

Exponential separations for one-way quantum communication complexity, with applications to cryptography

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Abstract

We give an exponential separation between one-way quantum and classical communication protocols for two partial Boolean functions, both of which are variants of the Boolean Hidden Matching Problem of Bar-Yossef et al. Earlier such an exponential separation was known only for a relational version of the Hidden Matching Problem. Our proofs use the Fourier coefficients inequality of Kahn, Kalai, and Linial. We also give a number of applications of this separation. In particular, we provide the first example in the bounded storage model of cryptography where the key is secure if the adversary has a certain amount of classical storage, but is completely insecure if he has a similar (or even much smaller) amount of *quantum* storage. Moreover, in the setting of privacy amplification, we show that there exist extractors which yield a classically secure key, but are insecure against a quantum adversary.

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1 Introduction

One of the main goals of quantum computing is to exhibit problems where quantum computers are much faster (or otherwise better) than classical computers. Preferably exponentially better. The most famous example, Shor’s efficient quantum factoring algorithm [Sho97], constitutes a separation only if one is willing to believe that efficient factoring is impossible on a classical computer—proving this would, of course, imply $P \neq NP$. One of the few areas where one can establish *unconditional* exponential separations is communication complexity.

Communication complexity is a central model of computation, first defined by Yao [Yao79], which has found applications in many areas [KN97]. Here two parties, Alice with input x and Bob with input y , collaborate to solve some computational problem that depends on both x and y . They want to do this with minimal communication. The problem to be solved could be either a function $f(x, y)$ or some relational problem where for each x and y , several outputs are valid. The protocols could either be *interactive*, in which case Alice and Bob can send each other as many rounds of communication as they want; *one-way*, in which case Alice sends one message to Bob who then has to determine the output; or *simultaneous*, where Alice and Bob each pass one message to a third party (the *referee*) who determines the output. The *communication complexity* of the problem is the worst-case communication of the best protocol that gives for every input x and y a correct output with probability at least $1 - \varepsilon$, for some fixed constant $\varepsilon \in [0, 1/2)$.

Allowing the players to use *quantum* instead of classical resources can reduce the communication complexity significantly. Examples of problems where quantum communication gives exponential savings were for instance given by Buhrman, Cleve, and Wigderson for zero-error one-way and interactive protocols [BCW98]; by Raz for bounded-error interactive protocols [Raz99]; and by Buhrman, Cleve, Watrous, and de Wolf for simultaneous protocols [BCWW01]. The first two problems are partial Boolean functions, while the third one is a total Boolean function (however that separation doesn’t hold in the presence of public coins). In fact, providing a separation for a total Boolean function in the presence of public coins is one of the main open questions in the area.

Moreover, Bar-Yossef, Jayram, and Kerenidis [BJK04] showed an exponential separation for one-way protocols and simultaneous protocols with public coins, but they only achieve this for a relational problem, called the *Hidden Matching Problem* (HMP). This problem can be solved efficiently by one quantum message of $\log n$ qubits, but classical one-way protocols need to send nearly \sqrt{n} bits to solve it. However, Boolean functions are much more natural objects than relations both in the model of communication complexity and in the cryptographic settings that we consider in this paper. Bar-Yossef et al. stated a Boolean version of their problem (a partial Boolean function) and conjectured that the same quantum-classical gap holds for this problem as well.

1.1 Exponential separation for the Boolean Hidden Matching Problem

In this paper we prove tight bounds for the bounded-error one-way communication complexity of a slight variant of the Boolean Hidden Matching Problem of [BJK04], which we call the *Noisy Perfect Matching problem* (NPM). A precise definition will be given in Section 2.

Theorem 1. *The classical bounded-error one-way communication complexity of the Noisy Perfect Matching problem is $R_\varepsilon^1(\text{NPM}) = \Theta(\sqrt{n})$, while the quantum bounded-error one-way complexity is $Q_\varepsilon^1(\text{NPM}) = \Theta(\log n)$.*

This provides the first exponential separation between quantum and classical one-way communication complexity for a partial Boolean function. Our contribution is the $\Omega(\sqrt{n})$ lower bound for classical one-way protocols. It is proved using the Fourier coefficients inequality of Kahn, Kalai,

Linial [KKL88], which in turn uses the Bonami-Beckner inequality [Bon70, Bec75]. Fourier analysis was previously used in communication complexity by Raz [Raz95] and Klauck [Kla01].

1.1.1 Application: streaming model

In the *streaming model* of computation, the input is given as a stream of bits and the algorithm is supposed to compute or approximate some function of the input, having only space of size S available. See for instance [AMS99, Mut05]. There is a well-established connection between one-way communication complexity and the streaming model (which was pointed out to us by Shengyu Zhang): if we view the input as consisting of two parts x and y , then the content of the memory after x has been processed, together with y , contains enough information to compute $f(x, y)$. Hence, a space- S streaming algorithm for f implies a one-way protocol for f of communication S . The classical lower bound for the above communication complexity problem, together with the observation that our quantum protocol can be implemented in the streaming model, implies a separation between the quantum and classical streaming model. Namely, there is a partial Boolean function f which can be computed in the streaming model with small error probability using quantum space of $O(\log n)$ qubits, but which requires $\Omega(\sqrt{n})$ bits if the space is classical.

Earlier this year, Le Gall [Gal06] constructed a problem that can be solved in the streaming model using $O(\log n)$ qubits of space, while any classical algorithm needs $\Omega(n^{1/3})$ classical bits. His $\log n$ -vs- $n^{1/3}$ separation is a bit smaller than our $\log n$ -vs- \sqrt{n} , but his separation is for a *total* Boolean function while ours is only partial (i.e., requires some promise on the input).

1.2 A variant of the problem, with links to cryptography

Our next result deals with another variant of the Boolean Hidden Matching Problem, called the α -Partial Matching problem (α PM), which is parametrized by some value α as defined in Section 2. The ability to vary this parameter α will be important for some of our applications. For this variant we can also establish an exponential gap:

Theorem 2. *Let $\alpha \in [0, O(1/\sqrt{\log n})]$. The classical bounded-error one-way communication complexity of the α -Partial Matching problem is $R_\varepsilon^1(\alpha\text{PM}) = \Theta(\sqrt{n/\alpha})$, while the quantum bounded-error one-way complexity is $Q_\varepsilon^1(\alpha\text{PM}) = O(\log(n)/\alpha)$.*

For instance for $\alpha \approx 1/\sqrt{\log n}$ the separation is $(\log n)^{3/2}$ qubits versus $\sqrt{n}(\log n)^{1/4}$ classical bits. The quantum protocol for α PM is somewhat less efficient than the quantum protocol for NPM ($(\log n)^{3/2}$ vs $\log n$ qubits), but the former has zero-error while the latter has bounded error.

Intuitively, the proof of this second separation shows that if Alice's message was too short, then Bob has hardly any information about a certain string z that can be computed from x given also Bob's input y . That is, from his perspective (given y and Alice's message) this string z is almost uniformly distributed. Our proof uses a result of Talagrand [Tal96] (which is easy to derive from, again, the KKL inequality, though Talagrand himself proves it differently) and a large deviation inequality for martingales [McD98].

1.2.1 Application: the bounded storage model

Our second result allows us to make some interesting connections with the area of cryptography. The *bounded storage model* was first introduced by Maurer [Mau92] with the aim of implementing information-theoretically secure *key expansion*. In this setting, a large random variable X is publicly but only temporarily available. Alice and Bob use a shared secret key Y to extract an additional key $Z(X, Y)$ from X . The secret key Y remains hidden from the adversary during this extraction phase,

but may be revealed later. The adversary is assumed to have only a bounded amount of storage and as a result his information about Z is limited *even* if he learns the secret key Y afterwards. “Limited information” means that the distribution on $Z(X, Y)$ is η -close to uniform even when conditioned on Y and on the information about X that the adversary stored in his memory, for some small security parameter $\eta \in [0, 1]$ (the smaller the better). Aumann, Ding, and Rabin [ADR02] were the first to prove a bounded-storage scheme secure, and essentially tight constructions have subsequently been found [DM04, Lu04, Vad04]. It is an important open question whether any of these constructions remains secure if the adversary is allowed to store *quantum information*. A number of researchers have conjectured that a bounded-storage protocol secure against classical adversaries with a certain amount of memory, should be roughly as secure against *quantum* adversaries with roughly the same memory bound. Using our second separation we can refute this conjecture.

