# Exponential Separation of Information and Communication

Anat Ganor* Gillat Kol† Ran Raz‡

**Abstract**

We show an exponential gap between communication complexity and information complexity, by giving an explicit example for a communication task (relation), with information complexity $\leq O(k)$, and distributional communication complexity $\geq 2^k$. This shows that a communication protocol cannot always be compressed to its internal information. By a result of Braverman [Bra12b], our gap is the largest possible. By a result of Braverman and Rao [BR11], our example shows a gap between communication complexity and amortized communication complexity, implying that a tight direct sum result for distributional communication complexity cannot hold.

## 1 Introduction

Communication complexity is a central model in complexity theory that has been extensively studied in numerous works. In the two player distributional model, each player gets an input, where the inputs are sampled from a joint distribution that is known to both players. The players' goal is to solve a communication task that depends on both inputs. The players can use both common and private random strings and are allowed to err with some small probability. The players communicate in rounds, where in each round one of the players sends a message to the other player. The communication complexity of a protocol is the total number of bits communicated by the two players. The communication complexity of a communication task is the minimal number of bits that the players need to communicate in order to solve the task with high probability, where the minimum is taken over all protocols. For excellent surveys on communication complexity see [KN97, LS09].

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†Institute for Advanced Study, Princeton, NJ. Research at the IAS supported by The Fund For Math and the Weizmann Institute of Science National Postdoctoral Award Program for Advancing Women in Science.

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The information complexity model, first introduced by [CSWY01, BYJKS04, BBCR10], studies the amount of information that the players need to reveal about their inputs in order to solve a communication task. The model was motivated by fundamental information theoretical questions of compressing communication, as well as by fascinating relations to communication complexity, and in particular to the direct sum problem in communication complexity, a problem that has a rich history, and has been studied in many works and various settings [FKNN95, CSWY01, JRS03, HJMR07, BBCR10, Kla10, Jai11, JPY12, BRWY12, BRWY13] (and many other works). In this paper we will mainly be interested in internal information complexity (a.k.a, information complexity and information cost). Roughly speaking, the internal information complexity of a protocol is the number of information bits that the players learn about each other's input, when running the protocol. The information complexity of a communication task is the minimal number of information bits that the players learn about each other's input when solving the task, where the minimum is taken over all protocols.

Many recent works focused on the problem of compressing interactive communication protocols. Given a communication protocol with small information complexity, can the protocol be compressed so that the total number of bits communicated by the protocol is also small? There are several beautiful known results, showing how to compress communication protocols in several cases. Barak, Braverman, Chen and Rao showed how to compress any protocol with information complexity $k$ and communication complexity $c$, to a protocol with communication complexity $\tilde{O}(\sqrt{ck})$ in the general case, and $\tilde{O}(k)$ in the case where the underlying distribution is a product distribution [BBCR10]. Braverman and Rao showed how to compress any one round (or small number of rounds) protocol with information complexity $k$ to a protocol with communication complexity $O(k)$ [BR11]. Braverman showed how to compress any protocol with information complexity $k$ to a protocol with communication complexity $2^{O(k)}$ [Bra12b] (see also [BW12, KLL+12]). This last protocol is the most related to our work, as it gives a compression result that works in the general case and doesn't depend at all on the communication complexity of the original protocol. Braverman also described a communication complexity task that has information complexity $O(k)$ and no known communication protocol with communication complexity smaller than $2^k$ [Bra13]. However, there is no known lower bound on the communication complexity of that problem.

Another line of works shows that many of the known general techniques for proving lower bounds for randomized communication complexity also give lower bounds for information complexity [Bra12b, BW12, KLL+12].

In this work we show the first gap between information complexity and communication complexity of a communication task. We give an explicit example for a communication task (a relation), called the *bursting noise game*, parameterized by $k \in \mathbb{N}$ and played with an input distribution $\mu$. We prove that the information complexity of the game is $O(k)$, while any communication protocol for solving this game, with communication complexity at most $2^k$, almost always errs. By the above mentioned compression protocol of Braverman [Bra12b],

our result gives the largest possible gap between information complexity and communication complexity.

**Theorem 1** (**Communication Lower Bound**). *Every randomized protocol (with shared randomness) for the bursting noise game with parameter $k$, that has communication complexity at most $2^k$, errs with probability $\epsilon \geq 1 - 2^{-\Omega(k)}$ (over the input distribution $\mu$).*

**Theorem 2** (**Information Upper Bound**). *There exists a randomized protocol for the bursting noise game with parameter $k$, that has information cost $O(k)$ and errs with probability $\epsilon \leq 2^{-\Omega(k)}$ (over the input distribution $\mu$).*

We note that both the inputs and the outputs in our bursting noise game example are very long. Namely, the input length is triple exponential in $k$, and the output length is double exponential. The protocol that achieves information complexity $O(k)$ has communication complexity double exponential in $k$.

As mentioned above, information complexity is also related to the direct sum problem in communication complexity. Braverman and Rao showed that information complexity is equal to the amortized communication complexity, that is, the limit of the communication complexity needed to solve $n$ tasks of the same type, divided by $n$ [BR11] (see also [Bra12a, Bra12b, Bra13]). Our result therefore shows a gap between distributional communication complexity and amortized distributional communication complexity, proving that tight direct sum results for the communication complexity of relations cannot hold.

**Organization.** The paper is organized as follows. In Section 2 we define the bursting noise game. Section 3 gives an overview of our main result, the lower bound for the communication complexity of the bursting noise game (Theorem 1). In Section 4 we give general definitions and preliminaries. In Section 5 we prove the graph correlation lemma, a central tool that we will use in the lower bound proof. In Section 6 we prove the communication complexity lower bound (Theorem 1). Section 7 gives a general tool that can be used to upper bound the information cost of a protocol, using the notion of a divergence cost of a tree. In Section 8 we give a protocol for the bursting noise game with low information cost, thus proving the upper bound required by Theorem 2. The appendix contains information theoretic lemmas that are used by the lower bound proof.

# 2 Bursting Noise Games

The *bursting noise game* is a communication game between two parties, called the *first player* and the *second player*. The game is specified by a parameter $k \in \mathbb{N}$, where $k > 2^{100}$. We set $c = 2^{4^k}$ and $w = 2^{100}k$.

The game is played on the binary tree $\mathcal{T}$ with $c \cdot w$ layers (the root is in layer 1 and the leaves are in layer $c \cdot w$), with edges directed from the root to the leaves. Denote the vertex set of $\mathcal{T}$ by $V$. Each player gets as input a bit for every vertex in the tree. Let $x$ be the input

given to the first player, and $y$ be the input given to the second player, where $x, y \in \{0,1\}^V$. For a vertex $v \in V$, we denote by $x_v$ and $y_v$ the bits in $x$ and $y$ associated with $v$. The input pair $(x, y)$ is selected according to a joint distribution $\mu$ on $\{0,1\}^V \times \{0,1\}^V$, defined below.

Denote by $\text{Even}(\mathcal{T}) \subseteq V$ the set of non-leaf vertices in an even layer of $\mathcal{T}$ and by $\text{Odd}(\mathcal{T}) \subseteq V$ the set of non-leaf vertices in an odd layer of $\mathcal{T}$. We think of the vertices in $\text{Odd}(\mathcal{T})$ as "owned" by the first player and the vertices in $\text{Even}(\mathcal{T})$ as "owned" by the second player. Let $v \in V$ be a non-leaf vertex. Let $v_0$ be the left child of $v$ and $v_1$ be the right child of $v$. Let $b \in \{0,1\}$. We say that $v_b$ is the *correct child* of $v$ with respect to $x, y$, if either the first player owns $v$ and $x_v = b$, or the second player owns $v$ and $y_v = b$.

We think of the $c \cdot w$ layers of the tree $\mathcal{T}$ as partitioned into $c$ multi-layers, each consisting of $w$ consecutive layers (e.g., the first multi-layer consists of layers 1 to $w$). We denote by $i^*$ the first layer of the $i^{th}$ multi-layer, that is, $i^* = (i-1)w + 1$.

For $s \leq t \in \mathbb{N}$, denote by $[s,t]$ the set $\{s, \ldots, t\}$ and by $[t]$ the set $\{1, \ldots, t\}$. Let $i \in [c]$ be a multi-layer. Denote $s = i^*$ and $t = s + w - 1 = (i+1)^* - 1$. Let $t' \in [(i+1)^*, cw]$, and let $v \in V$ be a vertex in layer $t'$ of $\mathcal{T}$. For $j \in [s, t+1]$, let $v_j$ be $v$'s ancestor in layer $j$. We say that $v$ is *typical* with respect to $i, x, y$, if the followings hold:

1. For at least 0.8-fraction of the indices $j \in [s,t] \cap \text{Odd}(\mathcal{T})$, the vertex $v_{j+1}$ is the correct child of $v_j$ with respect to $x, y$.

2. For at least 0.8-fraction of the indices $j \in [s,t] \cap \text{Even}(\mathcal{T})$, the vertex $v_{j+1}$ is the correct child of $v_j$ with respect to $x, y$.

Observe that in order to decide whether $v$ is typical with respect to $i, x, y$, it suffices to know the bits that $x, y$ assign to the vertices $v_s, \ldots, v_t$. When $x, y$ are clear from the context, we omit $x, y$ and say that $v$ is typical with respect to multi-layer $i$.

We next define the distribution $\mu$ on $\{0,1\}^V \times \{0,1\}^V$ by an algorithm for sampling an input pair $(x, y)$ (Algorithm 1 below). In the algorithm, when we say "set $v$ to be non-noisy", we mean "select $x_v \in \{0,1\}$ uniformly at random and set $y_v = x_v$". By "set $v$ to be noisy", we mean "select $x_v \in \{0,1\}$ and $y_v \in \{0,1\}$ independently and uniformly at random". Figure 1 illustrates Algorithm 1.

The players' mutual goal is to output the same leaf $v \in V$, where $v$ is typical with respect to $i, x, y$ (that is, $v$ is typical with respect to the noisy multi-layer; see Algorithm 1).

For $i \in [c]$, we denote by $\mu_i$ the distribution $\mu$ conditioned on the event that the noisy multi-layer selected by Step 1 of the algorithm defining $\mu$, is $i$. Note that $\mu = \frac{1}{c} \sum_{i \in [c]} \mu_i$.

**Remark.** *Observe that it is not always possible to deduce $i$ (i.e., the index of the noisy multi-layer used to construct the pair $(x, y)$) from the pair $(x, y)$. Therefore, the bursting noise game does not induce a relation. Nevertheless, with extremely high probability, the first multi-layer on which $x$ and $y$ disagree is $i$. Thus, the game can be easily converted to a relation, by omitting the rare inputs $(x, y)$ that agree on multi-layer $i$. Note that since the statistical distance between the two distributions is negligible, both our upper bound and lower*

**Algorithm 1** Sample $(x, y)$ according to $\mu$

1. Randomly select $i \in [c]$ (the noisy multi-layer).
2. Set every vertex in multi-layer $i$ (layers $[i^*, i^* + w - 1]$) to be noisy.
3. If $i < c$: Let $L$ be the set of all non-typical vertices in layer $i^* + w = (i + 1)^*$ with respect to $i, x, y$ (note that $x, y$ were already defined on layers $[i^*, i^* + w - 1]$, and therefore the typical vertices are defined). For every $v \in L$, set all the vertices in the subtree rooted at $v$ to be noisy.
4. Set all unset vertices in $V$ to be non-noisy.

