	5
1.3	Lectures 4–5: Counting distinct elements in a stream . . . . .	6
1.3.1	Lecture 4 (continued) . . . . .	6
1.3.2	Lecture 5 . . . . .	6
<b>2</b>	<b>Expander graphs and applications</b>	<b>8</b>
2.1	Lectures 6–7: Magical graphs and two applications . . . . .	8
2.1.1	Lecture 6 . . . . .	8
2.1.2	Lecture 7 . . . . .	10
2.2	Lectures 7, 8, 10: Testing connectivity of undirected graphs . . . . .	11
2.2.1	Lecture 7 (continued) . . . . .	11
2.2.2	Lecture 8 . . . . .	12
2.2.3	Lecture 10 . . . . .	14
2.3	Lecture 9: Expander graphs . . . . .	16
2.3.1	Lecture 9 . . . . .	16
<b>3</b>	<b>Reed-Solomon Codes</b>	<b>20</b>
3.1	Lecture 11: Introduction . . . . .	20
3.1.1	Lecture 11 . . . . .	20
3.2	Lecture 11-12: Decoding . . . . .	21
3.2.1	Lecture 11 (continued) . . . . .	21
3.2.2	Lecture 12 . . . . .	22
3.3	Lecture 14: Multiplicity codes . . . . .	24
3.3.1	Lecture 12 (continued) . . . . .	24
3.3.2	Lecture 14 . . . . .	24
<b>4</b>	<b>Hardness v. Randomness</b>	<b>27</b>
4.1	Lecture 13, 15: Hardness . . . . .	27
4.1.1	Lecture 13 . . . . .	27
4.1.2	Lecture 15 . . . . .	29
4.2	Lectures 15, 16: Average-case hardness to derandomization . . . . .	30
4.2.1	Lecture 15 (contd.) . . . . .	30
4.2.2	Lecture 16 . . . . .	30

4.3	Lecture 17: Worst-case hardness to average-case hardness	32
4.3.1	Lecture 17	32
4.3.2	Lecture 18	33

## §1. Introduction

### 1.1. Lecture 1: Matrix multiplication

We begin with a question.

**Problem.** Given three  $n \times n$  matrices  $A, B, C$ , decide whether  $AB = C$ .

One naïve way to do this is to compute  $AB$  and check if it is identical to  $C$ . The naïve implementation of this runs in  $O(n^3)$ , while the best known implementation at the time runs in about  $O(n^{2.373\dots})$ .

Can we do the required in  $O(n^2)$  time, perhaps in a random fashion (with some probability of failure)?

Consider the following algorithm to start with. For each row in  $C$ , choose an entry randomly and verify that it matches the corresponding entry in  $AB$ . In a similar spirit, a second algorithm is to choose  $n$  entries of  $C$  randomly and verify.

If  $AB = C$ , it is clear that no matter how we choose to test, we shall return that the two are indeed equal. The probability we would like to minimize is

$$\Pr[\text{the algorithm outputs yes} \mid AB \neq C].$$

Of course, this probability depends on  $A, B, C$ . This probability is over the randomness inherent in the algorithm, not in some choosing of  $A, B, C$ .

When  $AB$  and  $C$  differ at only one entry, the earlier proposed algorithm has a success probability of  $1/n$  (so the quantity mentioned above is  $1 - 1/n$ ). This is very bad, as it means that to reduce the failure probability to some constant, we would need to repeat this  $n$  times.

An algorithm that does the job is as follows.

Randomly choose  $r \in \{0, 1\}^n$ . Compute  $ABr$  and  $Cr$ , and verify that the two are equal. This is an  $O(n^2)$  algorithm, since multiplying a matrix with a vector takes  $O(n^2)$  and we perform this operation thrice, in addition to an  $O(n)$  verification step at the end.

We claim that the failure probability of this algorithm is at most  $1/2$ .

The failure probability can be rephrased as follows. Let  $x, y \in \mathbb{R}^{1 \times n}$ . What is  $\Pr[xr = yr \mid x \neq y]$ ? The earlier failure probability is at most equal to this, with equality attained (in a sense) when the two matrices differ at exactly one row.

This in turn is equivalent to the following. Let  $z \in \mathbb{R}^{1 \times n}$ . What is  $\Pr[ zr = 0 \mid z \neq 0 ]$ ? Suppose that  $z_i \neq 0$  for some  $i$ . For any choice of the remaining  $n - 1$  bits, at most one of the two options for the  $i$ th bit can result in  $zr = 0$ .

Let us do this slightly more formally. Assume wlog that  $z_n \neq 0$ . Then,

$$\begin{aligned} \Pr[z_1 r_1 + \dots + z_n r_n = 0 \mid z_n \neq 0] &= \Pr\left[r_n = -\frac{z_1 r_1 + \dots + z_{n-1} r_{n-1}}{z_n} \mid z_n \neq 0\right] \\ &\leq \max_{r_1, \dots, r_{n-1}} \Pr\left[r_n = -\frac{z_1 r_1 + \dots + z_{n-1} r_{n-1}}{z_n} \mid z_n \neq 0, r_1, \dots, r_{n-1}\right] \end{aligned}$$

which is plainly at most  $1/2$  – we cannot have that both 0 and 1 are equal to the quantity of interest!

*Remark.* If we instead choose  $r$  from  $\{0, 1, \dots, q - 1\}^n$  instead, the failure probability now goes down at most  $1/q$ . There is a tradeoff at play here between the reduction in the failure probability and the increase in the number of random bits (it goes from  $n$  to  $O(n \log q)$ ).

**Question.** Can we reduce the number of random bits in this algorithm? Can we make it deterministic?

To answer the question of determinism, suppose the algorithm designer chooses  $k$  vectors  $r^{(1)}, \dots, r^{(k)} \in \mathbb{R}^n$  and tests whether  $ABr^{(i)} = Cr^{(i)}$ . This will fail if  $k < n$ . Indeed, an adversarial input is a  $z$  that is nonzero but with  $zr^{(i)} = 0$  for  $1 \leq i \leq k$ .

The determinism here is in the sense that the vectors are chosen before the inputs are provided.

On the other hand, we *can* reduce the number of random bits used. In fact, we can go to about  $O(\log n)$  random bits. The goal of derandomization is to use a smaller number of random bits (perhaps by conditioning together previously independent bits), without losing the power of the earlier independent bits.

Let

$$A(x) = a_0 + a_1x + \cdots + a_dx^d$$

be a nonzero polynomial of degree  $d$ . Choose  $x$  randomly from  $\{0, 1, \dots, q-1\}$ . It is not difficult to see that

$$\Pr_{x \sim \{0, 1, \dots, q-1\}} [A(x) = 0] \leq \frac{d}{q}.$$

Inspired by this, we can reduce randomness as follows. Choose  $x$  randomly from  $\{0, 1, \dots, 2n-1\}$ , and set  $r = (1, x, x^2, \dots, x^{n-1})$ . Then,

$$\Pr[z_1r_1 + z_2r_2 + \cdots + z_nr_n = 0] = \Pr[z_1 + z_2x + z_2x^2 + \cdots + z_nx^n] \leq \frac{n-1}{2n-1} \leq \frac{1}{2}.$$

There are some other issues that enter the picture here, namely the bit complexity now that  $x^{n-1}$  has  $O(n)$  bits. One easy fix for this is to perform all operations modulo some prime.

## 1.2. Lectures 3–4: Pairwise independence

### 1.2.1. Lecture 3

Let  $X_1, \dots, X_n$  be random variables such that for any distinct  $i, j$ ,  $X_i, X_j$  are independent:

$$\Pr[X_i = \alpha, X_j = \beta] = \Pr[X_i = \alpha] \Pr[X_j = \beta].$$

This is referred to as *pairwise independence*. Analogously, we can define *k-wise independence*, which requires that any subset of at most  $k$  random variables is independent.

**Example.** Let random variables  $X_1, X_2$  take values in  $\{0, 1\}$  uniformly, and let  $X_3 = X_1 \oplus X_2$ . This set of random variables is pairwise independent, but not completely independent!

Given a cut  $(S, \bar{S})$  of a graph, denote

$$\partial S = \{(u, v) : u \in S, v \notin S\}.$$

Consider an algorithm that chooses a uniformly random cut  $S$  of the vertex set  $V$  (which corresponds to independently choosing each vertex with probability  $1/2$ ). Then,

$$\mathbb{E}[|\partial S|] = \sum_{e \in E} \Pr[e \in \partial S] = \sum_{\{u, v\} \in E} \Pr[u \in S, v \notin S] + \Pr[u \notin S, v \in S] = |E|/2.$$

In particular, this gives (in expectation) a  $1/2$ -approximation of a max-cut.<sup>1</sup>

Now, note that this algorithm does not require independence of all the  $|V|$  vertex-choosings, it suffices to have pairwise independence! This begs the question, how do we generate  $n$  pairwise independent while using a small number of actual random bits?

Bouncing off the idea in the previous example, we can take  $k$  random bits  $X_1, \dots, X_k$ , and generate  $2^k - 1$  pairwise independent random bits by considering  $\bigoplus_{i \in S} X_i$  for each non-empty  $S \subseteq [k]$  (why are these pairwise independent?).

Consequently, we can generate  $n$  pairwise random bits using just  $O(\log(n))$  random bits.

<sup>1</sup>Using Markov's inequality, it gives a  $1/2$ -approximation with probability at least  $1/2$ .

*Remark.* Since we have just  $\log n$  random bits, we can cycle through all the possible choices for the bits, since there are only  $n$  choices! This gives a deterministic polynomial time  $1/2$ -approximation algorithm for the max-cut problem. Instead of looking at all the  $O(2^n)$  cuts, it is enough to look at  $O(n)$  cuts. Interestingly, this does not even look at the structure of the graph!

**Proposition 1.1.** To generate  $n$  pairwise independent random bits, we require  $\Omega(\log n)$  independent random bits.

*Proof.* Suppose that given  $k$  independent random bits  $Y_1, \dots, Y_k$ , we can come up with  $n$  pairwise independent random bits  $X_1, \dots, X_n$ . Let  $f_i : \{0, 1\}^k \rightarrow \{0, 1\}$  for  $1 \leq i \leq n$  be defined by  $X_i = f_i(Y_1, \dots, Y_k)$ . Also, denote  $f_i^{-1}(1) = \{x \in \{0, 1\}^k : f_i(x) = 1\}$ .

The basic constraint that  $\Pr[X_i = 1] = 1/2$  means that  $|f_i^{-1}(1)| = 2^{k-1}$  and the pairwise independence constraint gives that for distinct  $i, j$ ,  $|f_i^{-1}(1) \cap f_j^{-1}(1)| = 2^{k-2}$ . Let  $M$  be the  $n \times 2^k$  matrix such that  $M_{ij} = f_i(j)$  (in the sense of the binary expansion of  $j$ ).

The previous constraints then just say that  $MM^\top = 2^{k-2}(I + J)$ , where  $J$  is the all ones matrix.

Note that the  $n \times n$  matrix  $2^{k-2}(I + J)$  is of rank  $n$ . It follows that  $\text{rank}(M) = \text{rank}(MM^\top) = n$ , so  $2^k \geq n$  and we are done! ■

Alternatively, after getting  $M$ , one may observe that if we replace 0 with  $-1$ , then the rows of  $M$  are orthogonal, which again gives the required.

Now, what happens if we want to generate pairwise independent functions instead of just bits? Can we do better? In particular, can we generate pairwise independent random variables  $X_1, \dots, X_n$  that uniformly take values in  $\mathbb{F}_q$ , where  $q$  is a prime power?

One simple construction is similar to the earlier one – take  $k := \log n$  random values  $y_1, \dots, y_k$  from  $\mathbb{F}_p$ , and consider  $\sum_{i \in S} y_i$  for each non-empty  $S \subseteq [k]$ . This takes  $\log n \cdot \log |\mathbb{F}|$  random bits.

A better construction for  $n = q$  is as follows – randomly choose  $a_0, a_1 \in \mathbb{F}$ , and let the required random variables be  $\{a_1 z + a_0 : z \in \mathbb{F}_q\}$ . This takes just  $\log n + \log |\mathbb{F}|$  bits! We leave the details of checking this to the reader.

### 1.2.2. Lecture 4

In the above construction for generating  $q$  pairwise independent random variables uniform in  $\mathbb{F}_q$ , if we set  $q = 2^r$ , then this in fact generates  $q$  pairwise independent random bits  $\log q$  times, using only  $2 \log q$  independent random bits!

The naïve method to do this would involve generating  $q$  pairwise independent random bits  $\log q$  times, which takes  $(\log q)^2$  bits.

Further, we can generalize the construction to  $n$  of the form  $q^r$  by considering  $\{a_0 + \sum_{i=1}^r a_i x_i : x_i \in \mathbb{F}_q\}$ , where the  $a_i$  are iid drawn from  $\mathbb{F}_q$ .

This idea can further be generalized to  $k$ -wise independence as well, taking a degree- $(k-1)$  polynomial  $\{\sum_{i=0}^{k-1} a_i x^i : x \in \mathbb{F}_q\}$  instead. Why are these  $k$ -wise independent? Fix distinct  $x_1, x_2, \dots, x_k \in \mathbb{F}_q$  and  $\alpha_1, \dots, \alpha_k \in \mathbb{F}$ . Is it true that

$$\Pr \left[ \sum_j a_j x_i^j = \alpha_i \text{ for all } i \right] = \frac{1}{q^k}?$$

Indeed, there is a unique solution  $(a_0, \dots, a_{k-1})$  to this since the matrix corresponding to the system of equations is a Vandermonde matrix, which has nonzero determinant (even over  $\mathbb{F}_q$ ).

**Exercise 1.1.** Show that a Vandermonde matrix is invertible.

**Solution**

Suppose instead that there is a nonzero vector  $v$  such that  $Mv = 0$ , where  $M$  is our  $k \times k$  Vandermonde matrix of interest. This gives a nonzero polynomial of degree at most  $k - 1$  with  $k$  roots, which is not possible.

## 1.3. Lectures 4–5: Counting distinct elements in a stream

## 1.3.1. Lecture 4 (continued)

Pseudorandomness has various applications in streaming algorithms. We generally have storage space that is far smaller than the input. We also have only one “pass” at the input and cannot look at older input. We can however run multiple copies of the same algorithm as we get the input, and in this case this can give better results.

**Problem.** Suppose we are getting a stream of items  $a_1, \dots, a_m$  in  $[n]$ . Count the number of distinct elements that appear.

A realistic example of the above is trying to find the number of unique visitors to a website.