The link to one-way communication comes from viewing Alice’s input as the temporarily available randomness X , while Bob’s input takes the role of the secret key Y . Alice’s message $m(X)$ (which she sends *without* knowing Y) represents the stored information of the adversary about the string X before he learns the key Y . Our lower bound proof for one-way communication shows that Bob cannot learn much about a certain αn -bit string $Z(X, Y)$ if Alice’s message is too short. This can be translated back to show that an adversary cannot learn much about the extracted key Z if his storage is too small. Our result gives the first example of a bounded storage protocol where the extracted key can be made η -secure against a classical adversary (for any constant η) but become completely insecure against a quantum adversary of the same or even much smaller memory size.

Theorem 3. *Let $\eta \in [0, 1]$ and $\alpha \in [0, O(\sqrt{\eta/\log n})]$. The extracted αn -bit string in the bounded storage protocol derived from the α -Partial Matching problem is η -secure against a classical adversary with memory bound $O(\sqrt{\eta^3 n/\alpha})$, while it is at most $(1 - 2^{-k})$ -secure against a quantum adversary with memory bound $O(k \log(n)/\alpha)$.*

Normally in cryptography one wants η -security for exponentially small η . Our classical bounded storage scheme is not secure in that strong sense, but it is secure for any constant η of our choice. In fact, if for some small $\delta > 0$ we set $\alpha \approx 1/\sqrt{n^\delta}$ (so the extracted key has $\alpha n \approx n^{1-\delta/2}$ bits) and set the classical memory bound to $\approx n^{1/2-5\delta/4}$, then η can be as small as $1/n^\delta$.

1.2.2 Application: extractors and privacy amplification

Our result is also closely related to the notion of strong *randomness extractors*. There the task is to extract almost uniform randomness from an *imperfect* (i.e. non-uniformly) source of randomness X with the help of an independent uniform seed Y . In other words, the output of an extractor is a random variable $Z(X, Y)$, such that the pair $(Y, Z(X, Y))$ is close to uniform. The main parameters of an extractor are the length of the random string Y and the randomness of the imperfect source, which is measured by the *min-entropy* of the source.

Extractors have found numerous applications in computer science, in particular in complexity theory and cryptography. One important application is that of *privacy amplification*, which was introduced in [BBR88, ILL89]. In this setting, Alice and Bob start with a shared random variable X about which the adversary has some partial information $m(X)$ and their goal is to generate a secret key Z about which the adversary has very little information. They can achieve this by communicating an independent uniform seed Y over an insecure channel, and using a strong extractor to generate the key $Z(X, Y)$. Assuming a certain bound on $m(X)$, the key $Z(X, Y)$ is secure even if the adversary has full knowledge of Y . Note that the length of $m(X)$ is an obvious upper bound on the amount of information that $m(X)$ contains about X .

Extractors and privacy amplification can also be considered in the *quantum* case where the prior partial information about the string X is a quantum state. Our communication result, as noted in recent work by König and Terhal [KT06], implies that there exist extractors which yield a classically secure key, but are insecure against a quantum adversary. More specifically, one can think of Alice’s input X as the shared random variable, her message $m(X)$ as the prior partial information of the adversary about X , and Bob’s input Y as the independent uniform seed. Our lower bound shows that in the classical setting, the αn -bit string $Z(X, Y)$ is close to uniform even if the size of the classical prior information $m(X)$ is as large as $O(\sqrt{n})$. However, in the quantum setting the key becomes insecure even if the quantum prior information is of size only $\text{poly}(\log n)$.

The dependence of the security on whether the adversary has quantum or classical memory is quite surprising, particularly in light of the following two facts: first, privacy amplification based on two-universal hashing provides exactly the same security against classical and quantum adversaries. The length of the key that can be extracted is given by the min-entropy both in the classical ([BBR88, ILL89]) and the quantum case ([RK05], [Ren05, Ch. 5]). Second, König and Terhal [KT06] have shown that for protocols that extract just *one* bit, the level of security against a classical and a quantum adversary (with the same information bound) is again comparable.

1.2.3 Application: simulations of quantum protocols by classical protocols

Another application of our second separation is in the context of simulating one-way quantum communication protocols by one-way classical protocols. As noted by Aaronson [Aar06, Section 5], Theorem 2 implies that his general simulation of bounded-error one-way quantum protocols by deterministic one-way protocols

$$D^1(f) = O(mQ_\varepsilon^1(f) \log Q_\varepsilon^1(f)),$$

is tight up to a polylogarithmic factor. Here m is the length of Bob’s input. This holds for any partial Boolean function f . Taking f to be our α PM for $\alpha \approx 1/\sqrt{\log n}$, one can show that $D^1(f) \approx n$, $m \approx \alpha n \log n \approx n\sqrt{\log n}$, $Q_\varepsilon^1(f) \approx (\log n)^{3/2}$.

It also implies that his simulation of quantum bounded-error one-way protocols by classical bounded-error one-way protocols

$$R_\varepsilon^1(f) = O(mQ_\varepsilon^1(f)),$$

cannot be much improved. In particular, the product on the right cannot be replaced by the sum: if we take $f = \alpha$ PM with $\alpha = 1/\sqrt{n}$, then by Theorem 2 we have $R_\varepsilon^1(f) \approx n^{3/4}$, $m \approx \sqrt{n} \log n$, and $Q_\varepsilon^1(f) \approx \sqrt{n} \log n$.

Remark. Let us finally point out that both of our separation results can be modified to give a separation in the simultaneous message passing model between the models of classical communication with shared entanglement and classical communication with shared randomness. Earlier, such a separation was known only for a relational problem [GKRW06], not for a Boolean function.

2 The problems and their quantum and classical upper bounds

We assume basic knowledge of quantum computation [NC00] and (quantum) communication complexity [KN97, Wol02].

Before giving the definitions of our two variants of the Boolean Hidden Matching Problem, we fix some notation. Part of Bob’s input will be a sequence M of αn disjoint edges $e_1 = (i_1, j_1), \dots, e_{\alpha n} = (i_{\alpha n}, j_{\alpha n})$ from $[2n]$, which we call an α -matching. If $\alpha < 1$, the matching is *partial*, if $\alpha = 1$ the

matching is *perfect*. We can also view an α -matching on $[2n]$ as an $(\alpha n \times 2n)$ matrix M over $GF(2)$, where each column corresponds to a number in $[2n]$ and the ℓ -th row corresponds to the ℓ -th edge of the matching. In other words, if the ℓ -th edge of the matching is (i_ℓ, j_ℓ) , then the ℓ -th row of the matrix contains two 1's at the positions i_ℓ and j_ℓ and 0's elsewhere. Let $x \in \{0, 1\}^{2n}$. Then the product Mx is an αn -bit string $z = z_1, \dots, z_\ell, \dots, z_{\alpha n}$ where $z_\ell = x_{i_\ell} \oplus x_{j_\ell}$. Denote by $h(\cdot, \cdot)$ the Hamming distance function and by $h(\cdot)$ the Hamming weight function.

Using this notation, we introduce the two partial functions we study, which differ only in the parameter α and in the promise. We call them the *Noisy Perfect Matching* (NPM) and the α -*Partial Matching* (α PM) respectively.

Alice: $x \in \{0, 1\}^{2n}$

Bob: an α -matching M on $[2n]$ and a string $w \in \{0, 1\}^{\alpha n}$ ($\alpha = 1$ for NPM)

a) Promise for NPM: There is a bit b such that $h(Mx \oplus b^n, w) \leq n/3$

b) Promise for α PM: There is a bit b such that $w = Mx \oplus b^{\alpha n}$

Function value: b

We can draw an analogy with two kinds of noise in transmission channels. In the Noisy Perfect Matching problem, Bob's input w results from the string Mx or $Mx \oplus 1^n$ after at most a $1/3$ -fraction of the bits have been "corrupted". In the α -Partial Matching problem, Bob's input w can be viewed as an n -bit string resulting from a perfect matching followed by the "erasure" of a $(1 - \alpha)$ -fraction of the bits. For the communication complexity separation using the α -Partial Matching problem, we could fix α to an appropriate value. However, the general parametric result is useful for our applications.

Quantum protocol: There is an easy $O(\log(n)/\alpha)$ -qubit protocol for both problems. Alice sends a uniform superposition of her bits $x = x_1 \dots x_{2n}$: $|\psi\rangle = \frac{1}{\sqrt{2^n}} \sum_{i=1}^{2^n} (-1)^{x_i} |i\rangle$. Bob completes his αn edges to a perfect matching in an arbitrary way, and measures with the corresponding set of n 2-dimensional projectors. With probability α he will get one of the edges $e_\ell = (i_\ell, j_\ell)$ of his input. The state then collapses to $(-1)^{x_{i_\ell}} |i_\ell\rangle + (-1)^{x_{j_\ell}} |j_\ell\rangle$, from which Bob can obtain the bit $z_\ell = x_{i_\ell} \oplus x_{j_\ell}$ by measuring in the corresponding $|\pm\rangle$ -basis.