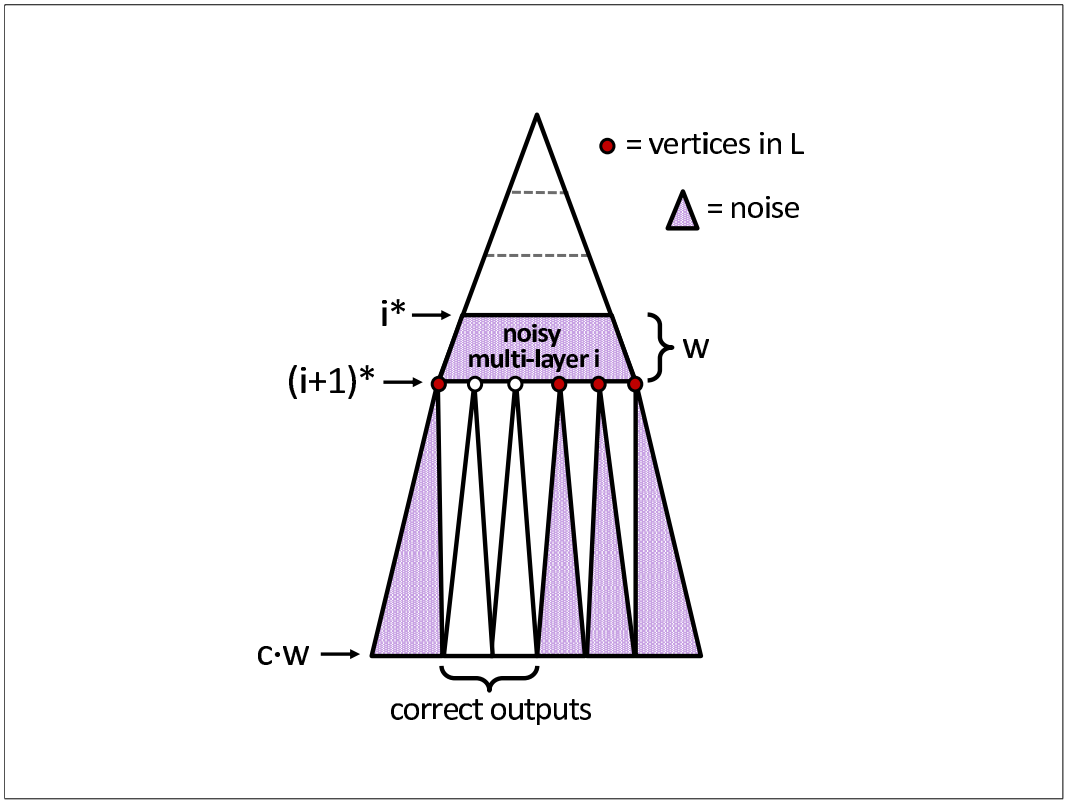


Figure 1: Illustration of Algorithm 1

*bound trivially apply to the new game as well. For that reason, it will be helpful to think of the supports of the different $\mu_i$'s as if they were pairwise disjoint.*

**Remark.** *Observe that $c$ is set to be double exponential in $k$. If $c$ were set to be just exponential in $k$, a simple binary search algorithm would have been able to find the location of the noisy multi-layer, and thus solve the bursting noise game with communication complexity polynomial in $k$.*

**The protocol with low information cost.** Consider the following protocol $\pi'$ for the bursting noise game. Starting from the root until reaching a leaf, at every vertex $v$, if the first player owns $v$, she sends the bit $x_v$ with probability 0.9, and the bit $1-x_v$ with probability 0.1. Similarly, if the second player owns $v$, she sends the bit $y_v$ with probability 0.9, and the bit $1 - y_v$ with probability 0.1. Both players continue to the child of $v$ that is indicated by the communicated bit. When they reach a leaf they output that leaf. By the Chernoff bound, the probability that the players output a leaf that is not typical with respect to the noisy multi-layer is at most $2^{-\Omega(w)}$. That is, the error probability of $\pi'$ is exponentially small in $k$.

It can be shown that if the protocol $\pi'$ does not reach a vertex in $L$ (a non-typical vertex with respect to the noisy multi-layer), then it reveals a small amount of information. Intuitively, this follows since in this case, the expected number of vertices reached by the protocol, on which the players' inputs disagree, is $O(k)$ (the disagreement is only on vertices in the noisy multi-layer). However, with exponentially small probability in $k$, the protocol $\pi'$ does reach a vertex in $L$. In this case, the information revealed by the protocol may be double exponential in $k$ (as $c = 2^{4^k}$), making the information cost of $\pi'$ too large.

For this reason, we consider a variant of $\pi'$, called $\pi$. Informally speaking, the protocol $\pi$ operates like $\pi'$ but aborts if too much information about the inputs is revealed. Specifically, a player decides to abort if the bits that she receives differ from the corresponding bits in her input too many times. In Section 8, we formally define $\pi$ and show that its information cost is $O(k)$.

# 3 Overview of the Lower Bound Proof

### Rectangle Partition

We will describe the proof of the lower bound for the communication complexity of the bursting noise game. We fix the random strings for the protocol so that we have a deterministic protocol. We show that if the protocol communicates at most $2^k$ bits, it errs with probability $1 - 2^{-\Omega(k)}$ on inputs sampled according to $\mu$. We will show that for almost all $i \in [c]$, the protocol errs with probability $1 - 2^{-\Omega(k)}$ on inputs sampled according to $\mu_i$, that is, the distribution $\mu$ conditioned on the event that the noisy multi-layer selected by Step 1 of Algorithm 1 defining $\mu$, is $i$. Note that the distribution $\mu_i$ is uniformly distributed

over $\text{supp}(\mu_i)$, and that for every pair of inputs $(x, y) \in \text{supp}(\mu_i)$, the projection of $x$ and $y$ on the first $i-1$ multi-layers is the same.

As mentioned above, it will be helpful to think of the supports of the different $\mu_i$'s as if they were pairwise disjoint (this property holds if we remove a $\mu_i$-negligible set of inputs from the support of each $\mu_i$).

Let $\{R^1, \ldots, R^m\}$ be the rectangle partition induced by the protocol, where $R^t = A^t \times B^t$, and $m \leq 2^{2^k}$. For $i \in [c]$ and an assignment $z$ to the first $i-1$ multi-layers, we denote by $R^{t,z} = A^{t,z} \times B^{t,z}$, the rectangle of all pairs of inputs $(x, y) \in R^t$, such that the projection of both $x, y$ on the first $i-1$ multi-layers is equal to $z$. Let $X^{t,z}$ be a random variable uniformly distributed over $A^{t,z}$. Let $Y^{t,z}$ be a random variable uniformly distributed over $B^{t,z}$. We denote by $X_i^{t,z}, Y_i^{t,z}$ the projections of $X^{t,z}, Y^{t,z}$, respectively, on multi-layer $i$.

For fixed $i, z$, we define $\rho^{i,z}$ to be a probability distribution that selects a rectangle in $\{R^{1,z}, \ldots, R^{m,z}\}$ according to its relative size. That is, $\rho^{i,z}$ is defined as follows: Randomly select $x, y$, such that the projection of both $x$ and $y$ on the first $i-1$ multi-layers is $z$. Select $t$ to be the index of the unique rectangle $R^{t,z}$ containing $(x, y)$.

### Bounding the Information on the Noisy Multi-Layer

The main intuition of the proof is that since $c$ is significantly larger than $2^k$, the protocol cannot make progress on all multi-layers $i \in [c]$ simultaneously. We first show that for a random $i \in [c]$, a random $z$, and a random rectangle $R^{t,z}$, chosen according to $\rho^{i,z}$, very little information is known about $X_i^{t,z}$ and $Y_i^{t,z}$.

Formally, we prove in Lemma 11 that

$$\mathop{\mathbf{E}}_{i} \mathop{\mathbf{E}}_{z} \mathop{\mathbf{E}}_{t \leftarrow \rho^{i,z}} \left[\mathbf{I}\left(X_i^{t,z}\right)\right] \leq \frac{m}{c}, \tag{1}$$

and similarly,

$$\mathop{\mathbf{E}}_{i} \mathop{\mathbf{E}}_{z} \mathop{\mathbf{E}}_{t \leftarrow \rho^{i,z}} \left[\mathbf{I}\left(Y_i^{t,z}\right)\right] \leq \frac{m}{c}, \tag{2}$$

where we denote by $\mathbf{I}(Z) := \log(|\Omega|) - \mathbf{H}(Z)$ the information known about a random variable $Z$, where $\Omega$ is the space that $Z$ is defined over.

The proof of Lemma 11 doesn't follow by a trivial application of super-additivity of information. That's because choosing $i, z$ at random and $t$ according to $\rho^{i,z}$ and then choosing a random variable $X$ to be uniformly distributed on $A^{t,z}$, gives a random variable $X$ with distribution that may be very far from uniform. Moreover, the probability that $X$ is in the set $A^t$, associated with a rectangle $R^t$, may be very far from the probability that a uniformly distributed input is in $A^t$. Nevertheless, we are still able to prove Lemma 11, using the fact that we have a bound of $m$ on the total number of times that an input $x$ appears in the cover $\{A^1, \ldots, A^m\}$.

We fix $\gamma = 2^{-k/4}$, and we fix $i, z, t$, such that,

1. $\mathbf{I}\left(X_i^{t,z}\right) \leq \frac{1}{\gamma} \cdot \frac{m}{c}$

2. $\mathbf{I}\left(Y_i^{t,z}\right) \leq \frac{1}{\gamma} \cdot \frac{m}{c}$

3. The rectangle $R^{t,z}$ is not too small.

By Equation (1) and Equation (2), and by Markov's inequality, we know that when we choose $i, z$ uniformly at random, and $t$ according to $\rho^{i,z}$, the triplet $(i, z, t)$ satisfies all three conditions with high probability. Therefore, we ignore triplets $(i, z, t)$ that do not satisfy all three conditions.

### Unique Answer Rectangles

In the rectangle $R^{t,z}$, the answer of each of the two players in the protocol may not be unique, as the answer of each player may also depend on the input that she gets. Nevertheless, in Lemma 12, using the fact that if the two players answer differently then the protocol errs, we are able to subdivide the rectangle $R^{t,z}$ into $\text{poly}(1/\gamma)$ sub-rectangles $R^{t,s,z}$, such that in each rectangle $R^{t,s,z}$ the answer is unique, except for a bad set of rectangles whose total size is negligible compared to the size of $R^{t,z}$. When subdividing $R^{t,z}$, we also need to change the answers given by the two players on each rectangle, but we are able to do that without adding errors to the protocol.

We ignore rectangles $R^{t,s,z}$ where the answer of the protocol is not unique, as their total size is small, and only consider rectangles $R^{t,s,z} = A^{t,s,z} \times B^{t,s,z}$ where the answer is unique. Let $X^{t,s,z}$ be a random variable uniformly distributed over $A^{t,s,z}$. Let $Y^{t,s,z}$ be a random variable uniformly distributed over $B^{t,s,z}$. For the rectangles $R^{t,s,z}$ we no longer have the strong bounds $\mathbf{I}\left(X_i^{t,z}\right) \leq \frac{1}{\gamma} \cdot \frac{m}{c}$, and $\mathbf{I}\left(Y_i^{t,z}\right) \leq \frac{1}{\gamma} \cdot \frac{m}{c}$, but rather the weaker bounds

$$\mathbf{I}\left(X_i^{t,s,z}\right) \leq O\left(\log\left(1/\gamma\right)\right),$$

and

$$\mathbf{I}\left(Y_i^{t,s,z}\right) \leq O\left(\log\left(1/\gamma\right)\right).$$

### How the Proof Works

Fix $i, z, t, s$. In the rectangle $R^{t,s,z}$ the answer is unique, denote that answer by $\omega^{t,s,z}$. We define $\Lambda^{t,s,z}$ to be the set of input pairs $(x, y) \in \text{supp}(\mu_i)$, such that $\omega^{t,s,z}$ is not a correct answer for the input $(x, y)$. Let $P_i$ be the probability for a uniformly distributed pair of inputs $(x, y)$, that have the same projection on the first $i-1$ multi-layers, to be in $\text{supp}(\mu_i)$. In Lemma 14, we prove that

$$\Pr\left[(X^{t,s,z}, Y^{t,s,z}) \in \Lambda^{t,s,z}\right] \geq \left(1 - 2^{-\Omega(k)}\right) P_i. \tag{3}$$

Summing over all possibilities for $t, s, z$, this implies, for almost all $i \in [c]$, that the protocol errs on $\mu_i$ with probability $1 - 2^{-\Omega(k)}$, which concludes the proof.

In what follows, we outline the proof of Equation (3).

### The Graph $G$

We define the complete bipartite graph $G = (U \cup W, E)$, where $U = W$ is the set of all possible assignments for multi-layer $i$ (for one player), and $E = U \times W$.

Let $M$ be the number of vertices in layer $(i+1)^*$ of the tree $\mathcal{T}$. We identify the set $[M]$ with the set of vertices in layer $(i+1)^*$. Let $u \in U, w \in W$. We define $T(u,w) \subset [M]$ to be the set of all vertices in layer $(i+1)^*$ that are set to be non-noisy for inputs $u, w$, by Algorithm 1 defining $\mu$, when the noisy multi-layer is $i$. Observe that $u$ and $w$ determine for every vertex in layer $(i+1)^*$ if it is noisy or not. Note that by a symmetry argument, $T(u,w)$ is of the same size $T$ for every $u, w$.

Let $\mathcal{E}^{t,s,z} \subseteq E$ be the set of all $(u,w) \in E$ for which the output $\omega^{t,s,z}$ is correct for inputs $(x,y) \in \text{supp}(\mu_i)$, where $x_i = u$ and $y_i = w$. Note that if the noisy multi-layer is $i$, then $u$ and $w$ determine the correctness of $\omega^{t,s,z}$. It holds that

$$|\mathcal{E}^{t,s,z}| \leq 2^{-20k}|E|,$$

as for any fixed $u$ and every $v \in [M]$, at most a fraction of $2^{-20k}$ of the sets $\{T(u,w)\}_{(u,w)\in E}$ contain $v$, and the output $\omega^{t,s,z}$ is correct only if it has an ancestor in $T(u,w)$.

Let $\Sigma$ be the set of all possible boolean assignments to the vertices of a subtree of $\mathcal{T}$ rooted at layer $(i+1)^*$.

For $u \in U$, we define the random variable $X^u$, over the domain $\Sigma^{[M]}$, to be the conditional variable $(X^{t,s,z}_{>i} | X^{t,s,z}_i = u)$, that is, $X^u$ has the distribution of $X^{t,s,z}_{>i}$ conditioned on the event $X^{t,s,z}_i = u$, where $X^{t,s,z}_{>i}$ denotes the projection of $X^{t,s,z}$ to all multi-layers after multi-layer $i$. Similarly, for $w \in W$, we define the random variable $Y^w$, over the domain $\Sigma^{[M]}$, to be $(Y^{t,s,z}_{>i} | Y^{t,s,z}_i = w)$, that is, $Y^w$ has the distribution of $Y^{t,s,z}_{>i}$ conditioned on the event $Y^{t,s,z}_i = w$.