One trivial way to do this is to store an array of size  $n$  of all the elements seen so far (or perhaps marking the elements which have been seen). This requires  $O(n)$  space.

Can we go to  $O(\log n)$  space, perhaps slightly giving up precision?

Let  $h$  be a function that maps each element in  $[n]$  to  $[0, 1]$  (the continuous interval) uniformly randomly. That is, each  $h(i)$  is independently uniformly randomly distributed in  $[0, 1]$ . We start with a variable  $m$  set at  $\infty$ . For a new  $a$  in the stream, we set  $m \leftarrow \min(m, h(a))$ . Finally, output  $1/m - 1$ .

The random variable  $m$  is essentially the minimum of  $k$  random variables iid drawn from  $[0, 1]$ , where  $k$  is the number of unique elements. Then,  $\mathbb{E}[m] = 1/(k + 1)$ .

## 1.3.2. Lecture 5

Before moving on, let us verify that  $\mathbb{E}[m] = 1/(k + 1)$ ? We have that for  $x \in [0, 1]$ ,

$$\Pr[m \geq x] = \Pr[h(i) \geq x \text{ for all } i] = (1 - x)^k.$$

Therefore,

$$\mathbb{E}[m] = \int_0^1 x \cdot k(1 - x)^{k-1} dx = \int_0^1 n(x^{k-1} - x^k) dx = \frac{1}{k + 1}.$$

Now, we still have to store all  $n$  outputs of  $h$ , so this has not really introduced any lower storage space. Choose a field  $\mathbb{F}$  with  $|\mathbb{F}| = N \geq n$ . We shall choose  $h(i)$  from  $\mathbb{F}$  (or rather,  $[N]$ ) such that they are pairwise independent. Recall that we had seen how to do this in Lectures 3 and 4. This construction only requires us to store the  $a$  and  $b$  from the algorithm, and we can compute  $h(i) = ai + b$  whenever needed. This also lowers the space requirement to  $O(\log n)$ . We shall now output  $(N/m) - 1$  instead of  $(1/m) - 1$ .

We want to show that  $(N/m) - 1$  is “close” to  $k$  with high probability. That is, let us try to bound

$$\Pr \left[ (1 - \epsilon) \frac{N}{k} \leq m \leq (1 + \epsilon) \frac{N}{k} \right]$$

from below.

Define

$$Y_{i,\lambda} = \mathbb{1}_{h(i) > \lambda} = \begin{cases} 1, & \text{if } h(i) > \lambda \\ 0, & \text{otherwise.} \end{cases}$$

Also define

$$Y_\lambda = \sum_{i \in S} Y_{i,\lambda},$$

where  $S$  is the set of the  $k$  distinct elements that are seen. Then, we want to find

$$\Pr \left[ Y_{(1-\epsilon)\frac{N}{k+1}} = 0 \text{ and } Y_{(1+\epsilon)\frac{N}{k+1}} \neq 0 \right].$$

Indeed,  $m$  is at least the lower bound iff no element in the stream is mapped to something less than it, and at most the upper bound iff at least one element is mapped to something less than it. Now,

$$\mathbb{E}[Y_\lambda] = \sum_{i \in S} \mathbb{E}[Y_{i,\lambda}] \approx k\lambda/N.$$

Then, using Markov's inequality,

$$\Pr[Y_\lambda \geq 1] \leq \mathbb{E}[Y_i] = \frac{k\lambda}{N}$$

and as a result,

$$\Pr[Y_{(1-\epsilon)\frac{N}{k}} = 0] = \Pr\left[m \geq (1-\epsilon)\frac{N}{k}\right] \geq \epsilon.$$

Observe that thus far, we have not used any sort of independence.

**Lemma 1.2.** If  $X_1, \dots, X_n$  are pairwise independent real-valued random variables,

$$\text{Var}\left[\sum_i X_i\right] = \sum_i \text{Var}[X_i].$$

*Proof.* We have

$$\begin{aligned} \text{Var}\left[\sum_i X_i\right] &= \mathbb{E}\left[\left(\sum_i X_i - \mathbb{E}[X_i]\right)^2\right] \\ &= \mathbb{E}\left[\sum_i (X_i - \mathbb{E}[X_i])^2 + 2\sum_{i < j} (X_i - \mathbb{E}[X_i])(X_j - \mathbb{E}[X_j])\right] \\ &= \sum_i \mathbb{E}\left[(X_i - \mathbb{E}[X_i])^2\right] + 2\sum_{i < j} \mathbb{E}\left[(X_i - \mathbb{E}[X_i])(X_j - \mathbb{E}[X_j])\right] \\ &= \sum_i \text{Var}[X_i] + 2\sum_{i < j} \mathbb{E}[X_i - \mathbb{E}[X_i]]\mathbb{E}[X_j - \mathbb{E}[X_j]]. \end{aligned} \quad (X_i, X_j \text{ are independent})$$

■

Now, set  $U = (1 + \epsilon)N/k$ , so we have  $\mathbb{E}[Y_U] = 1 + \epsilon$ . By the above lemma,

$$\text{Var}[Y_U] = k \text{Var}[Y_{i,U}] = k \cdot \frac{U}{N} \left(1 - \frac{U}{N}\right) = (1 + \epsilon) \left(1 - \frac{1 + \epsilon}{k}\right).$$

Therefore, using Chebyshev's inequality,

$$\begin{aligned} \Pr[Y_U \neq 0] &\geq 1 - \Pr[|Y_U - (1 + \epsilon)| \geq (1 + \epsilon)] \\ &\geq 1 - \frac{(1 + \epsilon) \left(1 - \frac{1 + \epsilon}{k}\right)}{(1 + \epsilon)^2} \geq 1 - \frac{1}{1 + \epsilon} = \frac{\epsilon}{1 + \epsilon}. \end{aligned}$$

Finally,

$$\begin{aligned} \Pr\left[Y_{(1-\epsilon)\frac{N}{k}} = 0 \text{ and } Y_{(1+\epsilon)\frac{N}{k}} \neq 0\right] &\geq 1 - \left(\Pr\left[Y_{(1-\epsilon)\frac{N}{k}} \neq 0\right] + \Pr\left[Y_{(1+\epsilon)\frac{N}{k}} \neq 0\right]\right) \\ &\geq \epsilon + \frac{\epsilon}{1 + \epsilon} - 1. \end{aligned}$$

## §2. Expander graphs and applications

### 2.1. Lectures 6–7: Magical graphs and two applications

#### 2.1.1. Lecture 6

Expander graphs are interesting because they are “pseudorandom” – they behave like random objects.

We recall the subject of error correcting codes, pioneered by Shannon in 1948. It studies the idea of introducing “redundancy” when transmitting messages so that the messages are understandable even in the presence of errors.

**Definition 2.1.** A code  $\mathcal{C}$  is a subset of  $\{0, 1\}^n$ . The elements of a code are called *codewords*.

**Definition 2.2.** Given  $x, y \in \{0, 1\}^n$ , the *Hamming distance*  $d_H(x, y)$  between  $x$  and  $y$  is  $|\{i \in [n] : x_i \neq y_i\}|$  and the *relative distance*  $\Delta(x, y)$  between  $x$  and  $y$  is  $d_H(x, y)/n$ . The distance  $d_H(\mathcal{C})$  of a code  $\mathcal{C}$  is  $\min_{x \neq y, x, y \in \mathcal{C}} d_H(x, y)$ .

The idea of this is that given a word in  $\{0, 1\}^k$ , we translate it bijectively into a codeword in  $\{0, 1\}^n$  and transmit it. Upon receiving the message, we decode the received word in some way to get a word.

One simple way is to decode a received word as the codeword closest to it, in the sense of the Hamming distance. This scheme allows the correction of errors if the received word is at Hamming distance less than  $(1/2)d_H(\mathcal{C})$  from the transmitted word.

**Definition 2.3.** The *rate* of a code is defined by

$$\text{Rate}(\mathcal{C}) = \frac{\log |\mathcal{C}|}{n}.$$

We also define the *relative distance*

$$\delta(\mathcal{C}) = \frac{d_H(\mathcal{C})}{n}.$$

One question that should immediately come to mind is: given a relative distance, what is the minimum rate required to achieve it? In less formal terms, what is the minimum amount of redundancy needed? We state it more formally.

**Problem.** Given constants  $\delta_0, r_0 \in (0, 1)$ , when can we construct codes  $\{\mathcal{C}_n\}_{n \in \mathbb{N}}$  such that  $\delta(\mathcal{C}_n) \rightarrow \delta_0$  and  $\text{Rate}(\mathcal{C}_n) \rightarrow r_0$ ?

This also presents another follow-up question: if codes of the above form exist, do there exist efficient encoding and decoding algorithms for the code? We do not look at this. Consider another question.

**Problem.** Suppose we have an algorithm  $\mathcal{A}$  with “one-sided error”. This means that if  $x$  is in the language  $L$  of interest,  $\mathcal{A}(x)$  is yes with probability 1, but if  $x$  is not in the language  $L$ ,  $\mathcal{A}(x)$  is no with probability  $\frac{15}{16}$ . How would one go about making the error probability very small, without using too many random bits?

One simple idea which we have discussed is to repeat the experiment a large number of times and output no if we get a no at any point. Indeed, if we repeat it  $\ell$  times, the error probability goes down to  $\leq (1/16)^\ell$ .

However, the fault with this is that if the algorithm uses  $k$  independent random bits (say), then repeating it  $\ell$  times requires  $\ell k$  independent random bits! Could we make it  $\ell + k$ ? It turns out that this is possible.

The two questions we have described seem incredibly different, but the answers to both are yes, with the ideas behind both involving “expander graphs”. Before getting to this, we define something else.

**Definition 2.4** (Magical graphs). A bipartite graph  $G = (L \sqcup R, E)$  is said to be  $(n, m, d)$ -magical,  $m \geq (3n/4)$ , if

1.  $|L| = n, |R| = m,$



2. for any  $v \in L$ ,  $\deg(v) = d$ , and
3. for every subset  $S \subseteq L$  with  $|S| \leq n/10d$ ,  $|\Gamma(S)| \geq (5d/8)|S|$ .

Above,  $\Gamma(S)$  denotes the neighbourhood of  $S$ .

Typically,  $n$  and  $m$  are of similar orders and  $d$  is a constant. This says that any “small” subset expands a lot – the neighbours of the vertices in the subset do not coincide too much. Ideally, with no intersection between neighbourhoods, we would have  $|\Gamma(S)| = d|S|$ , and we are demanding about half of this.

First, we shall see why magical graphs exist. Following this, we connect them to the questions we looked at earlier.

**Theorem 2.1.** For  $d \geq 24$  and sufficiently large  $n$ ,  $(n, m, d)$ -magical graphs exist.

*Proof.* For each vertex in  $L$ , choose its  $d$  neighbours randomly. Let  $S \subseteq L$  with  $|S| = s \leq n/10d$  and  $T = R$  with  $|T| = (5d/8)s$ ,

$$\Pr [\Gamma(S) \subseteq T] \leq \left( \frac{|T|}{m} \right)^{ds} \leq \left( \frac{5ds}{8m} \right)^{ds}.$$

This is for a *fixed*  $S, T$  however. Using the union bound,

$$\begin{aligned} \Pr [\exists S, T \text{ as above such that } \Gamma(S) \subseteq T] &\leq \sum_{S, T} \left( \frac{5ds}{8m} \right)^{ds} \\ &\leq \sum_{s=1}^{n/10d} \binom{n}{s} \binom{m}{5ds/8} \left( \frac{5ds}{8m} \right)^{ds} \\ &\leq \sum_{s=1}^{n/10d} \left( \frac{ne}{s} \right)^s \left( \frac{8me}{5ds} \right)^{5ds/8} \left( \frac{5ds}{8m} \right)^{ds} \quad \left( \binom{n}{k} \geq (ne/k)^k \right) \\ &= \sum_{s=1}^{n/10d} \left( \frac{ne}{s} \right)^s e^{5ds/8} \left( \frac{5ds}{8m} \right)^{3ds/8} \\ &\leq \sum_{s=1}^{n/10d} \left( \frac{ne}{s} \right)^s e^{5ds/8} \left( \frac{5ds}{6n} \right)^{3ds/8} \quad (m \geq 3n/4) \\ &= \sum_{s=1}^{n/10d} \left( \frac{s}{n} \right)^{s(3d/8-1)} e^{s(5d/8+1)} (5d/6)^{3ds/8} \\ &\leq \sum_{s=1}^{n/10d} (10d)^{-s(3d/8-1)} e^{s(5d/8+1)} (5d/6)^{3ds/8} \quad (s/n \leq 1/10d) \\ &\leq \sum_{s=1}^{\infty} (10d)^{-s(3d/8-1)} e^{s(5d/8+1)} (5d/6)^{3ds/8} \\ &= \frac{\alpha}{1-\alpha}, \end{aligned}$$

where  $\alpha = (10d)^{1-(3d/8)} e^{(5d/8)+1} (5d/6)^{3d/8}$ . The above is less than 1 when  $\alpha < 1/2$ . To check for what values of  $d$

this is true,

$$\begin{aligned}\log \alpha &= \left(1 - \frac{3d}{8}\right) (\log 10 + \log d) + 1 + \frac{5d}{8} + \frac{3d}{8} (\log(5/6) + \log d) \\ &= \log d + d \left(\frac{5}{8} - \frac{3}{8} \log(10) + \frac{3}{8} \log(5/6)\right) + (1 + \log 10) \\ \frac{d \log \alpha}{dd} &\approx \frac{1}{d} - 0.306,\end{aligned}$$

which is negative for  $1/d < 0.306$  (or equivalently,  $d \geq 5$ ). Since  $\alpha$  is decreasing in  $d$  for  $d \geq 24$ , it suffices to check that  $\alpha < 1/2$  when  $d = 24$ . Indeed, it is easily verified that  $\alpha \approx 0.413 < 1/2$  in this case, completing the proof. ■

Now, let us look at reduction of randomness using magical graphs.

Let  $\mathcal{A}$  be an algorithm that uses  $k$  random bits with error probability  $< 1/16$ . Take  $n = 2^k$ , and let  $G = (L \sqcup R, E)$  be a  $(n, n, d)$ -magical graph. Choose a random vertex  $v \in L$ , and take all  $d$  neighbours  $u_1, \dots, u_d$  of  $v$ . Each  $u_i$  can be thought of as a  $k$  bit string. For  $i = 1, \dots, d$ , run  $\mathcal{A}$  with  $u_i$  as the choice of random bits. Observe that we are only using  $k$  random bits here, namely in the choice of  $v$ .

Why does the error probability go down?