In the case of NPM, Bob outputs $z_\ell \oplus w_\ell$. The protocol is correct with probability at least $2/3$, and by repeating $O(\log(1/\varepsilon))$ times we can achieve correctness $1 - \varepsilon$ for any constant $\varepsilon > 0$.

In the case of α PM, Bob can obtain the bit $b = z_\ell \oplus w_\ell$ with certainty if he has measured one of his edges (which happens with probability α), otherwise he claims ignorance. Note that this protocol is *zero-error* (Bob knows when he didn't learn the bit b) and the success probability can be boosted to $1 - \varepsilon$ if Bob is given $O(\log(1/\varepsilon)/\alpha)$ copies of the log n -qubit state.

The above protocol can be repeated k times in parallel: if Bob is given $O(k/\alpha)$ copies of $|\psi\rangle$, then with high probability (at least while $k \ll \alpha n$) he can learn k bits of z .

Classical upper bound: The following is an easy $O(\sqrt{n/\alpha})$ classical upper bound for both functions. Suppose Alice uniformly picks a subset of $d \approx \sqrt{n/\alpha}$ bits of x to send to Bob. By the birthday paradox, with high probability Bob will have both endpoints of at least one of his αn edges and so he can compute the function value b with good probability. In this protocol Alice would need to send about $d \log n$ bits to Bob, since she needs to describe the d indices as well as their bitvalues. However, by Newman's Theorem [New91], Alice can actually restrict her random choice to picking one out of $O(n)$ possible d -bit subsets, instead of one out of all $\binom{2n}{d}$ possible subsets. Hence $d + O(\log n)$ bits of communication suffice.

In Section 4 we show that for NPM the classical upper bound of $O(\sqrt{n})$ is optimal, and in Section 5 we show for α PM that for $\alpha \ll 1/\sqrt{\log n}$ the classical upper bound of $O(\sqrt{n/\alpha})$ is

optimal. Choosing $\alpha \approx 1/\sqrt{\log n}$ gives a function that can be computed with $O(\log n^{3/2})$ qubits of one-way communication, but needs at least $\Omega(\sqrt{n}(\log n)^{1/4})$ classical bits of communication, which gives the exponential quantum-classical separation for α PM.

3 Strategy of the lower bound proofs

We prove a lower bound on classical communication with shared randomness for the problems of the previous section in two different ways. Let us first describe what is common among both proofs. By the Yao principle [Yao77], it suffices to prove a lower bound for *deterministic* protocols under some “hard” input distribution. For both problems we choose a distribution that is uniform on the x ’s, the matchings M , and b . In the case of α PM this fixes Bob’s second input $w = Mx \oplus b^{\alpha n}$. For the Noisy Perfect Matching problem we will in addition fix a distribution on the n -bit string w in the following way: independently choose each bit w_ℓ such that $\Pr[w_\ell = (Mx)_\ell \oplus b] = 3/4$. In other words, we can think of w as a *noisy* version of $Mx \oplus b^n = z \oplus b^n$ where each bit of $z \oplus b^n$ is flipped with probability $1/4$. Note that if (x, M, b, w) are picked according to this distribution, then the probability that $n/3 \leq h(Mx \oplus b^n, w)$ is exponentially small. Hence, any probabilistic protocol for NPM with error ε' gives a deterministic protocol for this distribution with distributional error $\varepsilon' + o(1)$. Therefore, for the rest of the proof we use this distribution.

Suppose we have a classical deterministic one-way protocol with c bits of communication and error probability at most ε under this distribution for either NPM or α PM. This protocol partitions the set of 2^{2n} x ’s into 2^c sets A_1, \dots, A_{2^c} , one for each possible message. At least half of the x ’s must occur in sets of size at least 2^{2n-c-1} , since the smaller sets together contain fewer than $2^c \cdot 2^{2n-c-1} = 2^{2n-1}$ x ’s. Hence there must be at least one set A that contains at least 2^{2n-c-1} x ’s and has error at most 2ε , otherwise the overall error would be larger than ε . Hereafter we will analyze this set A for both problems.

Our proof for NPM, directly bounds Bob’s probability to learn b . In order to learn b , Bob needs to determine whether his string w comes from a noisy version of $Mx \oplus 0^n$ or of $Mx \oplus 1^n$. We upper bound the total variation distance between these two distributions using Fourier analysis. This gives an upper bound on the size of A , and hence a lower bound on the communication c .

In our proof for α PM, we prove a stronger statement. From Bob’s point of view the following happens when he receives the message corresponding to the set A : a uniformly picked matching M of αn disjoint edges (i_ℓ, j_ℓ) , $\ell \in [\alpha n]$, is given, and an unknown x is picked uniformly from A . As before, define $z_\ell = x_{i_\ell} \oplus x_{j_\ell}$ and $z = z_1 \dots z_{\alpha n}$. Note that z is a function of x and M . Here Bob knows M and that x is a uniformly chosen element from the known set A . Bob needs to figure out whether his second input equals $z \oplus 0^{\alpha n}$ or $z \oplus 1^{\alpha n}$. We will use capital letters to denote the corresponding random variables. We show that Z is close to uniformly distributed when the edges are known but x is not: if the communication c is “small”, the total variation distance (conditioned on M) between Z and the uniform distribution $U_{\alpha n}$ on αn bits is

$$\mathbb{E}_M[d(Z, U_{\alpha n} \mid M)] = \frac{1}{2} \sum_{z \in \{0,1\}^{\alpha n}} |\Pr[Z = z \mid M] - 2^{-\alpha n}| \leq \eta$$

for some small η ; the expectation is taken over uniform M^1 . Then also $\mathbb{E}_M[d(Z \oplus 0^{\alpha n}, U_{\alpha n} \mid M)] \leq \eta$ and $\mathbb{E}_M[d(Z \oplus 1^{\alpha n}, U_{\alpha n} \mid M)] \leq \eta$, and hence by the triangle inequality

$$\mathbb{E}_M[d(Z \oplus 0^{\alpha n}, Z \oplus 1^{\alpha n} \mid M)] \leq \mathbb{E}_M[d(Z \oplus 0^{\alpha n}, U_{\alpha n} \mid M)] + \mathbb{E}_M[d(Z \oplus 1^{\alpha n}, U_{\alpha n} \mid M)] \leq 2\eta.$$

¹Note that we include a factor of $1/2$ in our definition of total variation distance. This means that the distance lies in the interval $[0, 1]$, and if distributions P and Q have distance η , the probability of any event cannot change by more than η , i.e., $|\Pr_P[E] - \Pr_Q[E]| \leq \eta$ for any event E .

Distinguishing between the two distributions $Z \oplus 0^{an}$ and $Z \oplus 1^{an}$ is exactly what Bob needs to do to determine b . It is well known that distinguishing between two distributions with variation distance 2η can be done with probability at most $1/2 + \eta$. Accordingly, if c is “small” then the success probability will be close to $1/2$. Conversely, since Bob’s success probability on the set A is at least $1 - 2\varepsilon$, c must have been large.

The link with the bounded storage model should be clear at this point. Alice’s input plays the role of the uniformly distributed bitstring X that is temporarily publicly available. The message in the communication protocol plays the role of the memory of the adversary, who stores c classical bits that depend on X . Bob’s matching plays the role of the initial secret key Y shared between the two parties who want to expand their shared secret key. Knowing Y , it is easy for these two parties to compute the an -bit string $Z(X, Y)$ from X , even in an online fashion. The fact that from Bob’s perspective the string Z is close to uniform if Alice’s message is too short, corresponds to the fact that the memory-bounded adversary in the cryptographic protocol knows hardly anything about the string $Z(X, Y)$ —even if the adversary learns Y after X ceases to be public.

Similarly, in the setting of privacy amplification, Alice’s input plays the role of the shared random variable X , the message in the communication protocol plays the role of the prior partial information that the adversary has about X (which is upper bounded by the communication c) and Bob’s matching plays the role of the independent uniform seed Y . The fact that Bob’s view about the string Z (which is the new secret key) is close to uniform when Alice’s message is short, means that any adversary that has full knowledge of Y and partial information about X bounded by c , knows hardly anything about the string $Z(X, Y)$.

4 Lower bound for Noisy Perfect Matching

In this section we prove Theorem 1 by giving a lower bound on the classical communication complexity of NPM. We begin by providing a few (standard) definitions.