### Application of the Graph Correlation Lemma

By the definition of the distribution $\mu_i$, the left hand side of Equation (3) is equal to

$$\sum_{(u,w)\in E\setminus\mathcal{E}^{t,s,z}} \Pr\left[X^{t,s,z}_i = u\right] \cdot \Pr\left[Y^{t,s,z}_i = w\right] \cdot \Pr\left[X^u_{T(u,w)} = Y^w_{T(u,w)}\right], \tag{4}$$

where $X^u_{T(u,w)}$ and $Y^w_{T(u,w)}$ are the projections of $X^u, Y^w$, respectively, to coordinates in $T(u,w)$. This is true because a pair $(x,y)$ is in $\text{supp}(\mu_i)$ if and only if $x, y$ agree on all the subtrees rooted at vertices in layer $(i+1)^*$ that are set to be non-noisy for inputs $x_i, y_i$, by Algorithm 1 defining $\mu$, when the noisy multi-layer is $i$.

Our graph correlation lemma (Lemma 9), that may be interesting in its own right, gives a general way to bound such expressions by

$$\geq \left(1 - 2^{-\Omega(k)}\right) |\Sigma|^{-T} \sum_{(u,w)\in E\setminus(\mathcal{E}^{t,s,z}\cup\mathcal{D}^{t,s,z})} \Pr\left[X^{t,s,z}_i = u\right] \cdot \Pr\left[Y^{t,s,z}_i = w\right], \tag{5}$$

where $\mathcal{D}^{t,s,z} \subset E$ is a small set, compared to the size of $E$, and $|\Sigma|^{-T}$ is a normalization factor that would have been equal to $\Pr[X^u_{T(u,w)} = Y^w_{T(u,w)}]$ if $X^u, Y^w$ were uniformly distributed (independent) random variables.

Thus, using Lemma 9, we are able to bound the left hand side of Equation (3), which is an expression that depends on the variables $X^{t,s,z}, Y^{t,s,z}$, by the expression in Equation (5) that depends only on the projections of these variables to multi-layer $i$.

We still need to bound from below the expression

$$\sum_{(u,w)\in E\setminus(\mathcal{E}^{t,s,z}\cup\mathcal{D}^{t,s,z})} \Pr\left[X_i^{t,s,z}=u\right]\cdot\Pr\left[Y_i^{t,s,z}=w\right]. \tag{6}$$

Since $\mathcal{E}^{t,s,z} \cup \mathcal{D}^{t,s,z}$ is a small set (compared to the size of $E$), we will first ignore the set $\mathcal{E}^{t,s,z} \cup \mathcal{D}^{t,s,z}$, and observe that

$$\sum_{(u,w)\in E} \Pr\left[X_i^{t,s,z}=u\right]\cdot\Pr\left[Y_i^{t,s,z}=w\right] = \sum_{u\in U}\Pr\left[X_i^{t,s,z}=u\right]\cdot\sum_{w\in W}\Pr\left[Y_i^{t,s,z}=w\right] = 1. \tag{7}$$

It remains to show that

$$\sum_{(u,w)\in\mathcal{E}^{t,s,z}\cup\mathcal{D}^{t,s,z}} \Pr\left[X_i^{t,s,z}=u\right]\cdot\Pr\left[Y_i^{t,s,z}=w\right],$$

is negligible.

### Bounding the Sum over the Bad Sets

In Lemma 15 we use the fact that $R^{t,s,z} \subseteq R^{t,z}$, to bound the last sum by

$$\frac{|R^{t,z}|}{|R^{t,s,z}|}\sum_{(u,w)\in\mathcal{E}^{t,s,z}\cup\mathcal{D}^{t,s,z}} \Pr\left[X_i^{t,z}=u\right]\cdot\Pr\left[Y_i^{t,z}=w\right].$$

Since $\mathbf{I}\left(X_i^{t,z}\right) \le \frac{1}{\gamma}\cdot\frac{m}{c}$, and $\mathbf{I}\left(Y_i^{t,z}\right) \le \frac{1}{\gamma}\cdot\frac{m}{c}$, we know that the distributions of $X_i^{t,z}$ and $Y_i^{t,z}$ are extremely close to uniform, and hence the sum in the last expression is negligible. Using also the fact that $\frac{|R^{t,z}|}{|R^{t,s,z}|} \le \mathrm{poly}(1/\gamma)$, we get that the entire expression is negligible.

A difficulty that we ignored in the discussion so far, is that the graph correlation lemma (Lemma 9) requires random variables $X^u, Y^w$ with bounded information for all $u, w$, while we have variables with bounded information for almost all $u, w$. To fix that, we just replace every $X^u$ or $Y^w$ that has large information, with a uniformly distributed random variable. This works since $G$ is the complete graph.

### Proof of the Graph Correlation Lemma and Shearer's Inequality

To bound expressions such as the expression in Equation (4), we show that if $\Pr[X^u_{T(u,w)} = Y^w_{T(u,w)}]$ is significantly larger than what is obtained by uniformly distributed variables, then either $\mathbf{I}(X^u_{T(u,w)})$ or $\mathbf{I}(Y^w_{T(u,w)})$ are non negligible (or both). We use this to show that for

some $u$ (or some $w$) we have that $\mathbf{I}(X^u)$ (or $\mathbf{I}(Y^w)$) are large, deriving a contradiction.

Our proof relies on a variant of Shearer's inequality [CGFS86, Kah01] that follows easily by Radhakrishnan's beautiful information theoretical proof [Rad03] (see Lemmas 7 and 8 and [MT10]).

# 4 Definitions and Preliminaries

## 4.1 General Notation

Throughout the paper, all logarithms are taken with base 2, and we define $0\log(0) = 0$. For a set $S$, when we write "$x \in_R S$" we mean that $x$ is selected uniformly at random from the set $S$. For a distribution $\tau$, when we write "$x \leftarrow \tau$" we mean that $x$ is selected according to the distribution $\tau$. For $Z$ that is either a random variable taking values in $\{0,1\}^V$ or an element in $\{0,1\}^V$, and a set $T \subseteq V$, we define $Z_T$ to be the projection of $Z$ to $T$.

## 4.2 Information Cost

**Definition 1** (**Information Cost**). *The* information cost *of a protocol $\pi$ over random inputs $(X,Y)$ that are drawn according to a joint distribution $\mu$, is defined as*

$$IC_\mu(\pi) = \mathbf{I}(\Pi; X|Y) + \mathbf{I}(\Pi; Y|X),$$

*where $\Pi$ is a random variable which is the transcript of the protocol $\pi$ with respect to $\mu$. That is, $\Pi$ is the concatenation of all the messages exchanged during the execution of $\pi$. The $\epsilon$ information cost of a computational task $f$ with respect to a distribution $\mu$ is defined as*

$$IC_\mu(f,\epsilon) = \inf_\pi IC_\mu(\pi),$$

*where the infimum ranges over all protocols $\pi$ that solve $f$ with error at most $\epsilon$ on inputs that are sampled according to $\mu$.*

## 4.3 Relative Entropy

**Definition 2** (**Relative Entropy**). *Let $\mu_1, \mu_2 : \Omega \to [0,1]$ be two distributions, where $\Omega$ is discrete (but not necessarily finite). The* relative entropy *between $\mu_1$ and $\mu_2$, denoted $\mathbf{D}(\mu_1 \| \mu_2)$, is defined as*

$$\mathbf{D}(\mu_1 \| \mu_2) = \sum_{x \in \Omega} \mu_1(x) \log\left(\frac{\mu_1(x)}{\mu_2(x)}\right).$$

**Proposition 3.** *Let $\mu_1, \mu_2 : \Omega \to [0,1]$ be two distributions. Then,*

$$\mathbf{D}(\mu_1 \| \mu_2) \geq 0.$$

The following relation is called Pinsker's inequality, and it relates the relative entropy to the $\ell_1$ distance.

**Proposition 4** (**Pinsker's Inequality**). *Let $\mu_1, \mu_2 : \Omega \to [0,1]$ be two distributions. Then,*

$$2\ln(2) \cdot \mathbf{D}(\mu_1 \| \mu_2) \geq \|\mu_1 - \mu_2\|^2,$$

*where*

$$\|\mu_1 - \mu_2\| = \sum_{x \in \Omega} |\mu_1(x) - \mu_2(x)| = 2 \max_{E \subseteq \Omega} \{\mu_1(E) - \mu_2(E)\}.$$

## 4.4 Information

**Definition 3** (**Information**). *Let $\mu : \Omega \to [0,1]$ be a distribution and let $\mathcal{U}$ be the uniform distribution over $\Omega$. The* information *of $\mu$, denoted $\mathbf{I}(\mu)$, is defined by*

$$\mathbf{I}(\mu) = \mathbf{D}(\mu \parallel \mathcal{U}) = \sum_{x \in \mathrm{supp}(\mu)} \mu(x) \log\left(\frac{\mu(x)}{\frac{1}{|\Omega|}}\right) = \sum_{x \in \mathrm{supp}(\mu)} \mu(x) \log\left(|\Omega| \mu(x)\right).$$

*Equivalently,*

$$\mathbf{I}(\mu) = \log(|\Omega|) - \mathbf{H}(\mu),$$

*where $\mathbf{H}(\mu)$ denotes the Shannon entropy of $\mu$.*

*For a random variable $X$ taking values in $\Omega$, with distribution $P_X : \Omega \to [0,1]$, we define $\mathbf{I}(X) = \mathbf{I}(P_X)$.*

**Proposition 5** (**Supper-Additivity of Information**). *Let $X_1, \ldots, X_m$ be $m$ random variables, taking values in $\Omega_1, \ldots, \Omega_m$, respectively. Consider the random variable $(X_1, \ldots, X_m)$, taking values in $\Omega_1 \times \ldots \times \Omega_m$. Then,*

$$\mathbf{I}\left((X_1, \ldots, X_m)\right) \geq \sum_{i \in [m]} \mathbf{I}(X_i).$$

*Proof.* Using the sub-additivity of the Shannon entropy function, we have

$$\begin{aligned} \mathbf{I}\left((X_1, \ldots, X_m)\right) &= \log(|\Omega_1 \times \ldots \times \Omega_m|) - \mathbf{H}(X_1, \ldots, X_m) \\ &\geq \sum_{i \in [m]} \log(|\Omega_i|) - \sum_{i \in [m]} \mathbf{H}(X_i) \\ &= \sum_{i \in [m]} \left(\log(|\Omega_i|) - \mathbf{H}(X_i)\right) = \sum_{i \in [m]} \mathbf{I}(X_i). \end{aligned}$$

□

## 4.5 Shearer-Like Inequality for Information

The following version of Shearer's inequality [CGFS86, Kah01] is due to [Rad03].

**Lemma 6** (**Shearer's Inequality**). *Let $X_1, \ldots, X_M$ be $M$ random variables. Let $X = (X_1, \ldots, X_M)$. Let $T = \{T_i\}_{i \in I}$ be a collection of subsets of $[M]$, such that each element of $[M]$ appears in at least $K$ members of $T$. For $A \subseteq [M]$, let $X_A = \{X_j : j \in A\}$. Then,*

$$\sum_{i \in I} \mathbf{H}[X_{T_i}] \geq K \cdot \mathbf{H}[X].$$

We state and prove here the following "Shearer-like" inequality for information. A variant of this lemma was proved in [MT10].

**Lemma 7** (**Shearer-Like Inequality for Information**). *Let $X_1, \ldots, X_M$ be $M$ random variables, taking values in $\Omega_1, \ldots, \Omega_M$, respectively. Let $X = (X_1, \ldots, X_M)$ be a random variable, taking values in $\Omega_1 \times \cdots \times \Omega_M$. Let $T = \{T_i\}_{i \in I}$ be a collection of subsets of $[M]$, such that each element of $[M]$ appears in at most $\frac{1}{K}$ fraction of the members of $T$. For $A \subseteq [M]$, let $X_A = \{X_j : j \in A\}$. Then,*

$$K \cdot \mathop{\mathbf{E}}_{i \in_R I}[\mathbf{I}(X_{T_i})] \leq \mathbf{I}(X).$$

*Proof.* Fix $i \in I$. By the definition of information,

$$\mathbf{I}(X_{T_i}) = \sum_{j \in T_i} \log(|\Omega_j|) - \mathbf{H}[X_{T_i}].$$

For every $j \in [M]$, define $\mathbf{H}[X_j | X_{<j}] = \mathbf{H}[X_j | (X_\ell : \ell < j)]$. By the chain rule for the entropy function,

$$\mathbf{I}(X) = \sum_{j \in [M]} \Big( \log(|\Omega_j|) - \mathbf{H}[X_j | X_{<j}] \Big),$$
$$\mathbf{I}(X_{T_i}) = \sum_{j \in T_i} \Big( \log(|\Omega_j|) - \mathbf{H}[X_j | (X_\ell : \ell \in T_i, \ell < j)] \Big).$$

For every $j \in T_i$ it holds that $\mathbf{H}[X_j | (X_\ell : \ell \in T_i, \ell < j)] \geq \mathbf{H}[X_j | X_{<j}]$. Therefore,

$$\mathbf{I}(X_{T_i}) \leq \sum_{j \in T_i} \Big( \log(|\Omega_j|) - \mathbf{H}[X_j | X_{<j}] \Big).$$

Summing over all $i \in I$ we get that

$$\sum_{i \in I} \mathbf{I}(X_{T_i}) \leq \sum_{i \in I} \sum_{j \in T_i} \Big( \log(|\Omega_j|) - \mathbf{H}[X_j | X_{<j}] \Big). \tag{8}$$

For every $j \in [M]$, the term $\log(|\Omega_j|) - \mathbf{H}[X_j | X_{<j}]$ appears on the right-hand side of

Equation (8) at most $\frac{|I|}{K}$ times. Therefore,

$$\sum_{i\in I}\mathbf{I}(X_{T_i}) \le \frac{|I|}{K}\cdot\sum_{j\in[M]}\Big(\log(|\Omega_j|)-\mathbf{H}[X_j|X_{<j}]\Big)$$
$$=\frac{|I|}{K}\cdot\mathbf{I}(X).$$

Dividing by $\frac{|I|}{K}$ we get that the claim holds. □

The next lemma generalizes Lemma 7, and gives a Shearer-like inequality for relative entropy. A variant of this lemma was proved in [MT10]. The lemma will not be used in the paper, but we include it here as it may be useful in this context. The proof is given in Appendix A.