Let  $B \subseteq \{0, 1\}^k$  be the set of “bad” inputs for algorithm  $\mathcal{A}$ . We know that  $|B| \leq n/16$ . What is the probability of failure when we run it  $d$  times as described above? The algorithm fails iff every  $u_i$  is in  $B$ .

We claim that there are less than  $n/10d$  such vertices  $v$  with every neighbour in  $B$ . Suppose instead that there are is a set  $S$  with  $n/10d$  vertices with all neighbours in  $B$ . Then,

$$\frac{n}{16} > |B| \geq \Gamma(S) \geq \frac{5d}{8} |S| \geq \frac{n}{16},$$

which is a contradiction.

Therefore, the probability of failure is at most  $1/10d$ .

We can make  $d$  very large (up to a limit forced by  $n = 2^k$ ), so the probability of failure can be made very small. Using the same number of random bits, we have managed to significantly decrease the error probability.

The issue now however is that the above scheme requires the construction of exponentially large magical graphs. We require a very efficient algorithm (polynomial in  $\log n$ ) to sample the neighbours of a random vertex in a magical graph.

### 2.1.2. Lecture 7

**Lemma 2.2.** Given a  $(n, m, d)$ -magical graph, for any  $S \subseteq L$  of size at most  $n/10d$ , there exists  $v \in \Gamma(S)$  such that  $v$  has a unique neighbour in  $S$ .

*Proof.* Suppose instead that no such  $v$  exists. Then,

$$d|S| = |e(\Gamma(S), S)| \geq 2|\Gamma(S)| \geq 2 \cdot \frac{5d}{8} |S|,$$

a contradiction. ■

Consider some  $(n, m = 3n/4, d)$ -magical graph, and let  $M$  be the  $m \times n$  adjacency matrix of the graph, where  $M_{ij}$  is 1 iff there is an edge between the  $i$ th vertex on the right and the  $j$ th vertex on the left, and 0 otherwise.

**Definition 2.5.** A code  $\mathcal{C} \subseteq \{0, 1\}^n$  is said to be a *linear code* if it is a subspace when viewed as a subset of the vector space  $\mathbb{F}_2^n$ .

This is equivalent to saying that if  $x, y \in \mathcal{C}$ , then  $x + y \in \mathcal{C}$ . Observe that the distance of a linear code is equal to the minimum weight of its codeword.

Consider the code defined by

$$\mathcal{C} = \{x : Mx = 0\},$$

the null space of  $M$  (over  $\mathbb{F}_2$ ). In coding theory lingo, one would say that  $M$  is the parity check matrix of  $\mathcal{C}$ . Evidently,

$$|\mathcal{C}| = 2^{n - \text{rank}(M)} \text{ and } \text{Rate}(\mathcal{C}) = \frac{n - \text{rank}(M)}{n} \geq \frac{n - m}{n} \geq \frac{1}{4}.$$

We claim that  $d_H(\mathcal{C}) > n/10d$ , so the relative distance is at least  $1/10d$ . Suppose instead that there is some  $x \in \mathcal{C}$  such that  $\text{wt}(x) \leq n/10d$ . Let  $S \subseteq [n]$  be the subset of  $L$  such that  $v \in S$  iff  $x_v \neq 0$ . By Lemma 2.2, there exists some  $u \in R$  such that  $M_{uv} = 1$  for precisely one  $v \in S$ . However, this implies that  $(Mx)_u = \sum_v M_{uv}x_v = 1$ .

## 2.2. Lectures 7, 8, 10: Testing connectivity of undirected graphs

### 2.2.1. Lecture 7 (continued)

**Problem.** Given a graph  $G$  and two vertices  $s, t \in V(G)$ , determine if there is a path between  $s$  and  $t$ .

To test connectivity of a graph, we can just test the above by iterating through all  $t \in V(G)$ . We work in the setting where running time is not an issue (as long as it is polynomial), but we have space constraints.

One way to do this, of course, is through a depth-first/breadth-first search. This requires  $\Omega(n)$  space however, which is more than we can afford.

Recall that there is a path between  $s, t$  iff  $((I + A)^n)_{st} \neq 0$ . Indeed,  $(A^i)_{st}$  gives the number of length  $i$  walks from  $s$  to  $t$ , and  $(I + A)^n$  is just some weighted sum of the  $A^i$ . Can we compute  $(I + A)^n$  in  $O(\log^2 n)$  space, say? We remark that here, space means that the size of the input and output tapes are “large”, but we cannot alter the input tape and once we write something to the output tape, we cannot change it; we have an intermediate tape of “small” size which is what we use for computation.

The answer is yes, but we do not say why.

Consider the following algorithm, that is also  $O(\log^2 n)$  space.

---

#### Algorithm 1: Checking connectivity of two vertices in an undirected graph

---

**Input:** An graph  $G$ , and vertices  $s, t \in V(G)$

**Output:** Connectivity of  $s, t$

```

1 isPath( $s, t, k = 2^\ell$ )                                     // Outputs whether there is a path between  $s$  and  $t$  of length at most  $k$ 
2   if  $\ell = 0$  then
3     return yes iff  $s, t$  are adjacent
4   foreach  $v \in V(G)$  do
5     return yes iff isPath( $s, v, 2^{\ell-1}$ ) and isPath( $s, v, 2^{\ell-1}$ )
6 return isPath( $s, t, n$ )

```

---

Observe that the depth of this recursion tree is  $O(\log n)$ . In each recursion call, we use  $O(\log n)$  space. As a result, we use  $O(\log^2 n)$  space in all.

All the algorithms we have presented thus far work in the setting of directed graphs as well.

Can we check connectivity in  $O(\log n)$  space? We present a randomized algorithm due to Reingold [Rei08] that does so. The algorithm is as follows.

**Algorithm 2:** Checking connectivity of two vertices in an undirected graph**Input:** An graph  $G$ , and vertices  $s, t \in V(G)$ **Output:** Connectivity of  $s, t$ 

```

1  $v \leftarrow s$ 
2 Uniformly randomly choose a neighbour  $u$  of  $v$ 
3 if  $u = t$  then
4   return yes
5 else
6    $v \leftarrow u$ 
7   go to line 2

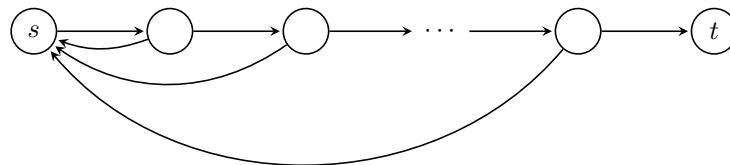
```

Suppose that  $G$  is connected. For the sake of simplicity, suppose that the graph is  $d$ -regular. If it is not, add self-loops to make it so. Also, after doing this, add another self-loop at each vertex.

We claim that

$$\Pr \left[ t \text{ is seen in } O(n^3 \log n) \text{ steps} \right] \geq \frac{1}{2}.$$

This algorithm does *not* work for directed graphs. Indeed, consider the graph



where it takes exponential time for the probability of seeing  $t$  to go over  $1/2$ .

The transition matrix of the random walk is defined by

$$M_{ij} = \frac{1}{d} e(i, j),$$

where  $e(i, j)$  is the number of edges between  $i$  and  $j$ .

Given the initial probability vector  $x^{(0)} = \mathbb{1}_s$  ( $x_s^{(0)} = 1$  and  $x_v^{(0)} = 0$  for  $v \neq s$ ), the probability distribution of vertices after  $t$  steps of the random walk is given by  $x^{(t)} = M^t x^{(0)}$ .

Consider the uniform probability vector  $u$ , where  $u_v = 1/n$  for all  $v$ , and observe that  $Mu = u$ .  $u$  is called a stationary distribution of the random walk.

We claim that  $u$  is the only stationary distribution of the walk (this assumes that  $G$  is connected – why?).

## 2.2.2. Lecture 8

Denote by  $p^{(t)}$  the probability vector at time  $t$ , and  $p_i^{(t)}$  the probability that we are at vertex  $i$  at times  $t$ . By definition,

$$p_i^{(t)} = \frac{1}{d} \sum_{j \leftrightarrow i} p_j^{(t-1)}.$$

Now, we wish to analyze  $M^t$  to show that  $p^{(t)}$  is close to the stationary distribution for somewhat large  $t$ .

Because  $M$  is symmetric, we can write  $M = UDU^\top$ , where  $U$  is an orthogonal matrix of eigenvectors and  $D$  is a diagonal matrix of the real eigenvalues. As a result,

$$p^{(t+1)} = (UDU^\top)^t p^{(0)} = UD^t U^\top p^{(0)}.$$

Therefore, what matters is  $D^t$ . Let the eigenvalues of  $M$  be  $\lambda_1 \geq \lambda_2 \geq \dots \geq \lambda_n$ .

We trivially have that  $\lambda_1 = 1$  is an eigenvalue of  $M$  with eigenvector being the stationary distribution  $u_1$  that we saw last lecture. Observe that 1 is the largest eigenvalue (in absolute value) of  $M$  because for any vector  $x$ ,  $\lambda \max_i x_i =$

$$\max_i (Mx)_i \leq \max_i x_i.$$

If  $|\lambda_j| < 1$  for  $j > 1$ , then  $\lim_{t \rightarrow \infty} p^{(t)} = u_1$  (all the other entries of  $D^t$  decay to 0).

To show fast convergence, we would like to show that the other eigenvalues are bounded away from 1.

**Definition 2.6.** Given a graph  $G$ ,  $\lambda(G) = \max\{|\lambda_2(G)|, |\lambda_n(G)|\}$ . Equivalently,

$$\lambda(G) = \max_{x: \langle x, u_1 \rangle = 0} \frac{\|Mx\|}{\|x\|}.$$

If  $G$  is obvious, we denote this as merely  $\lambda$ .

$(1 - \lambda)$  is referred to as the *spectral gap* of the graph. We would like to show that the spectral gap of the graph is “large”.

We have

$$\|p^{(t)} - u_1\| = \|M^t(p^{(0)} - u_1)\|$$

Observe that because  $p^{(0)}$  is a probability distribution,  $\langle p^{(0)}, u_1 \rangle = 1/n = \langle u_1, u_1 \rangle$ , so  $p^{(0)} - u_1$  is orthogonal to  $u_1$ . As a result,

$$\|p^{(t)} - u_1\| \leq \lambda^t \|p^{(0)} - u_1\|.$$

Suppose we want  $\|p^{(t)} - u_1\| \leq 1/n^2$ . For our specific choice of  $p^{(0)}$  concentrated on a single point, it is easily verified that  $\|p^{(0)} - u_1\| \leq 1$ . Therefore, for  $t > \log_\lambda \epsilon$ ,  $\|p^{(t)} - u_1\| \leq 1/n^2$ .

We have that the convergence time is

$$\log_\lambda \epsilon = O\left(\frac{1}{1 - \lambda} \log n\right).$$

We would like to bound the spectral gap from below, perhaps by an inverse polynomial.

It is true that

$$\lambda_2 = \max_{\substack{x \perp u_1 \\ \|x\|=1}} x^\top Mx$$

Indeed, if  $x = \alpha_2 u_2 + \dots + \alpha_n u_n$  with  $\sum \alpha_i^2 = 1$ ,

$$x^\top Mx = \lambda_2 \alpha_2^2 + \dots + \lambda_n \alpha_n^2 \leq \lambda_2.$$

We shall bound  $\lambda_2$  for every graph, and use it for  $H = G^2$ , the multigraph with an edge from  $i$  to  $j$  if there was a length 2 walk from  $i$  to  $j$ . The random walk matrix on  $H$  is just  $M^2$ , so  $\lambda_2(H) = \lambda(G)^2$ . Now,

$$\begin{aligned} x^\top Mx &= \sum_i x_i \sum_{j \in \Gamma(i)} \frac{1}{d} x_j \\ &= \sum_{(i,i) \in E} \frac{x_i^2}{d} + \sum_{\substack{(i,j) \in E \\ i \neq j}} \frac{2x_i x_j}{d} \\ &= \sum_{(i,i) \in E} \frac{x_i^2}{d} + \sum_{\substack{(i,j) \in E \\ i \neq j}} \frac{x_i^2 + x_j^2 - (x_i - x_j)^2}{d} \\ &= \underbrace{\sum_{(i,j) \in E} \frac{x_i^2}{d}}_1 + \sum_{\substack{(i,j) \in E \\ i \neq j}} \frac{(x_i - x_j)^2}{d}. \end{aligned}$$

We want to bound the second expression from below. Because  $\langle x, u_1 \rangle = 0$ , there exist coordinates of both positive and negative sign. Further, by the pigeonhole principle, there exists some coordinate  $i_1$  such that  $x_{i_1}$  is of absolute value at least  $1/\sqrt{n}$ . Assume that the sign of this coordinate is positive. Let  $i_k$  be a coordinate of negative sign, and  $i_1 i_2 \cdots i_k$  a path from  $i_1$  to  $i_k$ . Then, using the Cauchy-Schwarz inequality,

$$\begin{aligned} \sum_{\substack{(i,j) \in E \\ i \neq j}} (x_i - x_j)^2 &= \sum_{1 \leq j \leq (n-1)} (x_{i_{j+1}} - x_{i_j})^2 \\ &\geq \frac{1}{n} \left( \sum_{1 \leq j \leq (n-1)} |x_{i_{j+1}} - x_{i_j}| \right)^2 \\ &\geq \frac{1}{n} \left( \sum_{1 \leq j \leq (n-1)} x_{i_{j+1}} - x_{i_j} \right)^2 \geq \frac{1}{n^2}. \end{aligned}$$

Therefore, the largest eigenvalue of any graph is at most  $1 - 1/nd^2$ .

### 2.2.3. Lecture 10

By arguments discussed in Lecture 8, we know that the random walk on an expander graph mixes in  $O(\log n)$  time. This leads to a *deterministic* polynomial time algorithm to determine  $s, t$  connectivity in log-space – merely iterate over all paths of length  $\log n$ . That is, we iterate over all elements of  $[d]^{\log n}$  and for each such element, we follow the corresponding path and check if we see  $t$  anywhere.

How do we extend this to arbitrary graphs? Given a graph  $G$ , we would like to get a graph  $G'$  such that

- $G'$  has polynomially many vertices,
- $G'$  has “constant” degree (independent of  $n$ ),
- $G'$  preserves connectivity ( $s, t$  are connected in  $G$  iff some  $s', t'$  are connected in  $G'$ ), and
- each component of  $G'$  is an expander.

Let us look at a couple of operations one can do on a  $(n, d, \lambda)$ -expander  $G$ .