For functions $f, g : \{0, 1\}^n \rightarrow \mathbb{R}$ we define their inner product and the ℓ_1 and ℓ_2 norms by

$$\langle f, g \rangle = \frac{1}{2^n} \sum_{x \in \{0, 1\}^n} f(x)g(x) \quad , \quad \|f\|_1 = \frac{1}{2^n} \sum_{x \in \{0, 1\}^n} |f(x)| \quad , \quad \|f\|_2^2 = \langle f, f \rangle = \frac{1}{2^n} \sum_{x \in \{0, 1\}^n} |f(x)|^2.$$

In particular $\|f\|_2 \geq \|f\|_1$ because of Cauchy-Schwarz. The Fourier transform of f is a function $\hat{f} : \{0, 1\}^n \rightarrow \mathbb{R}$ defined by

$$\hat{f}(s) = \langle f, \chi_s \rangle = \frac{1}{2^n} \sum_{y \in \{0, 1\}^n} f(y)\chi_s(y),$$

where $\chi_s : \{0, 1\}^n \rightarrow \mathbb{R}$ is the character $\chi_s(y) = (-1)^{y \cdot s}$ with “ \cdot ” being the scalar product; $\hat{f}(s)$ is the Fourier coefficient corresponding to s . We have the following relation between f and \hat{f} :

$$f = \sum_{s \in \{0, 1\}^n} \hat{f}(s)\chi_s,$$

The convolution $f * g : \{0, 1\}^n \rightarrow \mathbb{R}$ for $f : \{0, 1\}^n \rightarrow \mathbb{R}$ and $g : \{0, 1\}^n \rightarrow \mathbb{R}$ is defined as

$$f * g(w) = \frac{1}{2^n} \sum_{y \in \{0, 1\}^n} f(y \oplus w)g(y).$$

Note that with this definition we have $\widehat{(f * g)}(s) = \hat{f}(s) \cdot \hat{g}(s)$.

We will use two tools in our analysis, Parseval’s identity and the KKL lemma.

Lemma 4 (Parseval’s Identity). *For every function $f : \{0, 1\}^n \rightarrow \mathbb{R}$, $\|f\|_2^2 = \sum_{s \in \{0, 1\}^n} \hat{f}(s)^2$.*

Lemma 5 ([KKL88]). *Let f be a function $f : \{0, 1\}^n \rightarrow \{-1, 0, 1\}$. Let $t = |\{x \mid f(x) \neq 0\}|/2^n$ be the uniform probability that $f \neq 0$. Then for every $\delta \in [0, 1]$ we have*

$$\sum_{s \in \{0, 1\}^n} \delta^{h(s)} \hat{f}(s)^2 \leq t^{\frac{2}{1+\delta}}.$$

Proof of Theorem 1. Following Section 3, we can assume that Bob can determine b with probability $1 - 2\varepsilon$ for x drawn uniformly from the set A , which is of size at least 2^{2n-c-1} . This means that he can distinguish whether his string w comes from a “noisy” Mx or from a “noisy” $Mx \oplus 1^n$. Recall that our hard distribution is uniform on the x ’s, the matchings M , and the bit b , and we pick w by independently choosing each bit w_ℓ such that $\Pr[w_\ell = (Mx)_\ell \oplus b] = 3/4$. Call \mathcal{D}_0 the distribution on the strings w induced by our hard distribution when we condition on $x \in A$ and $b = 0$, and call \mathcal{D}_1 the corresponding distribution when $b = 1$. Then

$$d(\mathcal{D}_0, \mathcal{D}_1) \geq 1 - 4\varepsilon \tag{1}$$

since the probability to distinguish two distributions of distance d is at most $(1 + d)/2$. In the following we give an upper bound on $d(\mathcal{D}_0, \mathcal{D}_1)$ using Fourier analysis, which will imply an upper bound on $|A| \geq 2^{2n-c-1}$ (and hence a lower bound on c).

To express $d(\mathcal{D}_0, \mathcal{D}_1)$, we will define the following probability distributions. Let μ be the distribution on a bit such that $\mu(0) = 3/4$ and $\mu(1) = 1/4$. For $b \in \{0, 1\}$ define the product distributions on $\{0, 1\}^n$ as $f_b(y) = \prod_{i=1}^n \mu(y_i \oplus b)$. In other words, f_0 is the distribution on n -bit strings where each bit is independently 0 with probability $3/4$ and 1 with probability $1/4$ and f_1 is the same distribution with the bits flipped. They represent the “noise” added to z . We also define the distribution $g_{\mathcal{M}}$ induced on the strings $z = Mx \in \{0, 1\}^n$ by picking a uniform $x \in A$ and a uniform matching M from the set of perfect matchings \mathcal{M} : let $g_{\mathcal{M}} = \frac{1}{|\mathcal{M}|} \sum_{M \in \mathcal{M}} g_M$ where

$$g_M(z) = \frac{|\{x \in A \mid Mx = z\}|}{|A|}.$$

The distribution \mathcal{D}_0 can be viewed as first picking a z according to $g_{\mathcal{M}}$ and then adding noise picked according to f_0 . This can be expressed as the *convolution* of f_0 and $g_{\mathcal{M}}$, i.e.

$$\Pr_{\mathcal{D}_0}[w] = \sum_{z \in \{0, 1\}^n} f_0(z \oplus w) \cdot g_{\mathcal{M}}(z) = 2^n \cdot f_0 * g_{\mathcal{M}}(w)$$

and similarly for \mathcal{D}_1 . This gives

$$\begin{aligned} d(\mathcal{D}_0, \mathcal{D}_1) &= \frac{1}{2} \sum_{w \in \{0, 1\}^n} \left| \Pr_{\mathcal{D}_0}[w] - \Pr_{\mathcal{D}_1}[w] \right| = 2^{n-1} \sum_{w \in \{0, 1\}^n} |f_0 * g_{\mathcal{M}}(w) - f_1 * g_{\mathcal{M}}(w)| \\ &= 2^{2n} \left\| \frac{f_0 - f_1}{2} * g_{\mathcal{M}} \right\|_1 \leq \frac{2^{2n}}{|\mathcal{M}|} \sum_{M \in \mathcal{M}} \left\| \frac{f_0 - f_1}{2} * g_M \right\|_1. \end{aligned} \tag{2}$$

To get an *upper* bound on $d(\mathcal{D}_0, \mathcal{D}_1)$ in terms of the Fourier coefficients, we upper bound the ℓ_1 norm by the ℓ_2 norm and use Parseval’s identity (Lemma 4) to go to the Fourier domain

$$\left\| \frac{f_0 - f_1}{2} * g_{\mathcal{M}} \right\|_1^2 \leq \left\| \frac{f_0 - f_1}{2} * g_{\mathcal{M}} \right\|_2^2 = \sum_{s \in \{0, 1\}^n} \left(\widehat{\frac{f_0 - f_1}{2} * g_{\mathcal{M}}}(s) \right)^2 = \sum_{s \in \{0, 1\}^n} \frac{\widehat{f_0 - f_1}(s)^2}{2} \cdot \widehat{g_{\mathcal{M}}}(s)^2 \tag{3}$$

It is easy to see that the Fourier coefficients of $\frac{f_0 - f_1}{2}$ are

$$\widehat{\frac{f_0 - f_1}{2}}(s) = \begin{cases} \frac{1}{2^{n+k}} & \text{for } s \text{ with } h(s) = k \text{ and } k \text{ odd} \\ 0 & \text{otherwise} \end{cases} \quad (4)$$

We will now relate the Fourier coefficients of g_M , the distribution induced by a matching M , to the flat distribution on $|A|$. Let $g : \{0, 1\}^{2n} \rightarrow \mathbb{R}$ be the uniform distribution over the set A

$$g(x) = \begin{cases} \frac{1}{|A|} & \text{for } x \in A \\ 0 & \text{for } x \notin A \end{cases}$$

Note that for $x \in \{0, 1\}^{2n}$ and $s \in \{0, 1\}^n$ we have $(Mx)^T \cdot s = (x^T M^T) \cdot s = x^T \cdot (M^T s)$. By the definition of g_M , its Fourier coefficients are

$$\begin{aligned} \widehat{g_M}(s) &= \frac{1}{2^n} \sum_{y \in \{0, 1\}^{2n}} g_M(y) (-1)^{y^T \cdot s} = \frac{1}{2^n |A|} \left(|\{x \in A \mid (Mx)^T \cdot s = 0\}| - |\{x \in A \mid (Mx)^T \cdot s = 1\}| \right) \\ &= \frac{1}{2^n |A|} \left(|\{x \in A \mid x^T \cdot (M^T s) = 0\}| - |\{x \in A \mid x^T \cdot (M^T s) = 1\}| \right) \\ &= \frac{1}{2^n} \sum_{x \in \{0, 1\}^{2n}} g(x) (-1)^{x^T \cdot (M^T s)} = 2^n \cdot \hat{g}(M^T s). \end{aligned} \quad (5)$$