**Lemma 8** (**Shearer-Like Inequality for Relative Entropy**). *Let $P,Q:\Omega_1\times\cdots\times\Omega_M\to[0,1]$ be two distributions, such that $Q$ is a product distribution, i.e., for every $j\in[M]$, there exists $Q_j:\Omega_j\to[0,1]$, such that $Q(x_1,\ldots,x_M)=\prod_{j\in[M]}Q_j(x_j)$. Let $T=\{T_i\}_{i\in I}$ be a collection of subsets of $[M]$, such that each element of $[M]$ appears in at most $\frac{1}{K}$ fraction of the members of $T$. For $A\subseteq[M]$, let $P_A$ and $Q_A$ be the marginal distributions of $A$ in the distributions $P$ and $Q$ (respectively). Then,*

$$K\cdot\mathop{\mathbf{E}}_{i\in_R I}[\mathbf{D}(P_{T_i}\|Q_{T_i})]\le\mathbf{D}(P\|Q).$$

# 5 The Graph Correlation Lemma

**Lemma 9** (**Graph Correlation Lemma**). [1] *Let $G=(U\cup W,E)$ be a bipartite (multi)-graph with sets of vertices $U,W$ and (multi)-set of edges $E$, such that, $G$ is bi-regular and $|U|=|W|$. Let $M>T>k\in\mathbb{N}$ be such that, $T\le 2^{-20k}M$, and $k\ge 4$. For every $(u,w)\in E$, let $T(u,w)\subset[M]$ be a set of size $T$, such that, for every $u\in U$, each element of $[M]$ appears in at most $2^{-20k}$ fraction of the sets in $\{T(u,w)\}_{(u,w)\in E}$, and for every $w\in W$, each element of $[M]$ appears in at most $2^{-20k}$ fraction of the sets in $\{T(u,w)\}_{(u,w)\in E}$.*

*Let $\Sigma$ be a finite set. For every $u\in U$, let $X^u\in\Sigma^M$ be a random variable, such that, $\mathbf{I}(X^u)\le 2^{4k}$, and for every $w\in W$, let $Y^w\in\Sigma^M$ be a random variable, such that, $\mathbf{I}(Y^w)\le 2^{4k}$, and such that, for every $u\in U$ and $w\in W$, the random variables $X^u$ and $Y^w$ are mutually independent.*

*For $(u,w)\in E$, denote*

$$\mu(u,w)=\frac{\Pr_{X^u,Y^w}[X^u_{T(u,w)}=Y^w_{T(u,w)}]}{|\Sigma|^{-T}}.$$

[1] Many variants of this lemma can be proven. In particular, a similar argument can be used to prove a similar statement with sets $T(u,w)$ that are not of the same size. We state the lemma here for sets $T(u,w)$ of the same size $T$, for convenience of notation.

*Let*

$$\mathcal{D} = \{(u,w) \in E \ : \ \mu(u,w) \leq 1 - 2^{-4k}\}.$$

*Then,*

$$\frac{|\mathcal{D}|}{|E|} \leq 2^{-4k}.$$

*Proof.* We will start by proving the following claim.

**Claim 10.** *If $(u,w) \in \mathcal{D}$ then at least one of the following two inequalities holds,*

$$\mathbf{I}\left(X^u_{T(u,w)}\right) \geq 2^{-8k-4},$$

$$\mathbf{I}\left(Y^w_{T(u,w)}\right) \geq 2^{-8k-4}.$$

*Proof.* Assume $(u,w) \in \mathcal{D}$. Thus,

$$-2^{-4k} \geq \mu(u,w) - 1 = |\Sigma|^T \cdot \left( \Pr_{X^u, Y^w}[X^u_{T(u,w)} = Y^w_{T(u,w)}] - |\Sigma|^{-T} \right) =$$

$$|\Sigma|^T \cdot \left( \left( \sum_{z \in \Sigma^{T(u,w)}} \Pr_{X^u}[X^u_{T(u,w)} = z] \cdot \Pr_{Y^w}[Y^w_{T(u,w)} = z] \right) - |\Sigma|^{-T} \right) =$$

$$|\Sigma|^T \cdot \sum_{z \in \Sigma^{T(u,w)}} \left( \Pr_{X^u}[X^u_{T(u,w)} = z] - |\Sigma|^{-T} \right) \cdot \left( \Pr_{Y^w}[Y^w_{T(u,w)} = z] - |\Sigma|^{-T} \right). \tag{9}$$

In the last sum, we can omit the positive summands (and the inequality still holds). As for the negative summands, we split them into summands where $(\Pr[X^u_{T(u,w)} = z] - |\Sigma|^{-T})$ is negative and $(\Pr[Y^w_{T(u,w)} = z] - |\Sigma|^{-T})$ is positive, and summands where it's the other way around. In the first case, we bound the first term by

$$\left( \Pr_{X^u}[X^u_{T(u,w)} = z] - |\Sigma|^{-T} \right) \geq -|\Sigma|^{-T},$$

and for the second term, we use

$$\left( \Pr_{Y^w}[Y^w_{T(u,w)} = z] - |\Sigma|^{-T} \right) = \left| \Pr_{Y^w}[Y^w_{T(u,w)} = z] - |\Sigma|^{-T} \right|.$$

Similarly, in the second case, we bound the terms the other way around. Note also that we can add to the sum arbitrary negative summands (and the inequality still holds). Thus, Equation (9) implies

$$\begin{aligned} -2^{-4k} \ \geq \quad & |\Sigma|^T \cdot \sum_{z \in \Sigma^{T(u,w)}} \left(-|\Sigma|^{-T}\right) \cdot \left| \Pr_{Y^w}[Y^w_{T(u,w)} = z] - |\Sigma|^{-T} \right| + \\ & |\Sigma|^T \cdot \sum_{z \in \Sigma^{T(u,w)}} \left| \Pr_{X^u}[X^u_{T(u,w)} = z] - |\Sigma|^{-T} \right| \cdot \left(-|\Sigma|^{-T}\right) = \\ & - \sum_{z \in \Sigma^{T(u,w)}} \left| \Pr_{Y^w}[Y^w_{T(u,w)} = z] - |\Sigma|^{-T} \right| - \sum_{z \in \Sigma^{T(u,w)}} \left| \Pr_{X^u}[X^u_{T(u,w)} = z] - |\Sigma|^{-T} \right|, \end{aligned}$$

that is,

$$\sum_{z\in\Sigma^{T(u,w)}}\left|\Pr_{Y^w}[Y^w_{T(u,w)}=z]-|\Sigma|^{-T}\right|+\sum_{z\in\Sigma^{T(u,w)}}\left|\Pr_{X^u}[X^u_{T(u,w)}=z]-|\Sigma|^{-T}\right|\geq 2^{-4k}.$$

Hence, for every $(u,w)\in\mathcal{D}$, at least one of the following two inequalities holds,

$$\sum_{z\in\Sigma^{T(u,w)}}\left|\Pr_{X^u}[X^u_{T(u,w)}=z]-|\Sigma|^{-T}\right|\geq 2^{-4k-1},$$

$$\sum_{z\in\Sigma^{T(u,w)}}\left|\Pr_{Y^w}[Y^w_{T(u,w)}=z]-|\Sigma|^{-T}\right|\geq 2^{-4k-1}.$$

The claim follows by Pinsker's inequality. □

We will now proceed with the proof of Lemma 9. By Claim 10, we know that one of the following two statements must hold:

1. For at least half of the edges $(u,w)\in\mathcal{D}$, we have $\mathbf{I}\left(X^u_{T(u,w)}\right)\geq 2^{-8k-4}$.
2. For at least half of the edges $(u,w)\in\mathcal{D}$, we have $\mathbf{I}\left(Y^w_{T(u,w)}\right)\geq 2^{-8k-4}$.

Without loss of generality, assume that the first statement holds.

Assume for a contradiction that

$$\frac{|\mathcal{D}|}{|E|}>2^{-4k}.$$

Thus, by an averaging argument, there exists $u\in U$, such that, for at least $2^{-4k-1}$ fraction of the edges $(u,w)\in E$, we have $\mathbf{I}\left(X^u_{T(u,w)}\right)\geq 2^{-8k-4}$. Fix $u\in U$ that has this property. Denote by $E(u)$ the (multi)-set of edges in $E$ that contain $u$, that is, $E(u)=\{(u,w):(u,w)\in E\}$. Thus,

$$\mathop{\mathbf{E}}_{(u,w)\in_R E(u)}\left[\mathbf{I}\left(X^u_{T(u,w)}\right)\right]\geq 2^{-4k-1}\cdot 2^{-8k-4}=2^{-12k-5}.$$

Since each element of $[M]$ appears in at most $2^{-20k}$ fraction of the sets in $\{T(u,w)\}_{(u,w)\in E(u)}$, we have by Lemma 7,

$$\mathbf{I}(X^u)\geq 2^{-12k-5}\cdot 2^{20k}=2^{8k-5},$$

in contradiction to the assumption of the lemma. □

# 6 Communication Lower Bound

In this section we prove Theorem 1. Assume that $\pi$ is a deterministic communication protocol for the bursting noise game with parameter $k$, that has communication complexity at most $2^k$.

The section is devoted to showing that $\pi$ has error $\epsilon \geq 1-2^{-\Omega(k)}$ (when the inputs are selected according to the distribution $\mu$). That is, the protocol almost always errs. Observe that this also implies that every probabilistic protocol errs with probability $\epsilon \geq 1-2^{-\Omega(k)}$, as it is a distribution over deterministic protocols.

## 6.1 Notation

Let $\{R^1, \ldots, R^m\}$ be the rectangle partition induced by the protocol $\pi$, where $R^t = A^t \times B^t$ for $A^t, B^t \subseteq \{0,1\}^V$ and $m \leq 2^{2^k}$. We assume for simplicity and without loss of generality that $m = 2^{2^k}$ (as empty rectangles can always be added). Let $t \in [m]$. Let $X^t$ be a random variable taking values in $\{0,1\}^V$, that is uniformly distributed over $A^t$. Let $Y^t$ be a random variable taking values in $\{0,1\}^V$, that is uniformly distributed over $B^t$.

Let $i \in [c]$ be a multi-layer. Define $V_{<i} \subseteq V$ to be the set of vertices in multi-layers 1 to $i-1$. Define $V_i \subseteq V$ to be the set of vertices in multi-layer $i$. Define $V_{\geq i} \subseteq V$ to be the set of vertices in multi-layers $i$ to $c$. Define $V_{>i} \subseteq V$ to be the set of vertices in multi-layers $i+1$ to $c$. For $Z$ that is either a random variable taking values in $\{0,1\}^V$ or an element in $\{0,1\}^V$, we define $Z_{<i}$, $Z_i$, $Z_{\geq i}$, $Z_{>i}$ to be the projections of $Z$ to $V_{<i}$, $V_i$, $V_{\geq i}$, $V_{>i}$ (respectively).

Let $i \in [c]$ and $z \in \{0,1\}^{V_{<i}}$. Define $\Psi^z$ to be the set of all elements $\psi \in \{0,1\}^V$ with $\psi_{<i} = z$. It holds that $|\Psi^z| = |\{0,1\}^{V_{\geq i}}|$.

Let $i \in [c]$, $z \in \{0,1\}^{V_{<i}}$ and $t \in [m]$. Define $A^{t,z} = A^t \cap \Psi^z$ and $B^{t,z} = B^t \cap \Psi^z$. Define $R^{t,z} = A^{t,z} \times B^{t,z}$. Let $X^{t,z}$ be a random variable taking values in $\Psi^z$, that is uniformly distributed over $A^{t,z}$. Let $Y^{t,z}$ be a random variable taking values in $\Psi^z$, that is uniformly distributed over $B^{t,z}$.

## 6.2 Bounding the Information on the Noisy Multi-Layer

Let $i \in [c]$ and $z \in \{0,1\}^{V_{<i}}$. We define $\rho^{i,z} : [m] \to [0,1]$ to be the distribution that selects a rectangle index $t \in [m]$ as follows: Randomly select an input pair $(x,y) \in \Psi^z \times \Psi^z$. Select $t$ to be the index of the unique rectangle $R^t$ containing $(x,y)$. That is, $\rho^{i,z}(t)$ is the density of the rectangle $R^{t,z}$ with respect to input pairs that agree with $z$,

$$\rho^{i,z}(t) = \frac{|R^{t,z}|}{|\Psi^z \times \Psi^z|}.$$

The following lemma shows that, in expectation, the distribution of the projections of inputs in $R^{t,z}$ to multi-layer $i$ is close to uniform.