**Squaring ( $G^2$ )** the vertex set is the same as that of  $G$ , and there is an edge between  $u, v$  for each length 2 walk between  $u, v \in G$ . This is a multigraph with self-loops.  $G^2$  is a  $(n, d^2, \lambda^2)$ -expander; the random walk matrix on  $G^2$  is just  $M^2$ , where  $M$  is the random walk matrix on  $G$ . Because the second eigenvalue has gone down, connectivity has improved. However, the degree has increased.

**Zig-zag product ( $G \circledast H$ )** Let  $G$  be a  $(n, D, \lambda_1)$ -expander and  $H$  a  $(D, d, \lambda_2)$ -expander. Note that the degree of the first graph is the number of vertices in the second! The goal of this product is to decrease the degree, without changing expansion too much.

The idea is very similar to that we saw in Lecture 9 to derandomize algorithms, and we shall “derandomize the random walk in  $G$  using  $H$ ” – instead of going to a random neighbour in  $G$ , we determine which vertex to travel to using  $H$ .

$G \circledast H$  has  $nD$  vertices and is  $d^2$ -regular. Suppose that the edges at each vertex in  $G$  are put in bijection with  $V(H)$  in some arbitrary manner.

- The vertex set of  $G \circledast H$  is  $V(G) \times V(H)$ ; we shall replace each vertex of  $G$  with the vertices of  $H$ .
- Let  $v = (a_G, a_H) \in V(G \circledast H)$  and  $(i, j) \in [d]^2$ . The  $(i, j)$ th neighbour  $(b_G, b_H)$  of  $(a_G, a_H)$  is as follows.
  1. Let  $a'_H$  be the  $i$ th neighbour of  $a_H$  in  $H$ .
  2. Let  $(b_G, b'_H)$  be the  $a'_H$ th neighbour of  $a_H$  in  $G$  (recall that we had put the neighbours of a given vertex in bijection with  $V(H)$ ). That is,  $b_G$  is the  $a'_H$ th neighbour of  $a_G$  and  $a_G$  is the  $b'_H$ th neighbour of  $b_G$ .

3. Let  $b_H$  be the  $j$ th neighbour of  $b'_H$ .

**Lemma 2.3.** Suppose that  $H$  is a non-bipartite graph, and let  $w_H, w'_H$  be any vertices in  $H$ . Then,  $s, t$  are connected in  $G$  iff  $(s, w_H)$  and  $(t, w'_H)$  are connected in  $G \mathbin{\textcircled{Z}} H$ .

*Proof.* It suffices to show that for any  $v \in V(G)$ , the “cloud”  $v \times V(H)$  is connected. Indeed, the connectivity of clouds is identical to the connectivity of  $G$ . Consider some arbitrary  $a_H, b_H$  with a 2-walk  $a_H v_H b_H$  in  $H$  between them. Due to the non-bipartiteness of  $H$ , we are done if we show that there is a path between  $(v, a_H)$  and  $(v, b_H)$ . Indeed, considering some neighbour  $(w, c_H)$  of  $(v, a_H)$  in  $G \mathbin{\textcircled{Z}} H$  such that  $w$  is the  $v_H$ th neighbour of  $v$ , it is seen that both  $(v, a_H)$  and  $(v, b_H)$  are adjacent to  $(w, c_H)$  in  $G \mathbin{\textcircled{Z}} H$ , so we are done. ■

What happens to the eigenvalue of  $G \mathbin{\textcircled{Z}} H$ ? For the sake of simplicity, let  $\gamma_1 = 1 - \lambda_1$  and  $\gamma_2 = 1 - \lambda_2$  be the corresponding spectral gaps. We claim that

**Lemma 2.4.** It is true that

$$\gamma(G \mathbin{\textcircled{Z}} H) \geq \gamma_1 \gamma_2^2.$$

We prove the above shortly, and first describe how this results in a conversion of our graph to an expander graph.

First of all, how do we construct the  $(D, d, \lambda_2)$  graph  $H$ ?

There exists a simple method to construct a  $(D^4, D, 1/8)$  graph  $H$  that we do not describe.

To generate a larger graph with bounded spectral value, we can do the following.

1. Set  $G_1$  as  $H^2$ .
2. For each  $k$ , set  $G_{k+1}$  as  $G_k^2 \mathbin{\textcircled{Z}} H$ .

We claim that  $G_k$  is a  $(D^{4k}, D^2, 1/2)$  graph (by  $1/2$  we mean that the eigenvalue is at most  $1/2$ ).

The first two parameters are easily verified. It may be shown using Lemma 2.4 without much trouble that  $\lambda(G \mathbin{\textcircled{Z}} H) \leq \lambda_1 + 2\lambda_2$ .

Then, an inductive argument yields the eigenvalue bound for  $G_k$ . Indeed,  $\lambda(G_1) \leq 1/2$  and  $\lambda(G_{k+1}) \leq \lambda(G_k)^2 + 2\lambda(H) \leq (1/2)^2 + 2/8 = 1/2$ . Before coming to the proof of Lemma 2.4, we describe the construction used to convert  $G$  to a graph  $G'$  satisfying the conditions described earlier.

---

**Algorithm 3:** Converting an arbitrary graph to an expander

---

**Input:** An graph  $G$

**Output:** A graph  $G'$  each of whose components is an expander

```

1  $G_0 \leftarrow \text{regularize}(G)$ 
2 for  $1 \leq k \leq O(\log n)$  do
3    $G_{k+1} \leftarrow G_k^2 \mathbin{\textcircled{Z}} H$ 
```

---

First, make  $G$  a regular graph.  $G_0$  is an  $(n, D^2, 1 - \frac{1}{\text{poly}(n)})$  graph. We claim that  $G_k$  is an  $(nD^{4k}, D^2, 17/18)$ .

The first two parameters are straightforward. Let  $\gamma_k = 1 - \lambda(G_k)$ . Assume that for some  $k$   $\gamma_{k-1} \leq 1/18$ . Then,

$$\begin{aligned}
 \gamma_{k+1} &\geq \left(1 - (1 - \gamma_k)^2\right) \left(\frac{7}{8}\right)^2 \\
 &= (2\gamma_k - \gamma_k^2) \frac{49}{64} \\
 &\geq \left(\frac{35}{18}\gamma_k\right) \frac{49}{64} \geq \frac{5}{4}\gamma_k.
 \end{aligned}$$

Consequently,  $\gamma$  increases from  $1/\text{poly}(n)$  to a constant in  $O(\log n)$  steps!

The only thing that remains is to prove Lemma 2.4.

**Lemma 2.5.** Let  $C$  be a random walk matrix with eigenvalue  $\lambda$ . Then, it is possible to write  $C = (1 - \lambda)J + \lambda E$  for some matrix  $E$  with spectral norm equal to 1.

*Proof.* Let  $v$  be a vector, and let  $v_1, v_2$  be its components along and orthogonal to the all ones vector. Observe that all eigenvalues of  $J$  other than the first are equal to 0. As a result,

$$\begin{aligned} \|Ev\| &= \frac{1}{\lambda} \|Cv - (1 - \lambda)Jv\| \\ &= \frac{1}{\lambda} \|v_1 + Cv_2 - (1 - \lambda)v_1\| \\ &= \frac{1}{\lambda} \|\lambda v_1 + Cv_2\| \\ &= \frac{1}{\lambda} \sqrt{\lambda^2 \|v_1\|^2 + \|Cv_2\|^2} \\ &\leq \frac{1}{\lambda} \sqrt{\lambda^2 \|v_1\|^2 + \lambda^2 \|v_2\|^2} = \|v\|. \end{aligned}$$

The above inequality is tight on setting  $v_2$  as the second eigenvector of  $C$ , so the spectral norm is precisely 1. ■

*Proof of Lemma 2.4.* Let  $B$  be the random walk matrix of  $H$ , and let  $\tilde{B} = I_n \otimes B$ .<sup>2</sup> Let  $\tilde{A}$  be the matrix that encodes the second edge (using  $G$ -edges) traversed in the construction of the zigzag product. Then, the transition matrix  $M$  of  $G \otimes H$  is just  $\tilde{B}\tilde{A}\tilde{B}$ !

Now, denote by  $J$  the matrix that has all elements  $1/n$ .

Now, use Lemma 2.5 on  $\tilde{B}$ .

$$\begin{aligned} M &= \tilde{B}\tilde{A}\tilde{B} \\ &= (\gamma_2 \tilde{J} + \lambda_2 \tilde{E}) \tilde{A} (\gamma_2 \tilde{J} + \lambda_2 \tilde{E}) \\ &= \gamma_2^2 \tilde{J}\tilde{A}\tilde{J} + \gamma_2(1 - \gamma_2) (\tilde{E}\tilde{A}\tilde{J} + \tilde{J}\tilde{A}\tilde{E}) + (1 - \gamma_2)^2 \tilde{E}\tilde{A}\tilde{E} \\ &= \gamma_2^2 \tilde{J}\tilde{A}\tilde{J} + (1 - \gamma_2^2)X, \end{aligned}$$

for some matrix  $X$ . Observe that the spectral norm of  $X$  is at most 1 because  $\tilde{E}, \tilde{A}, \tilde{J}$  all have spectral norms at most 1, and so by submultiplicativity of the spectral norm,<sup>3</sup>  $\tilde{E}\tilde{A}\tilde{J}, \tilde{J}\tilde{A}\tilde{E}$ , and  $\tilde{E}\tilde{A}\tilde{E}$  all have spectral norms at most 1.

Now, we wish to bound the second eigenvalue of  $M$ . Note that  $\tilde{J}\tilde{A}\tilde{J}$  is precisely the random walk matrix of  $G \otimes K_n$ , which has second eigenvalue equal to that of  $G$ . Recalling very carefully that the second eigenvector of  $M$  is some vector  $v$  orthogonal to  $\mathbf{1}$ , we have

$$\begin{aligned} \|Mv\|_2 &= \|\gamma_2^2 \tilde{J}\tilde{A}\tilde{J}x\| + \|(1 - \gamma_2^2)Xx\| \\ &\leq (\gamma_2^2 \lambda_1 + (1 - \gamma_2^2))\|v\| = (1 - \gamma_2^2 \gamma_1)\|v\|, \end{aligned}$$

completing the proof. ■

## 2.3. Lecture 9: Expander graphs

### 2.3.1. Lecture 9

**Definition 2.7.** A graph  $G$  is said to be an  $(n, d, \lambda)$ -expander if  $|V(G)| = n$ ,  $G$  is  $d$ -regular, and  $\lambda$  is the second largest eigenvalue  $\lambda(G)$  of  $G$  in absolute value.

<sup>2</sup>this is a  $nD \times nD$  matrix, with  $n$  blocks on the diagonal that are all  $B$ s.

<sup>3</sup> $\|AB\| \leq \|A\|\|B\|$ . This is obvious because  $\|ABx\| \leq \|A\|\|Bx\| \leq \|A\|\|B\|\|x\|$ .



**Definition 2.8** (Spectral expanders). A sequence  $\{G_n\}_{n \geq 0}$  of  $d$ -regular graphs is said to be a *spectral expander family* if for some  $\lambda < 1$ ,  $\lambda(G_i) \leq \lambda$  for all  $i$ .

We saw last lecture that random walks on expander graphs converge to the uniform distribution in  $O(\log n)$  steps. This means that there are only  $d^{O(\log n)} = \text{poly}(n)$  paths to explore. Therefore, there is a *deterministic* polynomial time algorithm to determine connectivity of expander graphs.

**Definition 2.9** (Sparsity). Given a  $d$ -regular graph  $G$ , define the *sparsity*

$$h(G) = \min_{\substack{S \subseteq V \\ |S| \leq (n/2)}} \frac{|E(S, \bar{S})|}{d|S|}.$$

Some definitions use  $|E(S, \bar{S})|/|S|$  instead.

It is natural to see that  $h(G)$  measures (in some sense) how well-connected a graph is. If it is low, there is some “bottleneck” in the graph where the random walk can get stuck – a set of high measure with very few outgoing edges.

It is clear that  $h(G) \leq 1$ .

**Definition 2.10** (Combinatorial expanders). A sequence  $\{G_n\}_{n \geq 0}$  of  $d$ -regular graphs is said to be a *combinatorial expander family* if for some  $h > 0$ ,  $h(G_i) \geq h$  for all  $i$ .

**Theorem 2.6** (Cheeger’s Inequality). For any graph  $G$  with second eigenvalue  $\lambda_2$  and sparsity  $h$ ,

$$\frac{1 - \lambda_2}{2} \leq h \leq \sqrt{2(1 - \lambda_2)}.$$

In particular, spectral expanders are combinatorial expanders and vice-versa.

We prove this later.

*Markov chain Monte Carlo* methods find many uses nowadays in problems such as sampling random spanning trees, random independent sets etc. The idea in these is that we start with an arbitrary spanning tree (say), and then randomly move to a “neighbouring” spanning tree – add a random edge not in the spanning tree and remove a random edge from the cycle thus formed. After sufficiently many steps, we are at a (n almost) uniformly random spanning tree. This massive graph composed of spanning trees as vertices ends up being an expander! Because the graph of spanning trees has only exponentially many vertices, we get a polynomial time algorithm to randomly sample spanning trees.

**Example.** The  $n$ -cycle  $C_n$  has sparsity  $h(C_n) = 2/n$ , and  $\lambda(C_n) = \cos(2\pi/n) \approx 1 - (2\pi/n)^2$ .

The hypercube graph  $H_n := P_2^{\otimes n}$ . We have  $h(H_n) = 1/n$  and  $\lambda(H_n) = 1 - \frac{1}{k}$ .

Each of these give a case where the appropriately inequality in **Cheeger’s Inequality** is (asymptotically) tight.

What guarantee do we even have that expanders exist? It turns out that a random  $d$ -regular graph is a (combinatorial) expander with high probability!

However, how do we construct expander graphs? Our goal is to use expander graphs to reduce randomness in algorithms, so it does not make sense to construct them using the above random argument. We also want the algorithm itself to run in polylog time – this requirement makes sense in light of our remarks towards the end of [Lecture 6](#).

**Example.** Let  $p$  be a prime and consider the 3-regular graph over  $\mathbb{F}_p^*$ , where each  $x$  is adjacent to  $x + 1, x - 1, x^{-1}$ . This graph is an expander, but the proof of this is not very straightforward.

**Theorem 2.7** (Expander Mixing Lemma). Let  $G$  be a  $(n, d, \lambda)$ -expander. Then, for any  $S, T \subseteq V$ ,

$$\left| E(S, T) - \frac{d}{n} |S| |T| \right| \leq d \lambda \sqrt{|S| |T|}.$$

If  $G$  were a random graph, then the expected number of edges between  $S, T$  is precisely  $(d/n)|S||T|$  – of the  $d|S|$  edges out of  $S$ , we expect a  $|T|/n$  fraction to be incident on  $T$ .