We rewrite inequality (2) in the Fourier domain and combine it with (1), (3), (4), and (5)

$$\begin{aligned} (1 - 4\varepsilon)^2 &\leq d(\mathcal{D}_0, \mathcal{D}_1)^2 \leq 2^{4n} \left(\frac{1}{|\mathcal{M}|} \sum_{M \in \mathcal{M}} \left\| \frac{f_0 - f_1}{2} * g_M \right\|_1 \right)^2 \\ &\leq \frac{2^{4n}}{|\mathcal{M}|} \sum_{M \in \mathcal{M}} \left\| \frac{f_0 - f_1}{2} * g_M \right\|_1^2 \leq \frac{2^{4n}}{|\mathcal{M}|} \sum_{M \in \mathcal{M}} \sum_{s \in \{0, 1\}^n} \widehat{\frac{f_0 - f_1}{2}}(s)^2 \cdot \widehat{g_M}(s)^2 \\ &\leq \frac{2^{4n}}{|\mathcal{M}|} \sum_{M \in \mathcal{M}} \sum_{\substack{s: h(s)=k \\ k \text{ odd}}} \frac{1}{2^{2k}} \cdot \hat{g}(M^T s)^2. \end{aligned} \quad (6)$$

Note that $h(M^T s) = 2h(s)$ and hence if $h(s)$ is odd, then $h(M^T s) = 2 \pmod{4}$. For $k = 2 \pmod{4}$ we define γ_k as follows: Let $v \in \{0, 1\}^{2n}$ be a string of Hamming weight k and M be a random matching. Then $\gamma_k = \Pr_M[\exists s \text{ s.t. } v = M^T s]$. Note that this probability depends only on k and not on the specific string v . With this definition we have

$$\sum_{\substack{s: h(s)=k \\ k \text{ odd}}} \frac{1}{2^{2k}} \frac{1}{|\mathcal{M}|} \sum_{M \in \mathcal{M}} \hat{g}(M^T s)^2 = \sum_{\substack{v: h(v)=2k \\ k \text{ odd}}} \frac{1}{2^{2k}} \gamma_{2k} \hat{g}(v)^2 = \sum_{\substack{v: h(v)=k \\ k=2 \pmod{4}}} \frac{1}{2^k} \gamma_k \hat{g}(v)^2 \quad (7)$$

Call $\delta_k = \gamma_k^{1/k}$. Combining (6) and (7) we get

$$(1 - 4\varepsilon)^2 \leq 2^{4n} \sum_{\substack{k=2 \\ k=2 \pmod{4}}}^{2n} \frac{1}{2^k} \sum_{v: h(v)=k} (\delta_k)^{h(v)} \hat{g}(v)^2 \leq 2^{4n} \sum_{\substack{k=2 \\ k=2 \pmod{4}}}^{2n} \frac{1}{2^k} \sum_{v \in \{0, 1\}^{2n}} (\delta_k)^{h(v)} \hat{g}(v)^2 \quad (8)$$

We can upper bound γ_k : For any even number $t \geq 2$, let $N(t)$ be the number of perfect matchings on $[t]$. Then, $N(2) = 1$, $N(t) = (t-1)N(t-2)$. It is not hard to see that the expression for γ_k is

$$\gamma_k = \frac{N(k)N(2n-k)}{N(2n)} = \frac{(k-1)(k-3)\cdots 1}{(2n-1)(2n-3)\cdots(2n-k+1)} \leq \left(\frac{k}{2n}\right)^{k/2}.$$

Then $0 \leq \delta_k \leq \sqrt{\frac{k}{2n}} \leq 1$ for $k \in [2, 2n]$. We now apply the KKL inequality (Lemma 5) to the function $g \cdot |A|$ (hence $t = |A|/2^{2n}$) and use $|A| \geq 2^{2n-c-1}$

$$\sum_{v \in \{0,1\}^{2n}} (\delta_k)^{h(v)} \hat{g}(v)^2 \leq \frac{1}{|A|^2} \left(\frac{|A|}{2^{2n}}\right)^{\frac{2}{1+\delta_k}} \leq 2^{-4n} \left(\frac{2^{2n}}{|A|}\right)^{2\delta_k} \leq 2^{-4n} \cdot 2^{(c+1)\sqrt{\frac{2k}{n}}}.$$

Finally, inequality (8) implies

$$(1 - 4\varepsilon)^2 \leq \sum_{\substack{k=2 \\ k=2(\text{mod } 4)}}^{2n} 2^{-k+(c+1)\sqrt{\frac{2k}{n}}} = \sum_{\substack{k=2 \\ k=2(\text{mod } 4)}}^{2n} 2^{-k/2} \left(2^{-k/2+(c+1)\sqrt{\frac{2k}{n}}}\right)$$

Since $\sum_{k=2; k=2(\text{mod } 4)}^{2n} 2^{-k/2} = \sum_{k=1; k \text{ odd}}^n 2^{-k} \leq \frac{2}{3}$ there is a k such that $\frac{3}{2}(1-4\varepsilon)^2 \leq 2^{-k/2+(c+1)\sqrt{\frac{2k}{n}}}$. Hence, for sufficiently small ε (so that $\frac{3}{2}(1-4\varepsilon)^2 \geq 1$) we have $c \geq \frac{\sqrt{n}}{2} - 1$. \square

5 Lower bound for α -Partial Matching

In this section we study the α PM problem. We start again from Section 3 and analyze the distribution of the αn -bit string Z , which depends on the known matching M with edges $e_1 = (i_1, j_1), \dots, e_{\alpha n} = (i_{\alpha n}, j_{\alpha n})$ as well as on the unknown $x \in A$. Remember that $|A| \geq 2^{2n-c-1}$. Intuitively, if c is small (i.e. A is large), then for most M and strings $z \in \{0,1\}^{\alpha n}$ we should have $\Pr[Z = z \mid M] \approx 2^{-\alpha n}$ and hence $\mathbb{E}_M[d(Z, U_{\alpha n} \mid M)]$ is small. Proving this will be quite technical.

We view the edges of M as being picked one by one. Since A is quite large, for most (i, j) -pairs roughly equally many x 's should have $x_i \oplus x_j = 1$ as have $x_i \oplus x_j = 0$. Thus we expect the first bit Z_1 to be close to uniformly distributed when x is picked uniformly from A . Similarly, we would like the later bits Z_ℓ to be more or less uniform when conditioned on values $Z_1 = z_1, \dots, Z_{\ell-1} = z_{\ell-1}$ for the earlier edges. More formally, once $(i_1, j_1), \dots, (i_{\ell-1}, j_{\ell-1})$ and $z_1, \dots, z_{\ell-1}$ have been fixed, we define the " ℓ th bias" by

$$\beta_\ell = \Pr_{x \in A} [Z_\ell = 1 \mid Z_1 = z_1, \dots, Z_{\ell-1} = z_{\ell-1}, M] - 1/2.$$

This is a function of the first ℓ edges of M and of the first $\ell-1$ bits of Z . Though we write ' M ' in the conditional for brevity, β_ℓ is actually independent of the last $\alpha n - \ell$ edges of M . It is positive if Z_ℓ is biased towards 1, and negative if Z_ℓ is biased towards 0. Note that a fixed M, z pair fully determines all biases $\beta_1, \dots, \beta_{\alpha n}$ and that

$$\Pr_{x \in A} [Z = z \mid M] = \prod_{\ell=1}^{\alpha n} \Pr_{x \in A} [Z_\ell = z_\ell \mid Z_1 = z_1, \dots, Z_{\ell-1} = z_{\ell-1}, M] = \prod_{\ell=1}^{\alpha n} (1/2 - (-1)^{z_\ell} \beta_\ell)$$

Fixing the first $\ell-1$ edges of M and conditioning on their bitvalues $Z_1 = z_1, \dots, Z_{\ell-1} = z_{\ell-1}$ will shrink the set of possible x 's. Let A_ℓ be the subset of A that is still consistent. Initially we have

$|A_1| = |A| \geq 2^{2n-c-1}$. When we pick the next edge (i_ℓ, j_ℓ) and its value z_ℓ , the new set $A_{\ell+1}$ will be smaller by a factor $1/2 + \beta_\ell$ if $z_\ell = 1$ and by a factor $1/2 - \beta_\ell$ if $z_\ell = 0$. We have

$$|A_\ell| = |A| \cdot \Pr_{x \in A} [Z_1 = z_1, \dots, Z_{\ell-1} = z_{\ell-1} \mid M] = |A| \cdot \prod_{i=1}^{\ell-1} (1/2 - (-1)^{z_i} \beta_i).$$

Hence we expect the set to shrink by a factor of about two for each new edge and bitvalue for that edge, i.e., $|A_\ell| \geq 2^{2n-c-\ell}$.