**Lemma 11.** *It holds that*

$$\mathop{\mathbf{E}}_{i \in_R [c]} \mathop{\mathbf{E}}_{z \in_R \{0,1\}^{V_{<i}}} \mathop{\mathbf{E}}_{t \leftarrow \rho^{i,z}} \left[\mathbf{I}\left(X_i^{t,z}\right)\right] \leq \frac{m}{c},$$

*and similarly,*

$$\mathop{\mathbf{E}}_{i \in_R [c]} \mathop{\mathbf{E}}_{z \in_R \{0,1\}^{V_{<i}}} \mathop{\mathbf{E}}_{t \leftarrow \rho^{i,z}} \left[\mathbf{I}\left(Y_i^{t,z}\right)\right] \leq \frac{m}{c}.$$

*Proof.* Fix $i \in [c]$. It holds that

$$\begin{aligned}
&\mathop{\mathbf{E}}_{z\in_R\{0,1\}^{V_{<i}}} \mathop{\mathbf{E}}_{t\leftarrow\rho^{i,z}} \left[\mathbf{I}\left(X_i^{t,z}\right)\right] \\
&= \sum_{z\in\{0,1\}^{V_{<i}}} \frac{1}{|\{0,1\}^{V_{<i}}|} \sum_{t\in[m]} \frac{|R^{t,z}|}{|\{0,1\}^{V_{\geq i}}|^2} \cdot \mathbf{I}\left(X_i^{t,z}\right) \\
&= \sum_{z\in\{0,1\}^{V_{<i}}} \frac{1}{|\{0,1\}^{V}|} \sum_{t\in[m]} \frac{|A^{t,z}| \cdot |B^{t,z}|}{|\{0,1\}^{V_{\geq i}}|} \cdot \mathbf{I}\left(X_i^{t,z}\right) \\
&\leq \sum_{z\in\{0,1\}^{V_{<i}}} \frac{1}{|\{0,1\}^{V}|} \sum_{t\in[m]} \frac{|A^{t,z}| \cdot |\{0,1\}^{V_{\geq i}}|}{|\{0,1\}^{V_{\geq i}}|} \cdot \mathbf{I}\left(X_i^{t,z}\right) \\
&= \sum_{t\in[m]} \frac{1}{|\{0,1\}^{V}|} \sum_{z\in\{0,1\}^{V_{<i}}} |A^{t,z}| \cdot \mathbf{I}\left(X_i^{t,z}\right) \\
&= \sum_{t\in[m]} \frac{|A^t|}{|\{0,1\}^{V}|} \sum_{z\in\{0,1\}^{V_{<i}}} \frac{|A^{t,z}|}{|A^t|} \cdot \mathbf{I}\left(X_i^{t,z}\right) \\
&= \sum_{t\in[m]} \frac{|A^t|}{|\{0,1\}^{V}|} \sum_{z\in\{0,1\}^{V_{<i}}} \frac{|A^{t,z}|}{|A^t|} \left(|V_i| - \mathbf{H}\left(X_i^{t,z}\right)\right).
\end{aligned}$$

Denote

$$s := \sum_{t\in[m]} \frac{|A^t|}{|\{0,1\}^{V}|} = \sum_{t\in[m]} \sum_{z\in\{0,1\}^{V_{<i}}} \frac{|A^{t,z}|}{|\{0,1\}^{V}|}.$$

Observe that $1 \leq s \leq m$. We have that

$$\begin{aligned}
&\mathop{\mathbf{E}}_{z\in_R\{0,1\}^{V_{<i}}} \mathop{\mathbf{E}}_{t\leftarrow\rho^{i,z}} \left[\mathbf{I}\left(X_i^{t,z}\right)\right] \\
&\leq s|V_i| - \sum_{t\in[m]} \frac{|A^t|}{|\{0,1\}^{V}|} \sum_{z\in\{0,1\}^{V_{<i}}} \frac{|A^{t,z}|}{|A^t|} \cdot \mathbf{H}\left(X_i^{t,z}\right) \\
&= s|V_i| - \sum_{t\in[m]} \frac{|A^t|}{|\{0,1\}^{V}|} \sum_{z\in\{0,1\}^{V_{<i}}} \frac{|A^{t,z}|}{|A^t|} \cdot \mathbf{H}\left(X_i^t | X_{<i}^t = z\right) \\
&= s|V_i| - \sum_{t\in[m]} \frac{|A^t|}{|\{0,1\}^{V}|} \mathop{\mathbf{E}}_{z\leftarrow X_{<i}^t} \left[\mathbf{H}\left(X_i^t | X_{<i}^t = z\right)\right] \\
&= s|V_i| - \sum_{t\in[m]} \frac{|A^t|}{|\{0,1\}^{V}|} \cdot \mathbf{H}\left(X_i^t | X_{<i}^t\right).
\end{aligned}$$

By the chain rule for the entropy function,

$$\begin{aligned}
&\mathop{\mathbf{E}}_{i\in_R[c]} \mathop{\mathbf{E}}_{z\in_R\{0,1\}^{V_{<i}}} \mathop{\mathbf{E}}_{t\leftarrow\rho^{i,z}} \left[\mathbf{I}\left(X_i^{t,z}\right)\right] \\
&\leq \mathop{\mathbf{E}}_{i\in_R[c]} \left[ s|V_i| - \sum_{t\in[m]} \frac{|A^t|}{|\{0,1\}^V|} \cdot \mathbf{H}\left(X_i^t | X_{<i}^t\right) \right] \\
&= \mathop{\mathbf{E}}_{i\in_R[c]} \left[s|V_i|\right] - \sum_{t\in[m]} \frac{|A^t|}{|\{0,1\}^V|} \cdot \mathop{\mathbf{E}}_{i\in_R[c]} \left[\mathbf{H}\left(X_i^t | X_{<i}^t\right)\right] \\
&= \frac{s|V|}{c} - \frac{1}{c} \sum_{t\in[m]} \frac{|A^t|}{|\{0,1\}^V|} \sum_{i\in[c]} \left[\mathbf{H}\left(X_i^t | X_{<i}^t\right)\right] \\
&= \frac{s}{c}\left( |V| - \sum_{t\in[m]} \frac{|A^t|}{s|\{0,1\}^V|} \cdot \mathbf{H}(X^t) \right) \\
&= \frac{s}{c}\left( |V| + \sum_{t\in[m]} \frac{|A^t|}{s|\{0,1\}^V|} \cdot \log\left(\frac{1}{|A^t|}\right) \right) \\
&\leq \frac{s}{c}\left( |V| + \log\left( \sum_{t\in[m]} \frac{|A^t|}{s|\{0,1\}^V|} \cdot \frac{1}{|A^t|} \right) \right) \\
&\leq \frac{s}{c} \log\left(\frac{m}{s}\right) \\
&\leq \frac{m}{c},
\end{aligned}$$

where the third to last inequality is by the concavity of the log function. The last inequality holds as $-x\log(x) < 1$ for $x \in [0,1]$. □

## 6.3 Unique Answer Rectangles

**Lemma 12** (**Unique Answer Lemma**). *Let $A$ be a set of inputs for the first player, $B$ be a set of inputs for the second player, and $\Omega$ be a set of possible outputs. Let $\pi_1 : A \to \Omega$, $\pi_2 : B \to \Omega$ be any functions determining the players' outputs.*

*Let $\gamma > 0$. There exist a partition of $A$ into a disjoint union $A = A^1 \cup \cdots \cup A^\ell$, a partition of $B$ into a disjoint union $B = B^1 \cup \cdots \cup B^\ell$, where $\ell = O(1/\gamma^4)$, and for every $s_1, s_2 \in [\ell]$ there exist functions $\pi_1^{s_1,s_2} : A^{s_1} \to \Omega$, $\pi_2^{s_1,s_2} : B^{s_2} \to \Omega$, such that the followings hold: Denote $R^{s_1,s_2} = A^{s_1} \times B^{s_2}$.*

1. *Let $s_1, s_2 \in [\ell]$ and let $(x,y) \in R^{s_1,s_2}$. If $\pi_1(x) = \pi_2(y)$ then*

$$\pi_1^{s_1,s_2}(x) = \pi_2^{s_1,s_2}(y) = \pi_1(x) = \pi_2(y).$$

2. *For $s_1, s_2 \in [\ell]$, we say that the rectangle $R^{s_1,s_2}$ is a* unique answer rectangle *if there exists $\omega \in \Omega$ such that for every $(x,y) \in R^{s_1,s_2}$, it holds that $\pi_1^{s_1,s_2}(x) = \pi_2^{s_1,s_2}(y) = \omega$.*

*Let $S$ be the union of all unique answer rectangles $R^{s_1,s_2}$, where $s_1, s_2 \in [\ell]$. Then,*

$$\Pr_{(x,y)\in_R A\times B}[(x,y) \notin S] \leq \gamma.$$

3. *For $s_1, s_2 \in [\ell]$, we say that the rectangle $R^{s_1,s_2}$ is a $\gamma$-*large rectangle *if $|A^{s_1}||B^{s_2}| \geq \frac{\gamma^4}{10^4}|A||B|$. Let $L$ be the union of all $\gamma$-large rectangles $R^{s_1,s_2}$, where $s_1, s_2 \in [\ell]$. Then,*

$$\Pr_{(x,y)\in_R A\times B}[(x,y) \notin L] \leq \gamma.$$

*Proof.* For $\Omega' \subseteq \Omega$, define

$$\begin{aligned}
p_1(\Omega') &= \Pr_{x\in_R A}[\pi_1(x) \in \Omega'],\\
p_2(\Omega') &= \Pr_{y\in_R B}[\pi_2(y) \in \Omega'],\\
A(\Omega') &= \{x \in A \ : \ \pi_1(x) \in \Omega'\},\\
B(\Omega') &= \{y \in B \ : \ \pi_2(y) \in \Omega'\}.
\end{aligned}$$

We define the partitions of $A$ and $B$ as follows: For every $\omega \in \Omega$ such that either $p_1(\{\omega\}) \geq \frac{\gamma}{10}$ or $p_2(\{\omega\}) \geq \frac{\gamma}{10}$, add $A(\{\omega\})$ to the partition of $A$ and $B(\{\omega\})$ to the partition of $B$. So far, we added at most $\frac{20}{\gamma}$ sets to each partition. Let

$$T = \left\{\omega \in \Omega \ : \ p_1(\{\omega\}), p_2(\{\omega\}) < \frac{\gamma}{10}\right\}.$$

Let $T_1, \ldots, T_t$ be a minimal partition of $T$, such that for every $j \in [t]$, we have

$$p_1(T_j), p_2(T_j) < \frac{\gamma}{10}.$$

Since $T$ is minimal, there is at most one set $T_j$ in $T$ with both $p_1(T_j), p_2(T_j) < \frac{\gamma}{20}$ (as if there were two such sets we could have merged them). Therefore, $t \leq 2 \cdot \frac{20}{\gamma} + 1$. For every $i \in [t]$, add $A(T_i)$ to the partition of $A$ and $B(T_i)$ to the partition of $B$. This concludes the definition of the partitions of $A$ and $B$, where the number of sets in each partition, denoted by $\ell$, is at most $\frac{20}{\gamma} + \frac{40}{\gamma} + 1 \leq \frac{70}{\gamma}$.

Fix $s_1, s_2 \in [\ell]$. Let $\Omega_A = \{\pi_1(x) \ : \ x \in A^{s_1}\}$ and $\Omega_B = \{\pi_2(y) \ : \ y \in B^{s_2}\}$. For every $(x,y) \in A^{s_1} \times B^{s_2}$, we define $\pi_1^{s_1,s_2}(x)$ and $\pi_2^{s_1,s_2}(y)$ by the following steps (once the conditions of a step are fulfilled and the outputs of the functions are defined we do not continue to the next step):

1. If $p_1(\Omega_A) \geq \frac{\gamma}{10}$, then $\Omega_A$ contains a single value $\omega$. Define $\pi_1^{s_1,s_2}(x) = \pi_2^{s_1,s_2}(y) = \omega$.
2. If $p_2(\Omega_B) \geq \frac{\gamma}{10}$, then $\Omega_B$ contains a single value $\omega$. Define $\pi_1^{s_1,s_2}(x) = \pi_2^{s_1,s_2}(y) = \omega$.
3. If $\Omega_A \cap \Omega_B \neq \emptyset$, define $\pi_1^{s_1,s_2}(x) = \pi_1(x)$ and $\pi_2^{s_1,s_2}(y) = \pi_2(y)$.
4. Define $\pi_1^{s_1,s_2}(x) = \pi_2^{s_1,s_2}(y) = w$ for some fixed $w \in \Omega$.

We prove that the three requirements of the lemma are met:

1. Assume that $\pi_1(x) = \pi_2(y) = \omega$. Then, $\omega \in \Omega_A \cap \Omega_B$. Thus, the outputs $\pi_1^{s_1,s_2}(x)$ and $\pi_2^{s_1,s_2}(y)$ are defined in one of the first three steps, and therefore, $\pi_1^{s_1,s_2}(x) = \pi_2^{s_1,s_2}(y) = \omega$.