*Proof.* Let  $M$  be the transition matrix of the random walk on  $G$ ; it is equal to  $(1/d)$  times the adjacency matrix of  $G$ . For any set  $X$ , let  $\mathbb{1}_X$  be the indicator vector of  $X$  with 1s at vertices in  $X$  and 0 elsewhere.

Observe that

$$\frac{1}{d} E(S, T) = \mathbb{1}_S^\top M \mathbb{1}_T.$$

Now, we have  $M = \sum_i \lambda_i u_i u_i^\top$  using the spectral theorem, where  $(u_i)$  are orthonormal eigenvectors of  $M$  with corresponding real eigenvalues  $(\lambda_i)$ . Note in particular that  $\lambda_1 = 1$  and  $u_1$  is the vector with all coordinates having value  $1/\sqrt{n}$ .

Let  $\mathbb{1}_S = \sum_i \alpha_i u_i$  and  $\mathbb{1}_T = \sum_i \beta_i u_i$ . Note in particular that  $\alpha_1 = \langle \mathbb{1}_S, u_1 \rangle = |S|/\sqrt{n}$  and  $\beta_1 = |T|/\sqrt{n}$ .

Using orthonormality,

$$\begin{aligned} \frac{1}{d} E(S, T) &= \left( \sum_i \alpha_i u_i \right)^\top \left( \sum_i \lambda_i u_i u_i^\top \right) \left( \sum_i \beta_i u_i \right) \\ &= \sum_i \alpha_i \beta_i \lambda_i \\ &= \alpha_1 \beta_1 + \sum_{i=2}^n \alpha_i \beta_i \lambda_i \\ &= \frac{|S||T|}{n} + \sum_{i=2}^n \alpha_i \beta_i \lambda_i. \end{aligned}$$

Therefore,

$$\begin{aligned}
 \left| E(S, T) - \frac{d}{n} |S| |T| \right| &= \left| \sum_{i=2}^n \alpha_i \beta_i \lambda_i \right| \\
 &\leq d \lambda \sum_{i=2}^n |\alpha_i \beta_i| \\
 &\leq d \lambda \sqrt{\left( \sum_{i=2}^n \alpha_i^2 \right) \left( \sum_{i=2}^n \beta_i^2 \right)} \\
 &\leq d \lambda \sqrt{\| \mathbb{1}_S \| \| \mathbb{1}_T \|} = d \lambda \sqrt{|S| |T|}.
 \end{aligned}$$

■

We now see how to save randomness using expanders.

Let  $\mathcal{A}$  be an algorithm that uses  $k$  independent random bits. Let  $G$  be a  $(2^k, d, \lambda)$ -expander. Starting at an arbitrary vertex  $v_1$ , perform a random walk for  $\ell$  steps through vertices  $v_1, v_2, \dots, v_\ell$ . Run the algorithm on each of these inputs  $v_1, \dots, v_\ell$  (interpreting the  $2^k$  elements of  $V(G)$  as length  $k$  bit strings).

Recall that if  $\mathcal{A}$  once (using  $k$  bits) has error probability  $\beta$ , running the algorithm  $\ell$  times (using  $k\ell$  bits) reduces this to error probability  $\beta^\ell$ . It turns out that running the algorithm  $\ell$  times as described above (using  $k + \ell \log d$  bits) reduces the error probability to  $(\beta + \lambda)^\ell$ !

In purely graph theoretic terms, this says the following.

**Theorem 2.8.** Let  $G$  be a  $(n, d, \lambda)$ -expander, and let  $B \subseteq V$  be of size  $\beta n$ . Starting at a random vertex  $v_1$ , consider  $\ell$  steps of the random walk going through vertices  $v_1, v_2, \dots, v_\ell$ . Then,

$$\Pr[\text{all } v_i \text{ are in } B] \leq (\beta + \lambda)^\ell.$$

*Proof sketch.* Consider the diagonal matrix  $D$  with 1s at vertices in  $B$  and 0 elsewhere. Let  $p^{(0)}$  be the initial (uniform) distribution of  $v_0$ . Observe that the  $\ell_1$  norm of  $(BM)^i p^{(0)}$  is precisely the probability that  $v_i$  is in  $B$ . To bound this, we split a given vector into its component along and orthogonal to the uniform distribution. The norm of the second part decreases by  $\lambda$  at every step. ■

**Theorem 2.9** (Alon-Boppana bound). For any  $(n, d, \lambda)$ -expander,

$$\lambda \geq \frac{2\sqrt{d-1}}{d} (1 - o(1)).$$

**Definition 2.11** (Ramanujan Graph). A  $(n, d, \lambda)$ -expander is said to be a *Ramanujan graph* if

$$\lambda \leq \frac{2\sqrt{d-1}}{d}.$$

It was proved in 2014 by Adam Marcus, Daniel Spielman, and Nikhil Srivastava that there exist infinite families of bipartite Ramanujan graphs of every degree greater than 2.

### §3. Reed-Solomon Codes

#### 3.1. Lecture 11: Introduction

##### 3.1.1. Lecture 11

Randomly choosing strings from  $\{0, 1\}^n$  tends to yield a good code, with high distance between points with high probability. This is true even for exponentially many points, say  $2^{0.1n}$ . At the heart of getting good codes is derandomizing this process to get good explicit codes.

In some sense, good codes, expander graphs, and extractors are equivalent. In fact, some of the best known expander constructions today come from coding theoretic constructions.

The subject of coding theory lies at the intersection of numerous disparate fields, such as (theoretical) computer science, electrical engineering, and math. Interestingly, the same objects are studied in all the disciplines, merely from different perspectives.

**Definition 3.1** (Reed-Solomon Code). Let  $\mathbb{F}$  be a finite field, and  $k, n \in \mathbb{N}$  with  $n \geq k, |\mathbb{F}|$ . Also fix some distinct  $\alpha_1, \dots, \alpha_n$ . The message space of the *Reed Solomon code*  $RS(k, n)$  is

$$\{g(x) \in \mathbb{F}[x] : \deg(g) \leq k - 1\}.$$

That is, we identify  $k$ -dimensional vectors in  $\mathbb{F}^k$  with corresponding polynomials. A polynomial  $g$  is encoded as

$$\text{Enc}(g) = (g(\alpha_1), \dots, g(\alpha_n)).$$

Note that the number of possible messages is  $|\mathbb{F}|^k$ . Let us look at some basic properties of this code.

1. A Reed-Solomon code is a linear code. This follows immediately from the fact that  $(\alpha g + h)(\alpha_i) = \alpha g(\alpha_i) + h(\alpha_i)$ , and if  $g, h$  are of degree at most  $k - 1$  then so is  $\alpha g + h$ .
2. The code has rate  $k/n$ .
3. The distance of the code is  $n - k + 1$ . Indeed, by the Fundamental Theorem of Algebra, two polynomials can coincide in value at at most  $k - 1$  points (otherwise their difference, a nonzero polynomial of degree at most  $k - 1$ , would have more than  $k - 1$  roots).

Observe that given a vector  $(g_1, \dots, g_k) \in \mathbb{F}^k$ , we encode it as

$$\begin{pmatrix} 1 & \alpha_1 & \alpha_1^2 & \cdots & \alpha_1^{k-1} \\ 1 & \alpha_2 & \alpha_2^2 & \cdots & \alpha_2^{k-1} \\ \vdots & \vdots & \vdots & \ddots & \vdots \\ 1 & \alpha_n & \alpha_n^2 & \cdots & \alpha_n^{k-1} \end{pmatrix} \begin{pmatrix} g_1 \\ g_2 \\ \vdots \\ g_k \end{pmatrix}.$$

Also note that the above properties do not depend on our choice of  $(\alpha_i)$ .

It turns out that Reed-Solomon codes are optimal in some sense.

**Theorem 3.1.** Reed-Solomon codes match the singleton bound.

This says that over large fields, they essentially match the best possible rate-distance trade-off.

### 3.2. Lecture 11-12: Decoding

#### 3.2.1. Lecture 11 (continued)

The matrix multiplication scheme above describes a simple way to encode RS codes. How do we decode them? That is, if  $r = (r_1, \dots, r_n) \in \mathbb{F}^n$ , what do we decode it to? This asks to find the RS codeword  $c$  closest to  $r$ ; find the RS codeword  $c$  such that  $d_H(c, r) < (n - k + 1)/2$  (if no such codeword exists, say no).

In the 60s, numerous decoding algorithms such as those of Peterson, Sugiyama, and Berlekamp-Massey were proposed. Despite being elementary, these are all very clever. In practice, Berlekamp-Massey is typically used (it is nearly linear in the input size).

We shall look at the Welch-Berlekamp algorithm, which was proposed in the 80s. Amusingly, this was not originally published in a paper but in a patent. Some time later, Madhu Sudan et al. decoded<sup>4</sup> the patent. The algorithm has had surprisingly pervasive effects in computer science, and the methods have helped resolve several open problems in math as well. Interested readers can see the “polynomial method” for more details; while it has been used in the past in math, this introduced it to mainstream theoretical computer science.

The algorithm is as follows.

1. Find a nonzero polynomial  $Q(x, y) = A(x) + yB(x)$  such that
  - (a)  $\deg(B) < \lceil (n - k + 1)/2 \rceil =: d$ ,
  - (b)  $\deg(A) < d + (k - 1) =: D$ , and
  - (c) for all  $i$ ,  $Q(\alpha_i, r_i) = 0$ .
2. Set  $g(x) = -A(x)/B(x)$ . If  $g$  is a polynomial and  $d_H(\text{Enc}(g), r) < (n - k)/2$  and  $\deg(g) \leq k - 1$ , output  $g$ . Otherwise, say that no codeword exists.

Observe that if  $g$  is a polynomial, the polynomial is certainly correct.

We must prove that if  $g$  is not a polynomial, then no codeword exists. First, however, let us describe how to determine the bivariate polynomial  $Q$ .

Set  $D = (n + k)/2$  and  $d = (n - k)/2$ . The desired  $Q$  is of the form

$$Q(x, y) = A_0 + A_1x + \dots + A_{D-1}x^{D-1} + B_0y + B_1xy + \dots + B_{d-1}x^{d-1}y.$$

Then, we wish to determine coefficients  $(A_i)$  and  $(B_i)$  such that for all  $i$

$$Q(x, y) = A_0 + A_1\alpha_i + \dots + A_{D-1}\alpha_i^{D-1} + B_0r_i + B_1r_i\alpha_i + \dots + B_{d-1}r_i\alpha_i^{d-1} = 0.$$

This is just a linear system of  $n$  equations. If  $D + d > n$ , which is indeed true, we may solve it for a non-trivial solution.

To compute  $g$  using  $A, B$ , we just have to do polynomial long division from high school (there are better methods to do this).

All that remains is to check correctness.

**Theorem 3.2.** If there exists a polynomial  $h \in \mathbb{F}[x]$  of degree at most  $k - 1$  such that  $d_H(\text{Enc}(h), r) < (n - k + 1)/2$ , the Welch-Berlekamp algorithms outputs it.

*Proof.* Let  $Q(x, y) = A(x) + yB(x)$  with  $\deg(A) < D$ ,  $\deg(B) < d$ , and  $Q(\alpha_i, r_i) = 0$  for all  $i$  be the polynomial output in the first step of the algorithm.

Consider  $U(x) = Q(x, h(x))$ . This is  $A(x) + h(x)B(x)$ . For starters, the degree of  $U$  is at most  $D - 1 < (n + k - 1)/2$ . If  $h(\alpha_i) = r_i$ , then  $U(\alpha_i) = 0$ . Consequently, the number of zeroes of  $U$  on  $\{\alpha_1, \dots, \alpha_n\}$  is at least the number of agreements between  $\text{Enc}(h)$  and  $r$ . Due to the distance constraint on  $h$ , there are at least  $(n + k - 1)/2$  such agreements. Therefore,  $U$  has more zeroes than degree, so  $U$  must be the zero polynomial (!) and therefore  $g$  is indeed a polynomial and equal to  $h$ . ■

<sup>4</sup>Pardon the pun.

The idea of the construction is just that the first step forces us to have a  $Q$  of large degree, while the second (assuming a valid  $h$  exists) forces  $Q$  to have small degree. The sweet spot in the middle is precisely where we lie.

Next class, we shall look at decoding Reed-Solomon codes beyond the error limit of half the minimum distance. One option is to output any codeword in the given radius. The more useful (albeit more stringent) notion is to output all these codewords. Such algorithms are called “list decoding algorithms”. We must ensure that the amount we go beyond  $(n - k + 1)/2$  is not so high that the list becomes exponentially large. This is guaranteed by the following.

**Theorem 3.3 (Johnson).** Let  $\mathcal{C}$  be a code of block length  $n$  with distance  $\Delta$ . Then, for all  $r \in \mathbb{F}^n$ , the number of codewords in  $\mathcal{C}$  within distance  $n - \sqrt{n(n - \Delta)}$  (the “Johnson radius”) of  $r$  is  $\text{poly}(q, n, \Delta)$ .

For Reed-Solomon codes, this value is equal to

$$n - \sqrt{n(n - (n - k + 1))} \approx n - \sqrt{nk}.$$

This is *far* larger than the minimum distance. If  $k = 0.01n$ , say, then the minimum distance is about  $0.49n$ , but the Johnson radius is about  $0.9n$ ! Next class, we shall describe how to output all the codewords within the prescribed radius.

We also remark that there exist Reed-Solomon codes where we go beyond the Johnson radius while still having polynomially many codewords in the given radius. Such codes are not *explicitly* known, however, and are obtained by taking a sufficiently large field and choosing random evaluation points  $\alpha_i$ . The question of when in general the Johnson bound is not tight does not have many satisfactory answers to date. In particular, we do not know if the Johnson bound is non-tight for *all* Reed-Solomon codes.

### 3.2.2. Lecture 12

Given  $S = \{(\alpha_i, \beta_i)\}_{i=1}^n$ , we aim to find *all* polynomials  $f \in \mathbb{F}_q[x]$ ,  $\deg f < k$ , such that

$$\text{agr}(f, S) := \left| \{u \in [n] : f(\alpha_i) = \beta_i\} \right| \geq t.$$

Last lecture, we saw the setting where  $t \geq (n + k - 1)/2$ .

Due to **Johnson**, the setting where  $t \geq \sqrt{nk}$  enters consideration. We shall now look at how to decode Reed-Solomon codes up to the Johnson radius.