We use a result of Talagrand [Tal96] to relate the expected squared bias β_ℓ^2 to the size of the set A_ℓ . Talagrand himself derived this using a large deviation inequality from [LT91], but Oded Regev showed us how it can be obtained in a simple way from the KKL inequality.

Lemma 6 ([Tal96], Eq. (2.9)). *There is an absolute constant $K \geq 1$ such that for all $A \subseteq \{0, 1\}^{2n}$*

$$\sum_{i,j \in [2n], i \neq j} \beta_{ij}^2 \leq K \left(\log \frac{K 2^{2n}}{|A|} \right)^2,$$

where $\beta_{ij} = \Pr_{x \in A} [x_i \oplus x_j = 1] - 1/2$.

Proof. Let $f : \{0, 1\}^{2n} \rightarrow \{0, 1\}$ be the characteristic function of our set A , and $t = |A|/2^{2n}$. Let $s_{ij} \in \{0, 1\}^{2n}$ be the string having a 1 only at positions i and j . Then we have

$$\hat{f}(s_{ij}) = \frac{1}{2^{2n}} \sum_{y \in \{0, 1\}^{2n}} f(y) (-1)^{y \cdot s_{ij}} = \frac{|A|}{2^{2n}} \cdot \frac{|\{y \in A \mid y \cdot s_{ij} = 0\}| - |\{y \in A \mid y \cdot s_{ij} = 1\}|}{|A|} = 2t\beta_{ij}.$$

Applying KKL (Lemma 5) to f , for every $\delta \in [0, 1]$ we have:

$$\sum_{i,j \in [2n], i \neq j} \delta^2 \hat{f}(s_{ij})^2 \leq \sum_{s \in \{0, 1\}^{2n}} \delta^{h(s)} \hat{f}(s)^2 \leq t^{2/(1+\delta)}.$$

Hence

$$\sum_{i,j \in [2n], i \neq j} \beta_{ij}^2 = \frac{1}{4t^2} \cdot \frac{1}{\delta^2} \sum_{i,j \in [2n], i \neq j} \delta^2 \hat{f}(s_{ij})^2 \leq \frac{1}{4t^2} \cdot \frac{1}{\delta^2} t^{2/(1+\delta)} = \frac{1}{4\delta^2} \cdot t^{-2\delta/(1+\delta)} \leq \frac{1}{4\delta^2} \cdot t^{-2\delta}.$$

Picking $\delta = 1/\log(1/t) = 1/\log(2^{2n}/|A|)$ gives the lemma. \square

This will allow us to show that β_ℓ is probably quite small *if* the set A_ℓ hasn't shrunk too fast. We allow some more shrinking than we expect: note the '3c' instead of 'c' in the exponent below. The way to read this corollary is as follows: the first $\ell - 1$ edges of M and the first $\ell - 1$ bits of z have already been fixed. This determines the set A_ℓ , and we assume this set is large enough. Choosing the ℓ th edge of M will now determine the value of the ℓ th bias β_ℓ . The corollary bounds the expectation of β_ℓ^2 , where the expectation is taken over all choices for the ℓ th edge of M .

Corollary 7. *There is an absolute constant $\gamma > 0$ such that if $|A_\ell| \geq 2^{2n-3c-\ell}$, then*

(1) $\mathbb{E}[\beta_\ell^2] \leq \gamma(c/n)^2$ and (2) $\Pr[|\beta_\ell| \geq \varepsilon] \leq \gamma(\frac{c}{n\varepsilon})^2$.

Proof. Note that fixing a bitvalue for the parity of an edge means that the two bits in that edge behave as one bit. Accordingly, we can view the set A_ℓ as a set of strings of length $m = 2n - (\ell - 1)$

bits. We can upper bound the sum of biases over all possible new edges (excluding ones touching earlier edges) by the sum over all possible edges (including ones touching earlier edges):

$$\sum_{i,j \in [m] \setminus \{i_1, \dots, i_{\ell-1}, j_1, \dots, j_{\ell-1}\}, i \neq j} \beta_{i_\ell, j_\ell}^2 \leq \sum_{i,j \in [m], i \neq j} \beta_{ij}^2 \leq O(c^2),$$

where the last inequality is by applying Lemma 6 to A_ℓ . Dividing by the number $\binom{2n-2(\ell-1)}{2} = \Theta(n^2)$ of possible new edges proves part (1). Part (2) now follows from Chebyshev's inequality. \square

Note that on the one hand we need to assume that the sets A_ℓ are not too small in order to show that the biases β_ℓ are probably not too large (via Corollary 7). But on the other hand we need to show that the earlier biases are not too large in order to be able to conclude that A_ℓ is not too small. To deal with this problem, below we give a proof in two “passes”. The first pass is quite coarse-grained and shows that (with high probability) the sets A_ℓ won't shrink by a factor of 2^{-2c} more than what we expect. This allows us to apply Corollary 7 to each of the αn biases during the second pass. In this second, more fine-grained pass we actually show that $d(Z, U_{\alpha n})$ is small, conditioned on M .

5.1 First pass: the sets A_ℓ probably don't shrink too much

We can only use Corollary 7 if the condition $|A_\ell| \geq 2^{2n-3c-\ell}$ is satisfied. We now show that with high probability (over the uniform distribution on M, z) this is indeed the case for all ℓ simultaneously. The proof uses the following concentration result from [McD98].

Lemma 8 ([McD98], special case of Thm. 3.7). *Let S_1, \dots, S_k be bounded random variables satisfying $E[S_j | S_1 = s_1, \dots, S_{j-1} = s_{j-1}] = 0$ for all $1 \leq j \leq k$ and all values s_1, \dots, s_k . Then for all $t, v \geq 0$*

$$\Pr \left[\sum_{j=1}^k S_j \geq t \right] \leq e^{-t^2/2v} + \Pr \left[\sum_{j=1}^k S_j^2 \geq v \right].$$

Lemma 9. *Let $\eta \in [0, 1]$ and $\alpha \leq \sqrt{\eta/256\gamma \log n}$. If the communication is $c \leq \sqrt{\eta n/64\alpha\gamma}$, then with probability at least $1 - \eta$ (over uniformly chosen M, z) the following holds: for each $\ell \in [\alpha n]$ we have $|A_\ell| \geq 2^{2n-3c-\ell}$ and $|\beta_\ell| \leq 1/4$.*

Proof. Assume $c = \sqrt{\eta n/64\alpha\gamma}$ for simplicity. Defining $S_i = -(-1)^{z_i} 2\beta_i$, we have

$$|A_\ell| = |A| \cdot \Pr_{x \in A} [Z_1 = z_1, \dots, Z_{\ell-1} = z_{\ell-1} \mid M] = |A| \cdot \prod_{i=1}^{\ell-1} (1/2 - (-1)^{z_i} \beta_i) \geq 2^{2n-c-\ell} \prod_{i=1}^{\ell-1} (1 + S_i).$$

To lower bound $|A_\ell|$ it thus suffices to lower bound $\prod_{i=1}^{\ell-1} (1 + S_i)$ by 2^{-2c} under distribution \mathcal{P} , which is uniform on the matching M and on z . Taking natural logarithms, we need to show for any ℓ

$$\ln \left(\prod_{i=1}^{\ell-1} (1 + S_i) \right) = \sum_{i=1}^{\ell-1} \ln(1 + S_i) \geq -c 2 \ln(2). \quad (9)$$

Let us divide the αn ℓ 's into blocks of size c each: for $1 \leq k \leq \alpha n/c$ define the k th block $B_k = \{(k-1)c+1, \dots, kc\}$ (assume for simplicity that $\alpha n/c$ is an integer). Let E_k be the following event:

- (a) $|\beta_i| \leq 1/4$ for each $i \in B_k$ and
- (b) $\sum_{i \in B_k} \ln(1 + S_i) \geq -c^2 \ln(2)/\alpha n$.