2. Assume that $R^{s_1,s_2}$ is not a unique answer rectangle. Then, there exists $(x,y) \in R^{s_1,s_2}$ such that $\pi_1^{s_1,s_2}(x) \neq \pi_2^{s_1,s_2}(y)$. Then, the outputs are defined in Step 3, and it holds that $p_1(\Omega_A), p_2(\Omega_B) < \frac{\gamma}{10}$ and $\Omega_A \cap \Omega_B \neq \emptyset$. By the definition of the partitions of $A$ and $B$, there exists $i \in [t]$ such that $\Omega_A, \Omega_B \subseteq T_i$ (we mention that if there is no $\omega \in T_i$ with $p_1(\omega) = 0$ or $p_2(\omega) = 0$ then $\Omega_A = \Omega_B = T_i$). Therefore,

$$\Pr_{(x,y)\in_R A\times B}[(x,y) \notin S] \leq \sum_{i\in[t]} p_1(T_i) \cdot p_2(T_i) \leq \frac{\gamma}{10} \sum_{i\in[t]} p_1(T_i) \leq \frac{\gamma}{10}.$$

3. Since the number of rectangles $R^{s_1,s_2}$, where $s_1, s_2 \in [\ell]$, is $\ell^2 \leq \frac{70^2}{\gamma^2}$,

$$\Pr_{(x,y)\in_R A\times B}[(x,y) \notin L] \leq \frac{70^2}{\gamma^2} \cdot \frac{\gamma^4}{10^4} \leq \gamma.$$

□

## 6.4 Good Rectangles

Fix $\gamma = \gamma(k) = 2^{-k/4}$ (in particular, $\gamma$ is sub-constant). Let $i \in [c]$. We say that $i$ is *good* if

$$\mathop{\mathbf{E}}_{z\in_R\{0,1\}^{V<i}} \; \mathop{\mathbf{E}}_{t\leftarrow\rho^{i,z}} \left[\mathbf{I}\left(X_i^{t,z}\right)\right] \leq \frac{m}{\gamma c}. \tag{10}$$

$$\mathop{\mathbf{E}}_{z\in_R\{0,1\}^{V<i}} \; \mathop{\mathbf{E}}_{t\leftarrow\rho^{i,z}} \left[\mathbf{I}\left(Y_i^{t,z}\right)\right] \leq \frac{m}{\gamma c}. \tag{11}$$

By Markov's inequality and Lemma 11,

$$\Pr_{i\in_R[c]}[\ i \text{ is good }] \geq 1 - 2\gamma. \tag{12}$$

For the rest of the lower bound proof, fix a good $i \in [c]$.

For every $z \in \{0,1\}^{V<i}$ and $t \in [m]$, consider the rectangle $R^{t,z} = A^{t,z} \times B^{t,z}$. Apply Lemma 12 to the sets $A^{t,z}$, $B^{t,z}$, where the functions $\pi_1, \pi_2$ are the outputs of the two players for the rectangle $R^t$, in the protocol $\pi$. Lemma 12 partitions each rectangle $R^{t,z}$ into $\ell = O(1/\gamma^8)$ new rectangles, denoted $R^{t,1,z}, \ldots, R^{t,\ell,z}$. Observe that $\{R^{t,s,z}\}_{t\in[m],s\in[\ell]}$ is a cover of $\Psi^z \times \Psi^z$,

$$\bigcup_{\substack{t\in[m]\\ s\in[\ell]}} R^{t,s,z} = \Psi^z \times \Psi^z. \tag{13}$$

For $t \in [m]$ and $s \in [\ell]$, denote $R^{t,s,z} = A^{t,s,z} \times B^{t,s,z}$, where $A^{t,s,z}, B^{t,s,z} \subseteq \Psi^z$. Let $X^{t,s,z}$ be a random variable taking values in $\Psi^z$, that is uniformly distributed over $A^{t,s,z}$. Let $Y^{t,s,z}$ be a random variable taking values in $\Psi^z$, that is uniformly distributed over $B^{t,s,z}$.

Let $z \in \{0,1\}^{V_{<i}}$. We define $\eta^{i,z} : [m] \times [\ell] \to [0,1]$ to be the distribution that selects rectangle indices $(t,s) \in [m] \times [\ell]$ as follows: Randomly select an input pair $(x,y) \in \Psi^z \times \Psi^z$. Select $(t,s)$ to be the indices of the unique rectangle $R^{t,s,z}$ containing $(x,y)$. That is, $\eta^{i,z}(t,s)$ is the density of the rectangle $R^{t,s,z}$ with respect to input pairs that agree with $z$,

$$\eta^{i,z}(t,s) = \frac{|R^{t,s,z}|}{|\Psi^z \times \Psi^z|}.$$

Observe that for every $t \in [m]$ it holds that

$$\rho^{i,z}(t) = \sum_{s \in [\ell]} \eta^{i,z}(t,s).$$

We say that $(i,z,t,s)$ is *good* if all the followings hold:

1. $R^{t,s,z}$ is a unique answer rectangle (as in Lemma 12). Recall that for each unique answer rectangle $R^{t,s,z}$, there is a unique leaf in $V$, denoted $\omega^{t,s,z}$, returned as an output by both players on all input pairs in the rectangle $R^{t,s,z}$.
2. $R^{t,s,z}$ is $\gamma$-large (as in Lemma 12). That is, $|R^{t,s,z}| \geq \frac{\gamma^4}{10^4}|R^{t,z}|$.
3. $\mathbf{I}\left(X^{t,s,z}\right) \leq 2\log(m)$, and therefore also, $\mathbf{I}\left(X^{t,z}\right) \leq 2\log(m)$.
4. $\mathbf{I}\left(Y^{t,s,z}\right) \leq 2\log(m)$, and therefore also, $\mathbf{I}\left(Y^{t,z}\right) \leq 2\log(m)$.
5. $\mathbf{I}\left(X_i^{t,z}\right) \leq \frac{m}{\gamma^2 c}$.
6. $\mathbf{I}\left(Y_i^{t,z}\right) \leq \frac{m}{\gamma^2 c}$.
7. $\mathbf{I}\left(X_i^{t,s,z}\right) \leq O\left(\log(1/\gamma)\right)$.
8. $\mathbf{I}\left(Y_i^{t,s,z}\right) \leq O\left(\log(1/\gamma)\right)$.

**Lemma 13.** *It holds that*

$$\Pr_{\substack{z \in_R \{0,1\}^{V_{<i}}, \\ (t,s) \leftarrow \eta^{i,z}}} [\ (i,z,t,s)\ \textit{is good}\ ] \geq 1 - O(\gamma).$$

*Proof.* We claim that each of the eight requirements in the definition of a good tuple $(i,z,t,s)$ is violated with probability $O(\gamma)$:

1. By the second item of Lemma 12.
2. By the third item of Lemma 12.

3. If $\mathbf{I}\left(X^{t,s,z}\right) > 2\log(m)$, then $\eta^{i,z}(t,s) = \frac{|R^{t,s,z}|}{|\Psi^z \times \Psi^z|} \leq 1/m^2$. Since for every $z$ there are at most $m \cdot \ell$ rectangles $R^{t,s,z}$, the $\eta^{i,z}$-measure of all such rectangles is at most $\ell/m < O(\gamma)$.

4. Same.

5. Since $i$ is good, follows from Equation (10) and Markov's inequality.

6. Since $i$ is good, follows from Equation (11) and Markov's inequality.

7. Since the the second and fifth properties of a good $(i,z,t,s)$ imply this (seventh) property as follows. Let $\tau$ and $\tau'$ be the distributions of the random variables $X_i^{t,s,z}$ and $X_i^{t,z}$ (respectively). By the second property of a good tuple $(i,z,t,s)$ (i.e., $\gamma$-largeness), for every $\omega \in \{0,1\}^{V_i}$ it holds that $\tau(\omega) \leq \frac{10^4}{\gamma^4} \cdot \tau'(\omega)$. We denote $\Psi = \{0,1\}^{V_i}$, $\Psi^{pos} = \{\omega \in \Psi : \log\left(|\Psi| \cdot \tau'(\omega)\right) \geq 0\}$ and $\Psi^{neg} = \Psi \setminus \Psi^{pos}$. Therefore,

$$\begin{aligned}
\mathbf{I}\left(X_i^{t,s,z}\right) = \mathbf{I}(\tau) &= \sum_{\omega\in\Psi} \tau(\omega)\log\left(|\Psi|\cdot\tau(\omega)\right) \\
&\leq \sum_{\omega\in\Psi} \tau(\omega)\log\left(|\Psi|\cdot\frac{10^4}{\gamma^4}\cdot\tau'(\omega)\right) \\
&\leq O\left(\log(1/\gamma)\right) + \sum_{\omega\in\Psi} \tau(\omega)\log\left(|\Psi|\cdot\tau'(\omega)\right) \\
&\leq O\left(\log(1/\gamma)\right) + \sum_{\omega\in\Psi^{pos}} \tau(\omega)\log\left(|\Psi|\cdot\tau'(\omega)\right) \\
&\leq O\left(\log(1/\gamma)\right) + \frac{10^4}{\gamma^4}\sum_{\omega\in\Psi^{pos}} \tau'(\omega)\log\left(|\Psi|\cdot\tau'(\omega)\right).
\end{aligned}$$

By the fifth property of a good tuple $(i,z,t,s)$, it holds that

$$\mathbf{I}(\tau') = \mathbf{I}\left(X_i^{t,z}\right) \leq \frac{m}{\gamma^2 c} < 0.01,$$

and thus by Lemma 5.11 in [KR13] (stated for convenience in Appendix A, Lemma 20)

$$-\sum_{\omega\in\Psi^{neg}} \tau'(\omega)\log\left(|\Psi|\cdot\tau'(\omega)\right) < 4\mathbf{I}(\tau')^{0.1}.$$

Therefore,

$$\begin{aligned}
\mathbf{I}\left(X_i^{t,s,z}\right) &< O\left(\log(1/\gamma)\right) + \frac{10^4}{\gamma^4}\left(\sum_{\omega\in\Psi} \tau'(\omega)\log\left(|\Psi|\cdot\tau'(\omega)\right) + 4\mathbf{I}(\tau')^{0.1}\right) \\
&= O\left(\log(1/\gamma)\right) + \frac{10^4}{\gamma^4}\left(\mathbf{I}(\tau') + 4\mathbf{I}(\tau')^{0.1}\right) \\
&\leq O\left(\log(1/\gamma)\right).
\end{aligned}$$

8. Since the the second and sixth properties of a good $(i, z, t, s)$ imply this (eighth) property, as above.

□

## 6.5 Proof of Theorem 1

In this section we prove Theorem 1. Recall that we fixed a good $i$. Let $\mathcal{G}_i$ be the set of all $(t, s, z) \in [m] \times [\ell] \times \{0,1\}^{V_{<i}}$ such that $(i, z, t, s)$ is good (see Section 6.4). Let $(t, s, z) \in \mathcal{G}_i$. The rectangle $R^{t,s,z}$ is a unique answer rectangle. Denote its unique answer by $\omega^{t,s,z}$. Let

$$\Lambda^{t,s,z} = \left\{(x, y) \in \mathrm{supp}(\mu_i) \ : \ \omega^{t,s,z} \text{ is an incorrect answer for } (x, y)\right\}.$$

Let $S_i$ be the set of inputs $(x, y) \in \mathrm{supp}(\mu_i)$ that the protocol $\pi$ errs on, when the noisy multi-layer is $i$. Our goal is to lower bound the size of $S_i$. Observe that the protocol $\pi$ errs on all the inputs in $\Lambda^{t,s,z} \cap R^{t,s,z}$, when the noisy multi-layer is $i$, for the following reason: If the protocol $\pi$ is correct on $(x, y) \in R^{t,s,z}$, then in the protocol $\pi$ both players output $\omega^{t,s,z}$ on the input $(x, y)$. This is true since if the protocol $\pi$ is correct on $(x, y)$, then, in particular, both players return the same output $\omega$ on $(x, y)$ in the protocol $\pi$. In this case, by the first item in Lemma 12, the output on the rectangle $R^{t,s,z}$ is the same as the original output, i.e., $\omega = \omega^{t,s,z}$. Thus,

$$\Lambda^{t,s,z} \cap R^{t,s,z} \subseteq S_i.$$

We lower bound the size of $S_i$ as follows:

$$\begin{aligned} |S_i| &\geq \sum_{(t,s,z)\in\mathcal{G}_i} |R^{t,s,z}| \cdot \Pr_{(x,y)\in_R R^{t,s,z}} \left[(x, y) \in \Lambda^{t,s,z}\right] \\ &= \sum_{(t,s,z)\in\mathcal{G}_i} |R^{t,s,z}| \cdot \Pr\left[(X^{t,s,z}, Y^{t,s,z}) \in \Lambda^{t,s,z}\right]. \end{aligned}$$

We apply Lemma 14 (stated and proved in Section 6.6), and get that for $(t, s, z) \in \mathcal{G}_i$,

$$\Pr\left[(X^{t,s,z}, Y^{t,s,z}) \in \Lambda^{t,s,z}\right] \geq \left(1 - O(2^{-2k}/\gamma^4)\right) \Pr_{(x,y)\in_R \Psi^z\times\Psi^z} \left[(x, y) \in \mathrm{supp}(\mu_i)\right].$$

Hence,

$$\begin{aligned} |S_i| &\geq \left(1 - O(2^{-2k}/\gamma^4)\right) \sum_{(t,s,z)\in\mathcal{G}_i} \Pr_{(x,y)\in_R \Psi^z\times\Psi^z} \left[(x, y) \in \mathrm{supp}(\mu_i)\right] \cdot |R^{t,s,z}| \\ &= \left(1 - O(2^{-2k}/\gamma^4)\right) \frac{|\mathrm{supp}(\mu_i)|}{|\{0,1\}^{V_{<i}}| \cdot |\{0,1\}^{V_{\geq i}}|^2} \sum_{(t,s,z)\in\mathcal{G}_i} |R^{t,s,z}|. \end{aligned}$$

By Lemma 13,

$$\sum_{(t,s,z)\in\mathcal{G}_i} |R^{t,s,z}| \geq (1 - O(\gamma)) \cdot |\{0,1\}^{V_{<i}}| \cdot |\{0,1\}^{V_{\geq i}}|^2.$$

Therefore,

$$|S_i| \geq \left(1 - O(2^{-2k}/\gamma^4) - O(\gamma)\right) \cdot |\mathrm{supp}(\mu_i)|,$$

which implies (recall that $\gamma = 2^{-k/4}$)

$$\frac{|S_i|}{|\mathrm{supp}(\mu_i)|} \geq 1 - O(2^{-2k}/\gamma^4) - O(\gamma) \geq 1 - 2^{-\Omega(k)}.$$

We conclude that the protocol $\pi$ errs on $1 - 2^{-\Omega(k)}$ fraction of the inputs in $\mathrm{supp}(\mu_i)$, when the noisy multi-layer is $i$. The lower bound follows, as by Equation (12), a fraction of at least $1 - O(\gamma)$ of the multi-layers $i$ are good.