First, we shall look at the case where  $t \geq 2\sqrt{nk}$ . The basic idea is that we take the Welch-Berlekamp algorithm, but look at polynomials that are higher degree in  $y$ .

1. Set  $\ell \approx \sqrt{n/(k-1)}$ . Find nonzero  $Q(x, y) = A_0(x) + A_1(x)y + A_2(x)y^2 + \dots + A_\ell(x)y^\ell$  such that  $\deg(A_i) \leq n/\ell$  for each  $i$ , and  $Q(\alpha_i, \beta_i) = 0$  for each  $i$ .
2. Find all factors of  $Q(x, y)$  of the form  $(y - h(x))$ , where  $\deg(h) < k$  and  $\text{agr}(h, S) \geq t$ . Output all such  $h$ .

This is very similar in spirit to the earlier algorithm. Indeed,  $Q(x, f(x)) = 0$  just says that  $(y - f(x))$  divides  $Q(x, y)$ . The first step of the algorithm is exactly as earlier and amounts to solving a system of linear equations. There is an algorithm, that factors polynomials of  $\deg d$  on  $\mathbb{F}_q$  in time  $\text{poly}(d, \log q)$ .<sup>5</sup> This is rather interesting, as it means we are able to factorize elements of the polynomial ring. Compare this to the integer ring, where we cannot factorize elements efficiently. We do not describe this algorithm.

*Proof of correctness.* The number of variables in the system of linear equations in the first step is  $(\ell + 1)(\frac{n}{\ell} + 1) > n$ , which is more than the number of datapoints, so such a  $Q$  exists.

Let  $f \in \mathbb{F}[x]$  be of degree  $< k$  and agree with  $S$  at more than  $2\sqrt{nk}$  points. Then, we wish to show that  $R(x) :=$

<sup>5</sup>This algorithm is randomized, and no such deterministic algorithm is known

$Q(x, f(x)) \equiv 0$ . We have  $\deg(R) \leq (k-1)\ell + n\ell$ . On the other hand, as before, if  $f(\alpha_i) = \beta_i$ ,  $R(\alpha_i) = Q(\alpha_i, \beta_i) = 0$ . If the number of agreements is more than  $\deg(R)$ , we are done. That is,

$$t > \frac{n}{\ell} + (k-1)\ell.$$

The quantity on the right is minimized for  $2\sqrt{n(k-1)}$  and the corresponding  $\ell$  is approximately  $\sqrt{n/(k-1)}$ . ■

Now, let us modify the algorithm slightly to  $\sqrt{2nk}$ .

1. Set  $D \approx \sqrt{2n(k-1)}$ . Find nonzero  $Q(x, y) = A_0(x) + A_1(x) + y^2 A_2(x) + \dots + y^{D/(k-1)} A_{D/(k-1)}(x)$  such that  $\deg(A_i) \leq D - (k-1)i$  for each  $i$ , and  $Q(\alpha_i, \beta_i) = 0$  for each  $i$ .
2. Find all factors of  $Q(x, y)$  of the form  $(y - h(x))$ , where  $\deg(h) < k$  and  $\text{agr}(h, S) \geq t$ . Output all such  $h$ .

*Proof of correctness.* The key observation is that in the above argument, we can push the degree of most  $A_i$  higher, without affecting the bound on the overall degree. Let  $\deg A_i = D_i$ , so the degree of  $f^i A_i$  is at most  $i(k-1) + D_i$ . If we want the overall degree to be  $D$ , then we get  $\deg(A_i) \leq D - (k-1)i$ .

The new number of variables is approximately

$$\sum_{i=0}^{D/(k-1)} D - (k-1)i \approx \frac{D^2}{k-1} - \frac{D^2}{2(k-1)} = \frac{D^2}{2(k-1)}.$$

So, we want a  $D$  such that the above is greater than  $D$ .

For the second part of the argument, we have  $\deg(R) \leq D$  by definition, so we are fine if  $t > D$ .

Overall, this gives a bound of around  $\sqrt{2n(k-1)}$ . ■

Finally, let us look at how to get a bound of  $t \geq \sqrt{nk}$ . This argument is slightly more involved than the short jump it took to get from  $2\sqrt{nk}$  to  $\sqrt{2nk}$ .

We shall begin with a brief discussion of the *method of multiplicities*, which is something like the polynomial method on steroids.

**Definition 3.2.** A polynomial  $Q(x, y)$  is said to have a zero of multiplicity  $\geq r$  at  $(\alpha, \beta)$  if for all  $i, j$  such that  $i + j < r$ ,

$$\frac{\partial Q}{\partial x^i \partial y^j}(\alpha, \beta) = 0.$$

1. Set  $D, r$  such that  $D/r \approx \sqrt{n(k-1)}$ . Find nonzero  $Q(x, y) = A_0(x) + A_1(x) + y^2 A_2(x) + \dots + y^{D/(k-1)} A_{D/(k-1)}(x)$  such that  $\deg(A_i) \leq D - (k-1)i$  for each  $i$ , and  $Q$  passes through  $(\alpha_i, \beta_i)$  with multiplicity  $r$ .
2. Find all factors of  $Q(x, y)$  of the form  $(y - h(x))$ , where  $\deg(h) < k$  and  $\text{agr}(h, S) \geq t$ . Output all such  $h$ .

*Proof of Correctness.* The number of variables now remains  $\frac{D^2}{2(k-1)}$ , but the number of constraints has increased to about  $\binom{r+1}{2}n$ . So, we require

$$\frac{D^2}{2(k-1)} \geq \binom{r+1}{2}n.$$

Approximately, this requires.

$$\frac{D^2}{r^2} \geq n(k-1).$$

Now, due to our additional multiplicity constraints, if  $f(\alpha_i) = \beta_i$ , then  $R(x) = Q(x, f(x))$  vanishes with multiplicity at least  $r$  at  $\alpha_i$ . Now, we have that there are at most  $D/r$  such distinct  $f$ . The observation is that each point of

agreement gives us  $r$  zeros, not just one.

We require

$$t \geq \frac{D}{r}.$$

In all, this gives us the required bound  $\sqrt{n(k-1)!}$  ■

We again draw attention to the part where we used the fact that a nonzero degree  $d$  univariate polynomial has at most  $d/r$  distinct zeros of multiplicity  $\geq r$ . Using this fact, we can consider another code.

### 3.3. Lecture 14: Multiplicity codes

#### 3.3.1. Lecture 12 (continued)

**Definition 3.3** (Univariate Multiplicity code). Let  $\mathbb{F}$  be a finite field of size at least  $n$ ,  $\alpha_1, \dots, \alpha_n \in \mathbb{F}$ . The message set is  $\{f \in \mathbb{F}[x], \deg f < k\}$ . We map  $f$  to the  $n$ -dimensional vector  $M$  over  $\mathbb{F}^s$ , where

$$(M_i)_j = f^{(j)}(\alpha_i) = \frac{\partial^{j-1} f}{\partial x^{j-1}}(\alpha_i).$$

Note that the messages are encoded in  $\mathbb{F}^s$ , so errors mean errors in the entire vector, not specific derivatives.

The rate of this code is approximately  $k/ns$ , which is worse than before. The distance however, jumps up to  $n - \frac{k-1}{s}!$

A unique decoding algorithm for the multiplicity is very similar to Berlekamp-Welch, and we omit the details.

Next class, we shall prove incredible list decoding results, namely that for any  $\epsilon > 0$ , for sufficiently large  $s$ , multiplicity codes can be efficiently decoded from fractional agreement  $\frac{k}{ns} + \epsilon$ . We can get arbitrarily close to the (hard) bound – we cannot hope to get a degree  $k$  polynomial with fewer than  $k$  datapoints!

Recent state-of-the-art expander graphs are constructed using multiplicity codes!

#### 3.3.2. Lecture 14

When talking about the derivative, we mean the *syntactic* derivative, which evaluates (on exponents of  $x$ ) exactly the same as ordinary derivatives.

Note that in contrast to Reed-Solomon codes, we can allow the degree of the polynomial to be more than  $n$ .

**Theorem 3.4** (Nielsen '01, Kopparty '13, Guruswami-Wang 14). For every  $\epsilon > 0$ , for sufficiently large  $s$ , univariate multiplicity codes are efficiently list decodable from fractional agreement  $\frac{k}{ns} + \epsilon$ .

We can get arbitrarily close to the (hard) bound (!) – we cannot hope to get a degree  $k$  polynomial with fewer than  $k$  datapoints. Further, this can be done with a constant list size, with the constant depending on  $\epsilon$  – this was shown by Kopparty-Saraf-RonZewi-Wooffer '17.

The fraction of agreement here is  $(1 + \epsilon) \frac{k}{sn} = (1 + \epsilon) \cdot (\text{Rate})$ . Compare this to what we had studied about Reed-Solomon codes, where we only had  $\sqrt{\text{Rate}}$ .

The remainder of this section is dedicated to the proof of this theorem; we shall look at the proof/algorithm of this due to Guruswami-Wang.

The input to the algorithm is an  $s \times n$  matrix  $Y$ . We wish to find all polynomials  $p$  of degree at most  $k$  whose encoding has “large” agreement with  $Y$ . More precisely, there is a set  $T \subseteq [n]$  of size greater than  $t$  such that for all  $i \in T$  and  $j \in [s]$ ,

$$p^{(j)}(\alpha_i) = Y_{ji}.$$



Call this set of all polynomials as  $\mathcal{L}$ . We want  $t$  to be as small as possible. Sticking with the Welch-Berlekamp idea, the proof (and algorithm) goes as follows.

1. Find a nonzero  $(m + 2)$ -variate polynomial

$$Q(x, z_0, z_1, \dots, z_m) = z_0 A_0(x) + z_1 A_1(x) + \dots + z_m A_m(x)$$

such that

- $\deg(A_i) < D$  for some  $D$  we shall fix later,
- multiplicity constraints which we shall come up with later, and
- $Q$  “explains” the given data: for every  $j \in [n]$ ,  $Q(\alpha_i, Y_{0,i}, Y_{0,i}, \dots, Y_{m,i}) = 0$ . We want it to explain the top  $m$  rows of the matrix,

2. Show that for all  $p \in \mathcal{L}$ ,

$$Q(x, p(x), p^{(1)}(x), \dots, p^{(m)}(x)) \equiv 0. \quad (3.1)$$

3. Find all low degree solutions to  $Q$  satisfying Equation (3.1). Note that we cannot rely on factoring for this, and it is more complicated.

Set  $R(x)$  equal to the LHS of Equation (3.1) for some polynomial  $p$ , so it is

$$R(x) = A_0 p + A_1 p^{(1)} + \dots + A_m p^{(m)}.$$

If  $Y$  and the encoding of  $p$  agree at  $\alpha_i$ , then  $R(\alpha_i) = 0$ .<sup>6</sup> The multiplicity constraint means that we also want the derivative of  $R$  to be zero at  $\alpha_i$ . We have

$$\frac{dR}{dx} = A_0^{(1)} p + A_0 p^{(1)} + A_1^{(1)} p^{(1)} + A_1 p^{(2)} + \dots + A_m^{(1)} p^{(m)} + A_m p^{(m+1)},$$

so if  $m < s$ ,

$$0 = \left. \frac{dR}{dx} \right|_{\alpha_i} = A_0^{(1)}(\alpha_i) Y_{0,i} + A_0(\alpha_i) Y_{1,i} + A_1^{(1)}(\alpha_i) Y_{1,i} + A_1(\alpha_i) Y_{2,i} + \dots + A_m^{(1)}(\alpha_i) Y_{m,i} + A_m(\alpha_i) Y_{(m+1),i}.$$

So, at each  $i$ , the aforementioned multiplicity constraints correspond to about  $s - m - 1$  additional constraints of the above form.

Now, we would like to set  $D$  in the first step such that it has a solution. There are  $Dm$  variables and  $n(s - m - 1)$  constraints. So, we require  $Dm \geq n(s - m - 1)$ . Set

$$D = \frac{n}{m}(s - m).$$

Let us now look at step 2. For a given polynomial in  $\mathcal{L}$ , the degree of  $R$  is at most  $D + k - 1$ . To ensure that  $R$  is identically zero, we need that  $t(s - m - 1) \geq D + k$ , since each point of agreement gives  $(s - m - 1)$  equations. That is, we need

$$\begin{aligned} t &> \frac{1}{s - m}(D + k) \\ &\approx \frac{n}{m} + \frac{k}{s - m} \\ \frac{t}{n} &> \frac{k}{n(s - m)} + \frac{1}{m}. \end{aligned}$$

Setting  $m$  as around  $1/\epsilon$  and  $s > 1/\epsilon^2$  does the job!

<sup>6</sup>Stopping here would lead to unique decoding, by setting  $m$  as  $s$  or  $s - 1$  or so.

Finally, it remains to see if it is possible to find all low degree solutions  $p$  to  $Q(x, p, p^{(1)}, \dots, p^{(m)}(x)) \equiv 0$ . Let us look at just the trivariate case, with  $Q(x, p, p') \equiv 0$ . That is, we wish to solve

$$A_0(x)p(x) + A_1(x)p^{(1)}(x) + A_2(x)p^{(2)}(x) \equiv 0.$$

Note that the space of all  $p$  satisfying this is a subspace of the space of all polynomials of degree  $< k$ . We may assume wlog that two of the  $A_i$  are nonzero, as the problem is not very interesting otherwise. Suppose that  $A_2 \not\equiv 0$ . This means that there exists some  $\beta \in \mathbb{F}$  such that  $A_2(\beta) \neq 0$ . We can assume wlog that  $\beta = 0$  by “shifting” the axis by  $\beta$  otherwise. Dividing by a constant, we can also assume that the constant term in  $A_2$  is 1, so

$$A_0p + A_1p^{(1)} + (1 + \tilde{A}_2)p^{(2)} \equiv 0,$$

where  $\tilde{A}_2$  has no constant term.

The  $p$  we wish to find is of the form

$$p(x) = p_0 + p_1x + p_2x^2 + \dots + p_{k-1}x^{k-1}.$$

Plugging this into the previous equation, we have

$$A_0(p_0 + p_1 + \dots) + A_1(p_1 + 2p_2x + \dots) + (1 + \tilde{A}_2)(2p_2 + 3 \cdot 2p_3 + \dots) \equiv 0.$$

This means that all the coefficients of the resulting polynomial is zero. This is just a linear system of equations, so we can solve it. It remains to argue that the number of solutions (the list size) is not too large.