We will show below in Claim 10 that for all k , $\Pr_{\mathcal{P}}[\neg E_k \mid E_1, \dots, E_{k-1}] \leq \frac{c}{\alpha n} \eta$. This implies

$$\Pr_{\mathcal{P}}[\neg(E_1, \dots, E_{\alpha n/c})] \leq \sum_{k=1}^{\alpha n/c} \Pr_{\mathcal{P}}[\neg E_k \mid E_1, \dots, E_{k-1}] \leq \frac{\alpha n}{c} \cdot \frac{c}{\alpha n} \eta = \eta.$$

If $E_1, \dots, E_{\alpha n/c}$ all hold, then from (b) for all k we have

$$\sum_{i=1}^{k \cdot c} \ln(1 + S_i) \geq -k \cdot c^2 \ln(2) / \alpha n \geq -c \ln(2)$$

and in particular Eq. (9) holds (even with righthand side of $-c \ln(2)$ instead of $-c2 \ln(2)$) whenever $\ell - 1$ is a multiple of c . For the other ℓ , pick k such that $\ell - 1 \in B_{k+1}$ and note that thanks to (a) we have $\ln(1 + S_i) \geq -\ln(2)$ and hence

$$\sum_{i=1}^{\ell-1} \ln(1 + S_i) = \sum_{i=1}^{kc} \ln(1 + S_i) + \sum_{i=kc+1}^{\ell-1} \ln(1 + S_i) \geq -c \ln(2) + \sum_{i=kc+1}^{\ell-1} -\ln(2) \geq -c2 \ln(2).$$

It thus remains to prove

Claim 10. For all $1 \leq k \leq \frac{\alpha n}{c}$, we have $\Pr_{\mathcal{P}}[\neg E_k \mid E_1, \dots, E_{k-1}] \leq \frac{c}{\alpha n} \eta$.

Proof. From the definition of the events E_i we have that

$$\Pr_{\mathcal{P}}[\neg E_k \mid E_1, \dots, E_{k-1}] \leq \Pr_{\mathcal{P}}[\neg(a) \mid E_1, \dots, E_{k-1}] + \Pr_{\mathcal{P}}[\neg(b) \mid E_1, \dots, E_{k-1}, (a)]$$

We bound the two terms on the righthand side separately, starting with the first.

Let $\ell_1 = (k-1)c$ be the last index in B_{k-1} . Conditioning on E_1, \dots, E_{k-1} means that $|A_{\ell_1+1}| \geq 2^{2n-2c-\ell_1-1}$. For each $i \in B_k$, if $|\beta_{\ell_1+1}|, \dots, |\beta_{i-1}| \leq 1/4$ then as before $\sum_{j=\ell_1+1}^{i-1} \ln(1 + S_j) \geq -c \ln(2)$ and hence $|A_i| \geq |A_{\ell_1+1}| \cdot 2^{-(i-\ell_1-1)-c} \geq 2^{2n-3c-i}$. By Corollary 7 (part 2) we have

$$\Pr_{\mathcal{P}}[|\beta_i| > 1/4] \leq \gamma \left(\frac{4c}{n} \right)^2 = \frac{\eta}{4\alpha n}$$

and hence (a) fails to hold for block k with probability

$$\Pr_{\mathcal{P}}[\neg(a) \mid E_1, \dots, E_{k-1}] \leq \frac{\eta c}{4\alpha n}$$

Next, conditioning on (a) and E_1, \dots, E_{k-1} we show that (b) holds for B_k with probability at least $1 - 3\eta c / 4\alpha n$, which implies the claim. Let \mathcal{P}' be the distribution on the edges and the string z when we condition on (a). We make some observations about \mathcal{P}' . First, like \mathcal{P} , we can view \mathcal{P}' as picking edges and bits z_i sequentially: select the i th edge uniformly at random among all edges that are disjoint from those already chosen *and* that have bias $\leq 1/4$; then pick z_i uniformly at random. The difference with \mathcal{P} is that the i th edge is not picked arbitrarily, but is restricted to edges having bias $\leq 1/4$. Second, the condition of Lemma 8 holds for each $S_i = -(-1)^{z_i} 2\beta_i$: the conditional expectations are all 0, because we first determine β_i and then give S_i the sign $+$ or $-$ with equal probability.

Since $|S_i| = 2|\beta_i| \leq 1/2$ for $i \in B_k$, we have $\ln(1 + S_i) \geq S_i - S_i^2$, and hence

$$\sum_{i \in B_k} \ln(1 + S_i) \geq \sum_{i \in B_k} S_i - \sum_{i \in B_k} S_i^2.$$

Let $v = c^2 \ln(2)/2\alpha n$; this is half of what (b) allows us to lose (note that $v \geq 2 \ln n$ by our choice of parameters). Then,

$$\begin{aligned} \Pr_{\mathcal{P}'}[\neg(b) \mid E_1, \dots, E_{k-1}] &= \Pr_{\mathcal{P}'} \left[\sum_{i \in B_k} \ln(1 + S_i) < -2v \right] \leq \Pr_{\mathcal{P}'} \left[\sum_{i \in B_k} S_i - \sum_{i \in B_k} S_i^2 < -2v \right] \\ &\leq \Pr_{\mathcal{P}'} \left[\sum_{i \in B_k} S_i < -v \right] + \Pr_{\mathcal{P}'} \left[\sum_{i \in B_k} S_i^2 > v \right]. \end{aligned} \quad (10)$$

First we bound the second term of the righthand side of Eq. (10). Corollary 7 implies, both under \mathcal{P} and \mathcal{P}' :

$$\mathbb{E} \left[\sum_{i \in B_k} S_i^2 \right] = \sum_{i \in B_k} 4\mathbb{E}[\beta_i^2] \leq c \cdot 4\gamma \left(\frac{c}{n}\right)^2 = \frac{4\gamma c^3}{n^2},$$

By Markov's inequality

$$\Pr_{\mathcal{P}'} \left[\sum_{i \in B_k} S_i^2 > v \right] \leq \frac{4\gamma c^3}{vn^2} = \frac{8\alpha\gamma c}{\ln 2 n} \leq \frac{\eta c}{4\alpha n},$$

where the equality follows from our value of c , and the last inequality follows simply from our upper bound on α .

Now we bound the first term on the right of Eq. (10). By Lemma 8 (with $t = v \geq 2 \ln n$),

$$\Pr_{\mathcal{P}'} \left[\sum_{i \in B_k} S_i < -v \right] \leq e^{-v/2} + \Pr_{\mathcal{P}'} \left[\sum_{i \in B_k} S_i^2 \geq v \right] \leq \frac{1}{n} + \frac{\eta c}{4\alpha n} \leq \frac{\eta c}{2\alpha n}.$$

Putting everything together:

$$\begin{aligned} \Pr_{\mathcal{P}}[\neg E_k \mid E_1, \dots, E_{k-1}] &\leq \Pr_{\mathcal{P}}[\neg(a) \mid E_1, \dots, E_{k-1}] + \Pr_{\mathcal{P}}[\neg(b) \mid E_1, \dots, E_{k-1}, (a)] \\ &= \Pr_{\mathcal{P}}[\neg(a) \mid E_1, \dots, E_{k-1}] + \Pr_{\mathcal{P}'}[\neg(b) \mid E_1, \dots, E_{k-1}] \\ &\leq \frac{\eta c}{4\alpha n} + \Pr_{\mathcal{P}'} \left[\sum_{i \in B_k} S_i < -v \right] + \Pr_{\mathcal{P}'} \left[\sum_{i \in B_k} S_i^2 > v \right] \\ &\leq \frac{\eta c}{4\alpha n} + \frac{\eta c}{2\alpha n} + \frac{\eta c}{4\alpha n} = \frac{\eta c}{\alpha n}. \end{aligned}$$

This concludes the proof of Claim 10. □

This concludes the proof of Lemma 9. □

5.2 Second pass: Z is close to uniform

We now prove the main result about α PM.

Theorem 11. *Let $\eta \in [0, 1]$ and $\alpha \leq \sqrt{\eta/256\gamma \log n}$. If the communication is*

$$c \leq \sqrt{\eta^3 n / 2^{14} \ln(64/\eta) \alpha \gamma} = O(\sqrt{\eta^3 n / \alpha})$$

bits, then

$$\mathbb{E}_M[d(Z, U_{\alpha n} \mid M)] \leq \eta.$$

Proof. Let $S_\ell = -(-1)^{z_\ell} 2\beta_\ell$ as before and $v = \eta^2/32 \ln(64/\eta)$. Call a pair M, z “good” if the following three things hold for it: (1) $|S_\ell| \leq 1/2$ for all ℓ , (2) $\sum_{\ell=1}^{\alpha n} S_\ell^2 \leq v$, and (3) $|\sum_{\ell=1}^{\alpha n} S_\ell| \leq \eta/4$. Call the pair M, z “bad” otherwise.