## 6.6 Applying the Graph Correlation Lemma

**Lemma 14.** *Let $(t, s, z) \in \mathcal{G}_i$. It holds that*

$$\Pr\left[(X^{t,s,z}, Y^{t,s,z}) \in \Lambda^{t,s,z}\right] \geq \left(1 - O(2^{-2k}/\gamma^4)\right) \Pr_{(x,y) \in_R \Psi^z \times \Psi^z}\left[(x, y) \in \mathrm{supp}(\mu_i)\right].$$

*Proof.* Denote $P := X^{t,s,z}$ and $Q := Y^{t,s,z}$, and note that $P$ and $Q$ are independent random variables over the domain $\Psi^z$. Denote $I := \max\{\mathbf{I}(P), \mathbf{I}(Q), 1\}$, and note that $I \leq 2\log(m) \leq 2^{k+1}$ (by the third and forth properties of a good tuple $(i, z, t, s)$).

Let $G = (U \cup W, E)$ be the complete bipartite graph with sets of vertices $U, W$ and set of edges $E$, defined as follows: Let $U = W = \{0,1\}^{V_i}$ be the set of all boolean assignments to the vertices in multi-layer $i$. Let $E = U \times W$.

Let $M$ be the number of vertices in layer $(i+1)^*$ of the tree $\mathcal{T}$. We identify the set $[M]$ with the set of vertices in layer $(i+1)^*$. Let $u \in U, w \in W$. We define $T(u, w) \subset [M]$ to be the set of all vertices in layer $(i+1)^*$ that are set to be non-noisy for inputs $u, w$, by Algorithm 1 defining $\mu$, when the noisy multi-layer is $i$. That is, $T(u, w)$ is the set of all typical vertices in layer $(i+1)^*$ with respect to $i, u, w$. Observe that $u$ and $w$ determine for every vertex in layer $(i+1)^*$ if it is noisy or not.

Note that by a symmetry argument, $T(u, w)$ is of the same size $T$ for every $u, w$. By the definition of the bursting noise game and by the Chernoff bound, for any fixed $u$ or $w$ and every $v \in [M]$, it holds that at most a fraction of $2^{-20k}$ of the sets $\{T(u, w)\}_{(u,w)\in E}$ contain $v$.

Define the bad sets:

$$\mathcal{D}_1 = \left\{u \in U \;:\; \Pr_P[P_i = u] = 0 \;\text{ or }\; \mathbf{I}(P_{>i}|P_i = u) > 2^{4k}\right\},$$

$$\mathcal{D}_2 = \left\{w \in W \;:\; \Pr_Q[Q_i = w] = 0 \;\text{ or }\; \mathbf{I}(Q_{>i}|Q_i = w) > 2^{4k}\right\}.$$

By the chain rule for the entropy function,

$$\begin{aligned} I \geq \mathbf{I}(P) &= \log(|\Psi^z|) - \mathbf{H}(P) \\ &= \log(|U| \cdot |\{0,1\}^{V_{>i}}|) - \mathbf{H}(P_i, P_{>i}) \\ &= \log(|U|) + \log(|\{0,1\}^{V_{>i}}|) - \mathbf{H}(P_i) - \mathbf{H}(P_{>i}|P_i) \\ &= \mathbf{I}(P_i) + \log(|\{0,1\}^{V_{>i}}|) - \mathop{\mathbf{E}}_{u \leftarrow P_i}\left[\mathbf{H}(P_{>i}|P_i = u)\right] \\ &\geq \mathop{\mathbf{E}}_{u \leftarrow P_i}\left[\mathbf{I}(P_{>i}|P_i = u)\right]. \end{aligned}$$

By Markov's inequality,

$$\Pr_P\left[P_i \in \mathcal{D}_1\right] \leq \frac{I}{2^{4k}}. \tag{14}$$

By a similar argument,

$$\Pr_Q\left[Q_i \in \mathcal{D}_2\right] \leq \frac{I}{2^{4k}}. \tag{15}$$

For $u \notin \mathcal{D}_1$, we define the random variable $X^u$ to be $(P_{>i}|P_i = u)$, that is, $X^u$ has the distribution of $P_{>i}$ conditioned on the event $P_i = u$. For $u \in \mathcal{D}_1$, we define the random variable $X^u$ to be uniformly distributed over $\{0,1\}^{V_{>i}}$. Similarly, for $w \notin \mathcal{D}_2$, we define the random variable $Y^w$ to be $(Q_{>i}|Q_i = w)$, that is, $Y^w$ has the distribution of $Q_{>i}$ conditioned on the event $Q_i = w$. For $w \in \mathcal{D}_2$, we define the random variable $Y^w$ to be uniformly distributed over $\{0,1\}^{V_{>i}}$. Let $\Sigma$ be the set of all possible boolean assignments to the vertices of a subtree of $\mathcal{T}$ rooted at layer $(i+1)^*$.

By Lemma 9 applied to the graph $G$, there exists a set $\mathcal{D} \subset E$ such that

$$\frac{|\mathcal{D}|}{|E|} \leq 2^{-4k}, \tag{16}$$

and for every $(u,w) \notin \mathcal{D}$ it holds that

$$\Pr_{X^u, Y^w}\left[X^u_{T(u,w)} = Y^w_{T(u,w)}\right] \geq \left(1 - 2^{-4k}\right)|\Sigma|^{-T}. \tag{17}$$

Let $\mathcal{D}' \subseteq E$ be the set of all $(u,w) \in E$ for which the output $\omega^{t,s,z}$ is correct for inputs $(x,y) \in \mathrm{supp}(\mu_i)$, with $x_i = u$ and $y_i = w$. Note that if the noise is taken on multi-layer $i$, then $u$ and $w$ determine the correctness of $\omega^{t,s,z}$. By the definition of the bursting noise game and by the Chernoff bound,

$$|\mathcal{D}'| \leq 2^{-4k}|E|, \tag{18}$$

as the output $\omega^{t,s,z}$ is correct only if it has an ancestor in $T(u,w)$.

It holds that

$$\begin{aligned}
&\Pr_{P,Q}\left[(P,Q)\in\Lambda^{t,s,z}\right]\\
&=\sum_{(u,w)\in E\setminus\mathcal{D}'}\Pr_{P}\left[P_i=u\right]\cdot\Pr_{Q}\left[Q_i=w\right]\cdot\Pr_{P,Q}\left[(P,Q)\in\mathrm{supp}(\mu_i)\ \mid P_i=u,Q_i=w\right]\\
&\geq\sum_{\substack{u\in U\setminus\mathcal{D}_1,\\ w\in W\setminus\mathcal{D}_2,\\ (u,w)\notin\mathcal{D}\cup\mathcal{D}'}}\Pr_{P}\left[P_i=u\right]\cdot\Pr_{Q}\left[Q_i=w\right]\cdot\Pr_{P,Q}\left[(P,Q)\in\mathrm{supp}(\mu_i)\ \mid P_i=u,Q_i=w\right].
\end{aligned}$$

By the definition of the bursting noise game (when the noisy multi-layer is $i$), for every $u,w$, the following holds: Conditioned on $P_i=u$ and $Q_i=w$, we have $(P,Q)\in\mathrm{supp}(\mu_i)$ if and only if $P_{>i}$ and $Q_{>i}$ agree on the subtrees rooted at vertices in $T(u,w)$ (these are the non-noisy subtrees). Therefore, using Equation (17) and the fact that $E$ contains all pairs $(u,w)$,

$$\begin{aligned}
&\Pr_{P,Q}\left[(P,Q)\in\Lambda^{t,s,z}\right]\\
&\geq\sum_{\substack{u\in U\setminus\mathcal{D}_1,\\ w\in W\setminus\mathcal{D}_2,\\ (u,w)\notin\mathcal{D}\cup\mathcal{D}'}}\Pr_{P}\left[P_i=u\right]\cdot\Pr_{Q}\left[Q_i=w\right]\cdot\Pr_{X^u,Y^w}\left[X^u_{T(u,w)}=Y^w_{T(u,w)}\right]\\
&\geq\left(1-2^{-4k}\right)|\Sigma|^{-T}\sum_{\substack{u\in U\setminus\mathcal{D}_1,\\ w\in W\setminus\mathcal{D}_2,\\ (u,w)\notin\mathcal{D}\cup\mathcal{D}'}}\Pr_{P}\left[P_i=u\right]\cdot\Pr_{Q}\left[Q_i=w\right].
\end{aligned}$$

To bound the last term, we consider four partial sums. Clearly,

$$\sum_{(u,w)\in U\times W}\Pr_{P}\left[P_i=u\right]\cdot\Pr_{Q}\left[Q_i=w\right]=1.$$

By Equation (14),

$$\sum_{u\in\mathcal{D}_1}\Pr_{P}\left[P_i=u\right]\leq\frac{I}{2^{4k}},$$

and by Equation (15),

$$\sum_{w\in\mathcal{D}_2}\Pr_{Q}\left[Q_i=w\right]\leq\frac{I}{2^{4k}}.$$

By Lemma 15 (stated and proved below), and Equations (16) and (18),

$$\sum_{(u,w)\in\mathcal{D}\cup\mathcal{D}'}\Pr_{P}\left[P_i=u\right]\cdot\Pr_{Q}\left[Q_i=w\right]\leq O(1/\gamma^4)\cdot 2^{-2k}.$$

Therefore,

$$\begin{aligned}\Pr_{P,Q}\left[(P,Q)\in\Lambda^{t,s,z}\right] &\geq \left(1-2^{-4k}\right)|\Sigma|^{-T}\left(1-\frac{I}{2^{4k}}-\frac{I}{2^{4k}}-O(1/\gamma^4)\cdot 2^{-2k}\right)\\ &= |\Sigma|^{-T}\left(1-O(2^{-2k}/\gamma^4)\right).\end{aligned}$$

Finally, note that for every $x,y\in\Psi^z$, such that $x_i=u$ and $y_i=w$, the following holds: $(x,y)\in\mathrm{supp}(\mu_i)$ if and only if $x$ and $y$ agree on the subtrees rooted at vertices in $T(u,w)$ (these are the non-noisy subtrees). Therefore,

$$\Pr_{(x,y)\in_R\Psi^z\times\Psi^z}\left[(x,y)\in\mathrm{supp}(\mu_i)\right]=|\Sigma|^{-T},$$

and the assertion follows. □

**Lemma 15.** *Let $(t,s,z)\in\mathcal{G}_i$. Let $U=W=\{0,1\}^{V_i}$. Let $\mathcal{D}\subseteq U\times W$ be such that $\frac{|\mathcal{D}|}{|U|\cdot|W|}\leq 2^{-4k+1}$. It holds that*

$$C^{t,s,z}:=\sum_{(u,w)\in\mathcal{D}}\Pr\left[X_i^{t,s,z}=u\right]\cdot\Pr\left[Y_i^{t,s,z}=w\right]\leq O(1/\gamma^4)\cdot 2^{-2k}.$$

*Proof.* First, observe that $C^{t,s,z}\cdot|R^{t,s,z}|$ is exactly the number of input pairs $(x,y)\in R^{t,s,z}$ with $(x_i,y_i)\in\mathcal{D}$. Consider the expression

$$C^{t,z}:=\sum_{(u,w)\in\mathcal{D}}\Pr\left[X_i^{t,z}=u\right]\cdot\Pr\left[Y_i^{t,z}=w\right].$$

Again, $C^{t,z}\cdot|R^{t,z}|$ is exactly the number of input pairs $(x,y)\in R^{t,z}$ with $(x_i,y_i)\in\mathcal{D}$. Since $R^{t,z}$ contains $R^{t,s,z}$, and by the second property of a good tuple $(i,z,t,s)$ (i.e., $\gamma$-largeness), it holds that

$$C^{t,s,z}\leq\frac{|R^{t,z}|}{|R^{t,s,z}|}C^{t,z}\leq O(1/\gamma^4)\cdot C^{t,z}. \tag{19}$$

Therefore, in order to bound $C^{t,s,z}$, it suffices to bound $C^{t,z}$.