The coefficient for the degree 0 coefficient is

$$A_{00}p_0 + A_{10}p_1 + 2p_2 = 0,$$

In fact, note that the equation for the coefficient of degree  $k$  being 0 involves only the first  $k + 2$  coefficients of  $p$ ! Consequently, the solution space lives in a 3-dimensional subspace, so it is solvable. In general, it lives in an  $(m + 1)$ -dimensional subspace.

These constructions give the best bipartite expanders, condensers, and extractors that we know.

## §4. Hardness v. Randomness

### 4.1. Lecture 13, 15: Hardness

#### 4.1.1. Lecture 13

A general question one can ask is this: for any polynomial-time randomized algorithm, is it possible to derandomize it (possibly using more space) to get a polynomial-time deterministic algorithm doing the same job?

It has been shown that given a “hard” function, one can construct very good pseudorandom bits. However, no explicit hard functions are known.

Consider the notion of *worst case hardness*. For example, if we can show for some language that no algorithm that runs in  $O(n^{10})$  can compute the output correctly on all inputs, then the language is hard in some sense.

We also have the notion of *average case hardness*. Here, if we can show for some language that no algorithm that runs in  $O(n^{10})$  can compute the output correctly on more than  $3/4$  of the inputs, then the language is hard in some sense. It is not too difficult to see that average case hardness is a stronger notion than worst case hardness.

Problems that are average case hard yield good pseudorandom generators. Another question of concern is converting worst case hardness to average case hardness, which is done through error correcting codes.

The rough idea behind the second question is the following. Error-correcting codes, when given two words as input that are close, make them far apart.

**Definition 4.1 (RP).** A language  $L$  is said to be in RP if there is a randomized algorithm  $\mathcal{A}$  running in polynomial time such that

1. for  $x \in L$ ,

$$\Pr_r [\mathcal{A}(x, r) = \text{yes}] \geq \frac{1}{2}.$$

2. for  $x \notin L$ ,

$$\Pr_r [\mathcal{A}(x, r) = \text{yes}] \leq 0.$$

For example, the algorithm we saw in Section 2.2 was in RP.

**Definition 4.2 (BPP).** A language  $L$  is said to be in BPP if there is a randomized algorithm  $\mathcal{A}$  running in polynomial time such that

1. for  $x \in L$ ,

$$\Pr_r [\mathcal{A}(x, r) = \text{yes}] \geq \frac{2}{3}.$$

2. for  $x \notin L$ ,

$$\Pr_r [\mathcal{A}(x, r) = \text{yes}] \leq \frac{1}{3}.$$

Here, by randomized algorithms, we mean probabilistic turing machines.

Next, let us look at pseudorandom distributions. The goal of these is to find some universal set of random bits which we can substitute in place of the (ideally) uniformly random bits.

Denote by  $U_m$  the uniform distribution on  $\{0, 1\}^m$ .

The idea of this is that a distribution  $D$  is pseudorandom (with respect to  $\mathcal{A}$ ) if for the given algorithm  $\mathcal{A}$ , for all inputs  $x$ ,

$$\left| \Pr_{r \sim U_m} [B(x, r) = \text{yes}] - \Pr_{r \sim D} [B(x, r) = \text{yes}] \right| \leq \epsilon.$$

One neat way to think about algorithms *with input* is **circuits**.

In fact, any deterministic algorithm can be viewed as a family  $(C_n)_{n \geq 1}$  of circuits, with  $C_n$  computing the output correctly if the input is of size  $n$ . We can view a randomized algorithm as a deterministic one with two inputs,  $x$  and  $r$ . The circuit then has  $n + m$  input leaves, where  $n$  is the size of the input  $x$  and  $m$  is the number of random bits  $r$ . If we fix the  $x$  part of the input, we get a circuit in the remaining input, namely  $r_1, \dots, r_m$ .

**Definition 4.3** (Pseudorandom distribution). A distribution  $D$  on  $\{0, 1\}^m$  is called  $(S, \epsilon)$ -pseudorandom if for any circuit  $C$  on  $m$  input gates and size at most  $S$ ,

$$\left| \Pr_{r \sim U_m} [C(r) = 1] - \Pr_{r \sim D} [C(r) = 1] \right| \leq \epsilon.$$

When the above happens, we say that  $D$  “fools” all circuits of size at most  $S$ .

**Definition 4.4** (Pseudorandom generator). A function  $G : \{0, 1\}^* \rightarrow \{0, 1\}^*$  is said to be a  $m(\ell)$ -pseudorandom generator if for  $r \in \{0, 1\}^\ell$ ,

- (a)  $G(r) \in \{0, 1\}^{m(\ell)}$ ,
- (b)  $G(r)$  can be computed in  $2^{O(\ell)}$  time, and
- (c) the distribution over  $\{0, 1\}^{m(\ell)}$  which takes a uniformly random element  $s$  of  $\{0, 1\}^\ell$  and takes value  $G(s)$  is  $(m(\ell)^3, 1/10)$ -pseudorandom.

The existence of the above would imply that any randomized algorithm in BPP using  $m$  random bits and running time  $m^3$  can be simulated by a deterministic algorithm with running time  $O(2^{O(\ell)} m^3)$ . We merely run the algorithm on  $G(r)$  for all  $r \in \{0, 1\}^\ell$ , and output whatever answer (yes or no) occurs more.

If we want to completely derandomize our randomized polynomial-time algorithm to get a deterministic polynomial time algorithm, we want that  $m = 2^{\Omega(\ell)}$ . In particular, if a pseudorandom generator with  $\ell = O(\log m)$  exists, then  $\text{BPP} = \text{P}$ . When this is true, we say that the PRG has “exponential stretch”.

Our goal in the next few lectures will be to show that “circuit lower bounds” imply the existence of pseudorandom generators.

**Theorem 4.1.** There exists a “PRG” with exponential stretch which satisfies only (a) and (c) in the definition.

*Proof.* Choose a random  $G$  – for each  $s \in \{0, 1\}^\ell$ , set  $G(s)$  to be a uniformly random string in  $\{0, 1\}^m$ . Fix a circuit  $C$  of size at most  $m^3$ . Suppose that  $\Pr_{r \sim U_m} [C(r) = 1] = p$ , and let

$$B_C := \{r \in \{0, 1\}^m : C(r) = 1\}.$$

Note that the random variable

$$X_C := \left| \left\{ s \in \{0, 1\}^\ell : G(s) \in B_C \right\} \right|$$

is the sum of  $2^\ell$  Bernoulli random variables equal to 1 with probability  $p$ . By the Chernoff bound,

$$\Pr_G \left[ |X_C - \mathbb{E}[X_C]| > \epsilon 2^\ell \right] \leq 2e^{-(\epsilon 2^\ell)^2 / 3\mathbb{E}[X_C]} = 2e^{-(\epsilon 2^\ell)^2 / (3p \cdot 2^\ell)} \leq 2e^{-\epsilon^2 2^\ell / 3}.$$

Now, the number of circuits with size at most  $m^3$  is bounded from above by  $2^{3m^3}$ . An application of the union bound yields that

$$\Pr_G [\text{for some } C \text{ of size at most } m^3, |X_C - \mathbb{E}[X_C]| > \epsilon 2^\ell] \leq e^{-(\epsilon^2 2^\ell / 3) + (\ln 2)(3m^3 + 1)}.$$

For  $\epsilon = 1/10$ ,  $\ell = 4 \log m$ , and sufficiently large  $m$ , the RHS is less than 1, so there exists some  $G$  that satisfies (a) and (c) in the definition of a PRG. ■

#### 4.1.2. Lecture 15

In the previous lecture, we had said the following (after defining what a PRG is).

**Theorem 4.2.** If a  $2^{\Omega(\ell)}$ -PRG exists,  $\text{BPP} = \text{P}$ .

Note that a  $2^\ell$ -PRG does not exist. Indeed, if it did, we could design a circuit that is 1 precisely at each point in  $\{0, 1\}^{2^\ell}$  that is mapped to by the PRG, and is 0 everywhere else.

While no PRGs are known that fool *all* circuits of size bounded by  $m(\ell)^3$ , there are PRGs known under more specific conditions on the circuit. For example, we can get a PRG that fools any randomized algorithm that is log-space.<sup>7</sup> It is also known that there exist (non-trivial) PRGs which fool constant-depth circuits.

Now, what are *circuit lower bounds*? We had remarked in the previous lecture that they imply the existence of PRGs.

**Definition 4.5** (Worst-case hardness). For  $f : \{0, 1\}^n \rightarrow \{0, 1\}$ , its *worst-case hardness*  $H_{\text{worst}}(f)$  is the largest number  $S$  such that for any circuit of size at most  $S$ , there exists some  $x \in \{0, 1\}^n$  such that  $C(x) \neq f(x)$ .

We cannot compute the function using a circuit of size any smaller than its worst-case hardness. The implementation of the truth table yields that the worst-case hardness of any function is at most about  $O(2^n)$ .

Does there exist any function which is actually this hard? There are  $2^{2^n}$  functions from  $\{0, 1\}^n \rightarrow \{0, 1\}$ , and there are (about)  $2^S$  circuits of size at most  $S$ . Consequently, some functions do require an  $S$  of at least about  $2^n/n$ .

However, no such function is explicitly known – this is another huge open question! In fact, the hardest explicit function we know has worst-case hardness just  $3n - o(n)$ . As mentioned at the beginning of this section, we can use (worst-case) hard functions to design good pseudorandom generators.

**Definition 4.6** (Average-case hardness). For  $f : \{0, 1\}^n \rightarrow \{0, 1\}$ , its *average-case hardness*  $H_{\text{avg}}(f)$  is the largest number  $S$  such that for any circuit of size  $S$ ,

$$\Pr_{x \sim U_n} [C(x) \neq f(x)] > \frac{1}{2} + \frac{1}{S}.$$

Note that we can trivially get a circuit that is equal to  $f$  with probability  $\geq 1/2$ , either set it as the constant 0 or the constant 1 (depending on which value  $f$  takes more often).

Clearly, the average-case hardness of any function is at least the worst-case hardness.

<sup>7</sup>This does not make sense in our current framework, but it is possible to modify the definition of PRGs appropriately. In this setting, we do not have exponential stretch, but we can go from  $\Omega(\log^2 m)$  to  $m$ . The question of whether  $\text{RL} = \text{L}$  is a huge open question.

## 4.2. Lectures 15, 16: Average-case hardness to derandomization

## 4.2.1. Lecture 15 (contd.)

**Theorem 4.3** (Nisan-Wigderson). If (for sufficiently large  $n$ ) there exists a function computable in time  $2^{O(n)}$  with  $H_{\text{avg}}(f) \geq 2^{2n/3}$ , then there exists a  $(2^{\ell/45})$ -PRG and in particular,  $\text{BPP} = \text{P}$ .

This links the worlds of algorithms (in the time complexity of  $f$ ), circuits, and derandomization.

Before moving to the proof of the above, let us try to go from  $\ell$  to  $\ell + 1$ .

**Proposition 4.4.** Let  $f : \{0, 1\}^\ell \rightarrow \{0, 1\}$  be such that  $H_{\text{avg}}(f) \geq \ell^4$ . Then,  $G$  defined by

$$G(r) = (r_1, \dots, r_\ell, f(r)) = (r, f(r))$$

is a PRG.

That is, we would like to say that the output of the function cannot be predicted for a given input. The above merely says that unpredictability implies indistinguishability.

**Theorem 4.5** (Yao's Theorem). Let  $D$  be a distribution on  $\{0, 1\}^m$ . Suppose that for any  $i$  and any circuit of size  $2S$ ,

$$\Pr_{y \sim D} [C(y_1, \dots, y_i) = y_{i+1}] < \frac{1}{2} + \epsilon.$$

Then, for any circuit  $B$  of size  $S$ ,

$$\left| \Pr_{y \sim D} [B(y) = 1] - \Pr_{y \sim U_m} [B(y) = 1] \right| < m\epsilon.$$

## 4.2.2. Lecture 16

*Proof of Yao's Theorem.* We shall show the contrapositive of the statement. Let  $B$  be a circuit of size  $S$  such that

$$\Pr_{y \sim D} [B(y) = 1] - \Pr_{y \sim U_m} [B(y) = 1] \geq m\epsilon.$$

We remove the modulus because we can consider the probability that it is 0 otherwise. Define a sequence of distributions  $D_0, D_1, \dots, D_m$  as follows, where  $D_i$  is obtained by drawing  $x$  from  $D$ , and then replacing the first  $i$  coordinates with draws from  $U_m$ . That is, a draw is  $(y_1, \dots, y_i, z_{i+1}, \dots, z_m)$ , where  $y \sim D$  and  $z \sim U_m$ . Note that  $D_0 = U_m$  and  $D_m = D$ . Let

$$P_i = \Pr_{r \sim D_i} [B(r) = 1].$$

Because  $P_m - P_0 \geq m\epsilon$ , there is some  $i$  such that  $P_i - P_{i-1} \geq \epsilon$ . Note that  $D_i$  and  $D_{i-1}$  differ only at the  $i$ th bit.

We shall give an algorithm to predict  $y_i$  given  $y_1, \dots, y_{i-1}$  for  $y \sim D$  and a random choice of  $z \sim U_m$ . If  $B(y_1, \dots, y_{i-1}, z_i, \dots, z_m)$ , then output  $z_i$ , and if it is 0 then output  $1 - z_i$ . For the sake of succinctness, let  $x = (y_1, \dots, y_{i-1}, z_i, \dots, z_m)$ . Now, the probability of success is

$$\frac{1}{2} \left( \underbrace{\Pr[B(x) = 1 \mid y_i = z_i]}_{P_i} + \underbrace{\Pr[B(x) = 0 \mid y_i = 1 - z_i]}_{(1-\alpha), \text{ say}} \right)$$

We have

$$P_{i-1} = \Pr[B(x) = 1] = \frac{1}{2} (\Pr[B(x) = 1 \mid y_i = z_i] + \Pr[B(x) = 1 \mid y_i = 1 - z_i]) = \frac{1}{2}(P_i + \alpha).$$

Therefore,

$$\text{probability of success} = \frac{1}{2}(P_i + 1 - \alpha) = \frac{1}{2} + P_i - P_{i-1} \geq \frac{1}{2} + \epsilon.$$

To get the final circuit  $C$ , note that on a random choice of  $z \sim U_m$  in our “algorithm”, we succeed with probability at least  $(1/2) + \epsilon$ . Therefore, there exists some specific choice which gives a probability of success at least  $(1/2) + \epsilon$ , which is precisely what we want. ■

Let us now come to the proof of Theorem 4.3. The idea is as follows. We would like to consider a bunch of subsets of  $[\ell]$ , and apply a hard function  $f$  to each of them to get one extra bit to append. In all, the number of bits we append is the number of subsets we choose. If we choose all subsets to be disjoint, then the resulting new bits are completely independent of each other, but we do not get exponentially many new bits. Therefore, we allow some small amount of intersection of the subsets, and thus some small amount of correlation, without compromising the uncorrelation of the new bits by too much.