Letting $\#M$ denote the number of possible α -matchings M , we rewrite the expected total variation distance over all matchings as:

$$\begin{aligned} \mathbb{E}_M [d(Z, U_{\alpha n} \mid M)] &= \frac{1}{2\#M} \sum_{M, z} \left| \Pr_{x \in A} [Z = z \mid M] - 2^{-\alpha n} \right| \\ &= \frac{1}{2\#M} \sum_{\text{good } M, z} \left| \Pr_{x \in A} [Z = z \mid M] - 2^{-\alpha n} \right| + \frac{1}{2\#M} \sum_{\text{bad } M, z} \left| \Pr_{x \in A} [Z = z \mid M] - 2^{-\alpha n} \right| \end{aligned}$$

Let \mathcal{P} be the uniform distribution on M, z . We start by bounding the probability that M, z is a bad pair over \mathcal{P} . Since $c \leq \sqrt{\eta^3 n / 2^{14} \ln(64/\eta) \alpha \gamma} \leq \sqrt{(\eta/128)n/64\alpha\gamma}$, we can apply Lemma 9 with value $\eta/128$ for η . Let B denote the bad event that at least one A_ℓ is too small and C the bad event that at least one S_ℓ has absolute value larger than $1/2$. Then by Lemma 9

$$\Pr_{\mathcal{P}}[B] \leq \eta/128 \quad \text{and} \quad \Pr_{\mathcal{P}}[C] \leq \eta/128$$

From Corollary 7 we have

$$\mathbb{E}_{\mathcal{P}} \left[\sum_{\ell=1}^{\alpha n} S_\ell^2 \mid \neg B \right] \leq 4\alpha n \gamma (c/n)^2 \leq \eta v / 128.$$

By Markov’s inequality

$$\Pr_{\mathcal{P}} \left[\sum_{\ell=1}^{\alpha n} S_\ell^2 \geq v \mid \neg B \right] \leq \eta/128$$

and

$$\Pr_{\mathcal{P}} \left[\sum_{\ell=1}^{\alpha n} S_\ell^2 \geq v \right] \leq \Pr_{\mathcal{P}}[B] + \Pr_{\mathcal{P}} \left[\sum_{\ell=1}^{\alpha n} S_\ell^2 \geq v \mid \neg B \right] \leq \eta/64.$$

We can now apply Lemma 8 with $t = \eta/4$ and our $v = \eta^2/32 \ln(64/\eta)$ to show that $\sum_{\ell=1}^{\alpha n} S_\ell$ is usually small (note that $t^2/2v = \ln(64/\eta)$):

$$\Pr_{\mathcal{P}} \left[\left| \sum_{\ell=1}^{\alpha n} S_\ell \right| > \eta/4 \right] \leq 2 \left(e^{-t^2/2v} + \Pr_{\mathcal{P}} \left[\sum_{\ell=1}^{\alpha n} S_\ell^2 \geq v \right] \right) \leq 2(\eta/64 + \eta/128) = 6\eta/128.$$

Applying the union bound, we see that the probability that M, z is a bad pair is at most $(\eta/128 + \eta/64 + 6\eta/128) < \eta/12$.

For any pair M, z we have

$$\left| \Pr_{x \in A} [Z = z \mid M] - 2^{-\alpha n} \right| = \left| \prod_{\ell=1}^{\alpha n} (1/2 - (-1)^{z_\ell} \beta_\ell) - 2^{-\alpha n} \right| = 2^{-\alpha n} \left| \prod_{\ell=1}^{\alpha n} (1 + S_\ell) - 1 \right|.$$

Next, we show that $|\prod_{\ell=1}^{\alpha n} (1 + S_\ell) - 1| \leq \eta/2$ for good pairs M, z . First,

$$\prod_{\ell=1}^{\alpha n} (1 + S_\ell) \leq e^{\sum_{\ell=1}^{\alpha n} S_\ell} \leq e^{\eta/4} \leq 1 + \eta/2.$$

Second, since $|S_\ell| \leq 1/2$ for all ℓ , we have

$$\sum_{\ell=1}^{\alpha n} \ln(1 + S_\ell) \geq \sum_{\ell=1}^{\alpha n} S_\ell - \sum_{\ell=1}^{\alpha n} S_\ell^2 \geq -\eta/4 - v$$

and so

$$\prod_{\ell=1}^{\alpha n} (1 + S_\ell) \geq e^{-\eta/4 - v} \geq 1 - \eta/4 - v \geq 1 - \eta/2.$$

Hence, for good pairs M, z

$$\frac{1}{\#M} \sum_{\text{good } M, z} \left| \Pr_{x \in A} [Z = z | M] - 2^{-\alpha n} \right| \leq \frac{1}{\#M} \sum_{\text{good } M, z} \frac{\eta}{2} \cdot 2^{-\alpha n} \leq \eta/2.$$

Moreover, using also that the probability that M, z is a good pair is at least $1 - \eta/12$, we have

$$\frac{1}{\#M} \sum_{\text{good } M, z} \Pr_{x \in A} [Z = z | M] \geq (1 - \eta/12)(1 - \eta/2) \geq 1 - 7\eta/12$$

and therefore

$$\frac{1}{\#M} \sum_{\text{bad } M, z} \left| \Pr_{x \in A} [Z = z | M] - 2^{-\alpha n} \right| \leq \frac{1}{\#M} \sum_{\text{bad } M, z} \left(\Pr_{x \in A} [Z = z | M] + 2^{-\alpha n} \right) \leq \frac{7\eta}{12} + \frac{\eta}{12} = \frac{2\eta}{3}.$$

Now we can finally bound the expected total variation distance over all matchings:

$$\mathbb{E}_M [d(Z, U_{\alpha n} | M)] = \frac{1}{2\#M} \sum_{M, z} \left| \Pr_{x \in A} [Z = z | M] - 2^{-\alpha n} \right| \leq \frac{1}{2} (\eta/2 + 2\eta/3) \leq \eta.$$

□

5.3 Consequences

As explained in Section 3, applying Theorem 11 with small constant η gives the lower bound on the classical communication complexity of α PM. Combining that classical lower bound with the bounds mentioned in Section 2, we obtain the separation stated in the introduction:

Theorem 2. *Let $\alpha \in [0, O(1/\sqrt{\log n})]$. The classical bounded-error one-way communication complexity of the α -Partial Matching problem is $R_\varepsilon^1(\alpha\text{PM}) = \Theta(\sqrt{n/\alpha})$, while the quantum bounded-error one-way complexity is $Q_\varepsilon^1(\alpha\text{PM}) = O(\log(n)/\alpha)$.*

As also explained in Section 3, we can reinterpret—in the bounded storage model—the fact that from Bob’s perspective, Z is close to uniformly distributed if the message is short. The classical lower bound can be reinterpreted as follows: if the $2n$ -bit string X is uniformly distributed and the adversary has $\sigma\sqrt{\eta^3 n/\alpha}$ classical bits of memory available (for sufficiently small positive constant σ), then even given his memory and even given Y , his distribution on the αn -bit string $Z(X, Y)$ is η -close to uniform. In contrast, the quantum upper bound from Section 3 shows that an adversary with $O(\log(n)/\alpha)$ qubits of memory gives the adversary one bit of the string $Z(X, Y)$. Hence the distance between Z and the uniform distribution is at least $1/2$. Using k times as much memory gives the adversary roughly k bits of $Z(X, Y)$, and the resulting distribution will be $1 - 2^{-k}$ far from uniform. This gives the theorem stated in the introduction:

Theorem 3. Let $\eta \in [0, 1]$ and $\alpha \in [0, O(\sqrt{\eta/\log n})]$. The extracted αn -bit string in the bounded storage protocol derived from the α -Partial Matching problem is η -secure against a classical adversary with memory bound $O(\sqrt{\eta^3 n/\alpha})$, while it is at most $(1 - 2^{-k})$ -secure against a quantum adversary with memory bound $O(k \log(n)/\alpha)$.

So far, we have proved that if the $2n$ -bit string X is uniformly distributed over a set A with $|A| \geq 2^{2n-c-1}$ (i.e., a flat distribution on A), then $Z(X, Y)$ is close to uniform even knowing Y , where Y is the uniformly chosen matching. In order to conclude the result about extractors mentioned in Section 1.2.2, we need to prove the same result in the more general situation when X has min-entropy greater than $2n - c - 1$ (instead of just being uniform on a set of size at least 2^{2n-c-1}). However, a result by Chor and Goldreich [CG88, Lemma 5] based on the fact that any distribution can be thought of as a convex combination of flat distributions, shows that the second statement follows from the first: flat distributions are the “worst distributions” for extractors.

6 Future work

We state some open problems:

- Our lower bound on α PM uses $\alpha = O(1/\sqrt{\log n})$. We would like to make it work also for constant α . Recently we learned of a simplified proof of the lower bound on α PM by Oded Regev, who used the Fourier methods from our proof of the lower bound on NPM together with the Bonami-Beckner inequality, which may resolve this.
- Our bounded storage example is not very natural as the length of the extracted key (αn bits) is shorter than the length of the initial key ($\approx \alpha n \log n$ bits). Find a constructive example with a much shorter initial key.

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