**Definition 4.7.** An  $(l, k, d)$ -combinatorial design is a collection  $I_1, \dots, I_r$  of size  $k$  subsets such that for distinct  $i, j \in [r]$ ,  $|I_i \cap I_j| \leq d$ .

**Proposition 4.6.** For  $k = \ell/30, d = k/3$ , there exists an  $(l, k, d)$ -design of size at least  $2^{d/10} \geq 2^{\ell/900}$ .

We do not prove this.

*Proof of Nisan-Wigderson.* Set  $\ell = 900 \log n$ , and  $k$  as from the above.

Fix some combinatorial design  $\mathcal{I} = \{I_1, \dots, I_n\}$  guaranteed by the above, and let  $f : \{0, 1\}^k \rightarrow \{0, 1\}$  be a hard function. Then, the final pseudorandom bits we output are  $f(z_{I_r})$  for each  $r \in [n]$ . For simplicity, denote  $f(I_r) = f(z_{I_r})$ .

Let  $f$  be computable in time  $2^{O(k)}$  and  $H_{\text{avg}}(f) \geq 2^{2k/3}$ . Denote the resulting PRG by  $\text{NW}_\mathcal{I}^f$ . We shall show that  $\text{NW}_\mathcal{I}^f(U_\ell)$  is  $(2^{2k/3}/2, 1/10)$ -pseudorandom, which is  $(n^{20}/2, 1/10)$ -pseudorandom.

Now, we shall use Yao’s Theorem, by showing unpredictability instead. That is, we are done if we show that for any circuit  $C$  of size at most  $n^{20}$ ,

$$\Pr_{z \sim U_\ell} [C(f(z_{I_1}), \dots, f(z_{I_{i-1}})) = f(z_{I_i})] \leq \frac{1}{2} + \frac{\epsilon}{n}.$$

Let  $f_j(z) = f(z_{I_j})$  for each  $j$ .

Suppose otherwise. Let  $z' = z_{[\ell] \setminus I_i}$ , and  $z'' = z_{I_i}$ , so

$$\Pr_{z \sim U_\ell} [C(f_1(z', z''), \dots, f_{i-1}(z', z'')) = f(z'')] > \frac{1}{2} + \frac{\epsilon}{n}.$$

Ignore  $z'$  for now, fixing their values as something (this will be done in precisely the same way as in Yao’s Theorem), abuse notation to denote the new functions by  $f_j$  as well. Then,

$$\Pr_{z \sim U_\ell} [C(f_1(z''), \dots, f_{i-1}(z'')) = f(z'')] > \frac{1}{2} + \frac{\epsilon}{n}.$$

using this, we get a circuit for  $f$  that succeeds with probability at least  $(1/2) + \frac{\epsilon}{n}$ . Note that each  $f_j(z'')$  uses at most  $d$  bits. By taking  $(i-1)$  trivial circuits of  $f$ , which are each of size at most about  $d2^d$ , we get a circuit for  $f$  of size  $d2^d 2^{d/10} + 2^{2d}/2 \leq 2^{2d}$ , contradicting the hardness of  $f$ . ■

### 4.3. Lecture 17: Worst-case hardness to average-case hardness

#### 4.3.1. Lecture 17

Now, we would like to convert a function that is worst-case hard to some other function that is average-case hard.

**Definition 4.8.** Given  $0 < \rho < 1$  and a function  $f : \{0, 1\}^n \rightarrow \{0, 1\}$ , define  $H_{\text{avg}}^\rho(f)$  to be the largest  $S$  such that for any circuit  $C$  of size  $S$ ,

$$\Pr_{x \sim U_m} [f(x) = C(x)] < \rho.$$

In particular, we have  $H_{\text{worst}}(f) = H_{\text{avg}}^1(f)$ .

Given a function  $f : \{0, 1\}^n \rightarrow \{0, 1\}$  with worst-case hardness  $H_{\text{worst}}(f)$ , we may view it as a string  $\text{TruthTable}(f)$  in  $\{0, 1\}^{2^n}$ . For any circuit  $S$  of size less than  $H_{\text{worst}}(f)$ , we have  $\text{TruthTable}(f) \neq \text{TruthTable}(C)$ . This idea of using a few inequalities to generate a large number of inequalities is reminiscent of error correcting codes – maybe we can convert the truth table to an element of  $\{0, 1\}^{2^{2^n}}$ , say, using an encoder  $E$  of relative distance  $1/4$ , and our new average-case hard function  $f'$  is defined by  $\text{TruthTable}(f') = E(\text{TruthTable}(f))$ .

However, this trick does not allow us to show that *all* circuits differ from  $f'$  on a constant fraction! It only does so for circuits whose truth table is equal to some image of the encoding function.

Can we do something in the backward direction? Given a circuit  $B$  on  $2n$  inputs of size  $S^{1/5}$  (say) that agrees with  $f'$  significantly (on a  $\frac{1}{2} + \frac{1}{S}$  fraction), is it possible to construct a small circuit  $C$  on  $n$  inputs of size  $S$  that agrees with  $f$  everywhere?

Let us do something slightly weaker, and suppose that  $B$  agrees with  $f'$  on a 0.9 fraction. This essentially asks us to decode the code. However, the size of the input string to the decoding algorithm is  $2^{2^n}$ , so we need a very fast decoding algorithm.

In usual decoding, we look at the entire corrupted encoded string, and try to retrieve the entire message. In *local decoding*, we read a small part of the corrupted string to recover some portion of the original message. Indeed, finally, we do not want to know the values of  $f$  at *every* point, only at the input point  $x$ .

We can compute the value of  $B$  at this small part, and use it to recover the specific bit corresponding to  $f(x)$ .

**Definition 4.9** (Local decoding). Let  $E : \{0, 1\}^N \rightarrow \{0, 1\}^M$  be an encoder that runs in  $\text{poly}(N)$ . A *local decoder* for handling  $\rho$  errors is an algorithm  $D$  such that given random access to a string  $y \in \{0, 1\}^M$  with  $\Delta(y, E(x)) \leq \rho$  for some  $x \in \{0, 1\}^N$  and an index  $j \in [N]$ , runs in time  $\text{polylog}(M)$  and outputs  $x_j$  with probability at least  $2/3$ .

**Theorem 4.7.** Suppose we have an encoder  $E$  with a local decoder  $D$  for handling  $\rho$  errors. Further suppose we have functions  $f_n : \{0, 1\}^n \rightarrow \{0, 1\}$  in EXP that have worst-case hardness  $H_{\text{worst}}(f) \geq S(n)$ . Then, there exists  $\epsilon > 0$  and  $\hat{f}_m \in \text{EXP}$  on  $m$  bits such that

$$H_{\text{avg}}^{1-\rho}(\hat{f}_m) \geq (S(2\epsilon m))^\epsilon.$$

The proof is exactly as in the preceding paragraphs. We have  $(\hat{f}(x))_{x \in \{0, 1\}^m} = E((f(x))_{x \in \{0, 1\}^n})$ . Let  $N = 2^n$  and  $M = 2^m$ . Suppose instead that  $B$  is a circuit of size  $T = (S(\epsilon m))^\epsilon$  such that

$$\Pr_{x \sim U_m} [B(x) = f(x)] \geq 1 - \rho.$$

Now, we have that

$$\Delta((B(x))_{x \in \{0, 1\}^m}, (\hat{f}(x))_{x \in \{0, 1\}^m}) < \rho.$$



Fix  $x$  to be the string in  $\{0, 1\}^n$  that we wish to obtain the value of (of  $f$ ). Each time we compute  $B(y)$  for some  $y$ , we incur a time of  $T$ .

We still need to eliminate the randomness in the local decoding algorithm. To do this, modify the encoding algorithm by repeating it enough times that the error probability is less than  $1/N^2$ . Then, for a fixed  $f, x$ , this implies that there exists some *fixed* set of random bits which outputs  $f(x)$  correctly for all  $x \in [N]$ .

#### 4.3.2. Lecture 18

It is possible to go from functions with high  $H_{\text{avg}}^{1-\rho}$  to functions which are average-case hard, but we omit the details. In this lecture, we elaborate a bit more on local decoding. Recall Reed-Solomon codes from Section 3. Are these locally decodable?

Given a polynomial

$$m = a_0 + a_1x + \dots + a_{d-1}x^{d-1}$$

and  $(m(\alpha_1), m(\alpha_2), \dots, m(\alpha_n))$ , can we recover a specific  $a_i$  by looking at a few of the  $m(\alpha_i)$ ? They do not seem very suitable for local decoding, so let us look at some other codes that are more amenable to this.

**Definition 4.10** (Reed-Muller codes). Let  $\mathbb{F}$  be a finite field, and  $\ell, d \in \mathbb{N}$  such that  $|\mathbb{F}| > d$ . Also fix  $S_1, \dots, S_\ell \subseteq \mathbb{F}$ . The message space of the *Reed Solomon code*  $\text{RM}(n, \ell, d)$  is

$$\{p(x_1, \dots, x_\ell) \in \mathbb{F}[x_1, \dots, x_\ell] : \deg(p) \leq d\}.$$

A polynomial  $p$  is encoded as

$$\text{Enc}(p) = (p(\alpha_1, \dots, \alpha_\ell))_{\alpha \in (S_1 \times \dots \times S_\ell)}.$$

In our setting, we fix all the  $S_i$  to be  $\mathbb{F}$ .

That is, the encoding goes from  $\mathbb{F}^{\binom{d+\ell}{\ell}}$  to  $\mathbb{F}^{|\mathbb{F}|^\ell}$ . What is the distance of this code? Given a nonzero polynomial  $p$  over  $\ell$  variables of degree at most  $d$ , what is the largest number of zeros it can have?

**Proposition 4.8.** Any polynomial  $p \in \mathbb{F}[x_1, \dots, x_\ell]$  of degree at most  $d$  has at most  $d|\mathbb{F}|^{\ell-1}$  zeros. That is,

$$\Pr_{\alpha \sim \mathbb{F}^\ell} [p(\alpha) = 0] \leq \frac{d}{|\mathbb{F}|}$$

*Proof.* Assume wlog that the degree of  $p$  is  $d$ . The idea is that we will partition  $\mathbb{F}^\ell$  into a bunch of “lines” and show that on each line, the probability is at most  $d/|\mathbb{F}|$ . For  $\alpha \in \mathbb{F}^\ell, r \in \mathbb{F}^\ell$ , consider the line

$$L_{\alpha, r} = \{\alpha + tr : t \in \mathbb{F}\}.$$

We shall show that for some clever choice of  $r$ , the polynomial does not become the zero polynomial on this line for any  $\alpha$ . Restricted to this line, the function becomes a polynomial in  $t$ . We want to show that this is a nonzero polynomial

$$p(\alpha_1 + tr_1, \dots, \alpha_\ell + tr_\ell).$$

in  $t$ . Let  $P_d$  be the degree  $d$  part of  $P$ , and note that the coefficient of  $t^d$  in this polynomial is  $P_d(r_1, \dots, r_\ell)$ , independent of  $\alpha$ ! Further,  $P_d$  cannot be identically zero on  $\mathbb{F}^\ell$  because this would imply that the degree of  $p$  is less than  $d$  (this uses the fact that  $|\mathbb{F}| > d$ ). Therefore, the polynomial is nonzero for some choice of  $r$ . This means that the univariate polynomial is nonzero, so has at most  $d/|\mathbb{F}|$  zeros, and we are done. ■

Are the Reed-Muller codes locally decodable? Let us change our perspective slightly, changing the message space from the coefficients to the evaluations of  $\mathbb{F}$  at some  $\binom{\ell+d}{d}$  (fixed and specific) points – there exists a choice of such

points which uniquely determines the polynomial.

In the absence of errors, this makes local decoding trivial. What do we do in the presence of errors?

Suppose we want to evaluate the polynomial at some point  $\beta$  given the evaluations at all points (with an  $\epsilon$  fraction of errors). If we manage to come up with some line through  $\beta$  that has relatively few errors, then we can use Reed-Solomon decoding on this line to compute what  $p(\beta)$  is precisely. Suppose that we choose this line randomly. Then, the expected number of errors is

$$\begin{aligned} \mathbb{E}_{\text{random line } \ell \text{ through } \beta} [\text{number of corruptions on } \ell] &= \mathbb{1}_{\text{error at } \beta} \frac{1}{(|\mathbb{F}|^\ell - 1)/(|\mathbb{F}| - 1)} (\text{number of errors not at } \beta) \\ &\leq 1 + \epsilon|\mathbb{F}|. \end{aligned}$$

Therefore, by a Markov argument,

$$\Pr_{\ell}[\ell \text{ has less than } 3(\epsilon|\mathbb{F}| + 1) \text{ errors}] \geq \frac{2}{3}$$

and we are done.

In all, we choose a random line through  $\beta$ , apply Reed-Solomon coding on this line, then use the resultant polynomial to compute  $p(\beta)$ .

Here, the local decoding algorithm runs in  $O(|\mathbb{F}|)$  time, which we wish to be  $\text{polylog}(|\mathbb{F}|^\ell)$ . For sufficiently large  $\ell$  ( $\Omega((|\mathbb{F}|/\log |\mathbb{F}|)^\delta)$  for some constant  $\delta > 0$ ), this is indeed true.

When we try to convert this to the binary setting however, one major issue pops up. We can of course view  $\mathbb{F}$  as a string over  $\{0, 1\}^{\log |\mathbb{F}|}$ , but in this case the notion of “error” changes. An  $\epsilon$  fraction corruption means that an  $\epsilon$  fraction of the *bits* are corrupted, not points in  $\mathbb{F}^\ell$ . Indeed, an  $\epsilon$  fraction of bits being corrupted means that an  $\epsilon \log |\mathbb{F}|$  fraction of the points in  $\mathbb{F}^\ell$  could be corrupted.

I would like a coding scheme over the binary alphabet that can tolerate a constant fraction of errors, and Reed-Muller codes do not seem to satisfy this.

The way we fix this is by further concatenating a string over the binary alphabet, to yield a coding known as the *Walsh-Hadamard coding*. We shall see this next lecture.

## References

[Rei08] Omer Reingold. Undirected connectivity in log-space. *J. ACM*, 55(4), sep 2008.