By the third and fourth properties of a good tuple $(i,z,t,s)$, we have $\mathbf{I}\left(X^{t,z}\right),\mathbf{I}\left(Y^{t,z}\right)\leq 2\log(m)$, which means that $|R^{t,z}|\geq\frac{1}{m^4}|\Psi^z\times\Psi^z|$. This implies that for every set $L\subseteq\Psi^z$, the probability that $X^{t,z}$ is in $L$ is at most $m^4$ the probability that a uniformly distributed variable over $\Psi^z$ obtains a value in $L$. In particular, for every $u\in U$,

$$\Pr\left[X_i^{t,z}=u\right]\leq\frac{m^4}{|U|}. \tag{20}$$

Similarly, for $w\in W$,

$$\Pr\left[Y_i^{t,z}=w\right]\leq\frac{m^4}{|W|}. \tag{21}$$

Define

$$U' = \left\{ u \in U \ : \ \Pr\left[X_i^{t,z} = u\right] \geq \frac{2}{|U|} \right\},$$

$$W' = \left\{ w \in W \ : \ \Pr\left[Y_i^{t,z} = w\right] \geq \frac{2}{|W|} \right\}.$$

By the fifth and sixth properties of a good tuple $(i, z, t, s)$, we have

$$\mathbf{I}\left(X_i^{t,z}\right) \leq \frac{m}{\gamma^2 c},$$

$$\mathbf{I}\left(Y_i^{t,z}\right) \leq \frac{m}{\gamma^2 c}.$$

Using Lemma 5.12 in [KR13] (stated for convenience in Appendix A, Lemma 21) it holds that

$$\Pr\left[X_i^{t,z} \in U'\right] < 5 \cdot \left(\frac{m}{\gamma^2 c}\right)^{0.1}, \tag{22}$$

Similarly,

$$\Pr\left[Y_i^{t,z} \in W'\right] < 5 \cdot \left(\frac{m}{\gamma^2 c}\right)^{0.1}. \tag{23}$$

The expression $C^{t,z}$ is a sum over pairs $(u, w) \in \mathcal{D}$. We bound $C^{t,z}$ by a sum of three partial sums, and work on each partial sum separately. The first partial sum is over pairs $(u, w) \in U \times W$ with $u \in U'$, the second is over pairs $(u, w) \in U \times W$ with $w \in W'$, the third is over pairs $(u, w) \in \mathcal{D}$ with $u \notin U'$ and $w \notin W'$.

We bound the first partial sum as follows. We use Equation (21) for the first step, and Equation (22) for the third.

$$\begin{aligned}
&\sum_{\substack{(u,w)\in U\times W \\ u\in U'}} \Pr\left[X_i^{t,z} = u\right] \cdot \Pr\left[Y_i^{t,z} = w\right] \\
&\leq \frac{m^4}{|W|} \sum_{\substack{(u,w)\in U\times W \\ u\in U'}} \Pr\left[X_i^{t,z} = u\right] \\
&= m^4 \cdot \sum_{u\in U'} \Pr\left[X_i^{t,z} = u\right] \\
&\leq 5m^4 \cdot \left(\frac{m}{\gamma^2 c}\right)^{0.1} \leq c^{-0.05}.
\end{aligned}$$

The second partial sum is bounded in a similar way. We bound the third partial sum using the bound that we have on the size of $\mathcal{D}$,

$$\sum_{\substack{(u,w)\in\mathcal{D} \\ u\notin U', w\notin W'}} \Pr\left[X_i^{t,z} = u\right] \cdot \Pr\left[Y_i^{t,z} = w\right] \leq |\mathcal{D}| \cdot \frac{2}{|U|} \cdot \frac{2}{|W|} < 2^{-3k}.$$

We conclude that $C^{t,z} \leq 2^{-2k}$. Using Equation (19),

$$C^{t,s,z} \leq O(1/\gamma^4) \cdot 2^{-2k}.$$

□

# 7 Bounding Information Cost by Tree Divergence Cost

In this section we give a general tool that can be used to upper bound the information cost of a protocol $\pi$, using the notion of a divergence cost of a tree. This notion is implicit in [BBCR10] and was formally defined in [BR11].

Let $\pi$ be a communication protocol between two players. We assume that the first player has the private input $x$ and the second player has the private input $y$, where $(x, y)$ were chosen according to some joint distribution $\mu$. In this section, we assume without loss of generality that $\pi$ does not use public randomness (but may use private randomness), as for the purpose of upper bounding the information cost, the public randomness can always be replaced by private randomness. We also assume, without loss of generality, that the players take alternating turns sending bits to each other. That is, in odd rounds, the first player sends a bit to the second player, and in even rounds the second player sends a bit to the first player (if this is not the case, we can add dummy rounds that do not change the information cost).

We denote by $\mathcal{T}_\pi$ the binary tree associated with the communication protocol $\pi$. That is, every vertex $v$ of $\mathcal{T}_\pi$ corresponds to a possible transcript of $\pi$, and the two edges going out of $v$ are labeled by 0 and 1, corresponding to the next bit to be transmitted. We think of the first player as owning the vertices in odd layers of $\mathcal{T}_\pi$ (where the root is in layer 1), and of the second player as owning the vertices in even layers of $\mathcal{T}_\pi$. When the protocol $\pi$ reaches a non-leaf vertex $v$, the player who owns $v$ sends a bit to the other player.

Every input pair $(x, y)$ for the protocol $\pi$ induces a distribution $P_v = (p_v, 1-p_v)$ for every non-leaf vertex $v$ of the tree $\mathcal{T}_\pi$, where $p_v$ is the probability that the next bit transmitted by the protocol $\pi$ on the vertex $v$ and inputs $x, y$ is 0. We think of $P_v$ as a distribution over the two children of the vertex $v$. Observe that the player who owns $v$ knows $P_v$. Given the binary tree $\mathcal{T}_\pi$ and the distributions $P_v$ for every non-leaf vertex $v$ of $\mathcal{T}_\pi$, where for each $v$ the player who owns $v$ knows $P_v$, we can assume without loss of generality that the protocol $\pi$ operates as follows: Starting from the root until reaching a leaf, at every vertex $v$, the player who owns $v$ samples a bit according to $P_v$ and sends this bit to the other player. Both players continue to the child of $v$ that is indicated by the communicated bit.

Assume that for every non-leaf vertex $v$ of $\mathcal{T}_\pi$, we have an additional distribution $Q_v = (q_v, 1 - q_v)$ that is known to the player who doesn't own $v$. We think of every $P_v$ as the "correct" distribution over the two children of $v$. This distribution is known to the player who owns $v$. We think of $Q_v$ as an estimation of $P_v$, based on the knowledge of the player who doesn't own $v$. For the rest of the section, we think of $\mathcal{T}_\pi$ as the tree $\mathcal{T}_\pi$ together

with the distributions $P_v$ and $Q_v$, for every non-leaf vertex $v$ in the tree $\mathcal{T}_\pi$.

To upper bound the information cost of a protocol $\pi$ it is convenient to use the notion of divergence cost of a tree [BBCR10, BR11].

**Definition 4** (**Divergence Cost** [BBCR10, BR11])**.** *Consider a binary tree $\mathcal{T}$, whose root is $r$, and distributions $P_v = (p_v, 1-p_v), Q_v = (q_v, 1-q_v)$ for every non-leaf vertex $v$ in the tree. We think of $P_v$ and $Q_v$ as distributions over the two children of the vertex $v$. We define the* divergence cost *of the tree $\mathcal{T}$ recursively, as follows. $\mathbf{D}(\mathcal{T}) = 0$ if the tree has depth 0, otherwise,*

$$\mathbf{D}(\mathcal{T}) = \mathbf{D}(P_r \| Q_r) + \mathop{\mathbf{E}}_{v \sim P_r}[\mathbf{D}(\mathcal{T}_v)], \tag{24}$$

*where for every vertex $v$, $\mathcal{T}_v$ is the subtree of $\mathcal{T}$ whose root is $v$.*

*An equivalent definition of the divergence cost of $\mathcal{T}$ is obtained by following the recursion in Equation (24) and is given by the following equation:*

$$\mathbf{D}(\mathcal{T}) = \sum_{v \in V} \tilde{p}_v \cdot \mathbf{D}(P_v \| Q_v), \tag{25}$$

*where $V$ is the vertex set of $\mathcal{T}$, and for a vertex $v \in V$, $\tilde{p}_v$ is the probability to reach $v$ by following the distributions $P_v$, starting from the root. Formally, if $v$ is the root of the tree $\mathcal{T}$, then $\tilde{p}_v = 1$, otherwise,*

$$\tilde{p}_v = \begin{cases} \tilde{p}_u \cdot p_u & \text{if } v \text{ is the left-hand child of } u \\ \tilde{p}_u \cdot (1-p_u) & \text{if } v \text{ is the right-hand child of } u. \end{cases}$$

Let $X$ be the input to the first player and $Y$ be the input to the second player. In the protocol $\pi$, the players use two private random strings and no public randomness. Denote the private random string of the first player by $R_1$, and the private random string of the second player by $R_2$. For a layer $d$ of $\mathcal{T}_\pi$, let $\Pi_d$ be the vertex in layer $d$ that the players reach during the execution of the protocol $\pi$, when the inputs are $(X, Y)$ and the private random strings are $R_1$ and $R_2$ (if $\pi$ ends before layer $d$, then $\Pi_d$ is undefined).

Let the tree $\mathcal{T}'_\pi$ be the same as $\mathcal{T}_\pi$, except that every distribution $Q_v$, for every non-leaf vertex $v$ in $\mathcal{T}_\pi$, is replaced with the distribution $Q'_v = (q'_v, 1-q'_v)$, where $q'_v$ is defined as follows: Let $d$ be the layer of $v$. If $v$ is owned by the first player, $q'_v$ is the function of $v, y$ and $r_2$, defined as

$$q'_v = \mathop{\mathbf{E}}_{X,R_1}[p_v | Y = y, R_2 = r_2, \Pi_d = v].$$

If $v$ is owned by the second player, $q'_v$ is the function of $v, x$ and $r_1$, defined as

$$q'_v = \mathop{\mathbf{E}}_{Y,R_2}[p_v | X = x, R_1 = r_1, \Pi_d = v].$$

We think of $Q'_v$ as the best estimation of the correct distribution $P_v$, based on the knowledge of the player who doesn't own $v$, whereas $Q_v$ is some estimation. Intuitively, $\mathbf{D}(P_v \| Q_v)$ is the information that the player who doesn't own $v$ learns on $P_v$ from the bit

sent during the protocol at the vertex $v$, assuming that she expects this bit to be distributed according to $Q_v$, whereas $\mathbf{D}(P_v\|Q'_v)$ is the information that she learns based on the best possible estimation of $P_v$. Therefore, intuitively, the divergence cost of $\mathcal{T}'_\pi$ is at most the divergence cost of $\mathcal{T}_\pi$, in expectation. This is formulated in the following lemma.

Observe that the protocol $\pi$ induces the distributions $P_v$ (known to the player who owns $v$) and $Q'_v$ (known to the player who doesn't own $v$), while the distribution $Q_v$ may be any distribution known to the player who doesn't own $v$.

**Lemma 16.** *For every protocol $\pi$ and distributions $Q_v$ known to the player who doesn't own $v$, as above, it holds that*

$$\mathbf{E}[\mathbf{D}(\mathcal{T}'_\pi)] \leq \mathbf{E}[\mathbf{D}(\mathcal{T}_\pi)],$$

*where the expectation is over the sampling of the inputs according to $\mu$ and over the randomness.*

*Proof.* By Equation (25),

$$\mathop{\mathbf{E}}_{X,Y,R_1,R_2}[\mathbf{D}(\mathcal{T}_\pi) - \mathbf{D}(\mathcal{T}'_\pi)] = \mathop{\mathbf{E}}_{X,Y,R_1,R_2}\left[\sum_v \tilde{p_v}\left(\mathbf{D}(P_v\|Q_v) - \mathbf{D}(P_v\|Q'_v)\right)\right],$$

where $\tilde{p}_v$ is as in Definition 4. We separate the sum on the vertices to layers and work on each layer separately. Fix a layer $d$ in the tree. Let $L_d$ be the set of vertices in layer $d$. To simplify notation, let $A$ denote $(X, R_1)$, let $B$ denote $(Y, R_2)$, and let $V$ denote $\Pi_d$. Then,

$$\mathop{\mathbf{E}}_{X,Y,R_1,R_2}\left[\sum_{v\in L_d} \tilde{p_v}\left(\mathbf{D}(P_v\|Q_v) - \mathbf{D}(P_v\|Q'_v)\right)\right] = \mathop{\mathbf{E}}_{A,B,V}\left[\mathbf{D}(P_V\|Q_V) - \mathbf{D}(P_V\|Q'_V)\right].$$

(Recall that $V$ is undefined when the protocol ends before layer $d$. In that case, for simplicity, we think of $P_V$, $Q_V$ and $Q'_V$ as all being equal, and hence $\mathbf{D}(P_V\|Q_V) = \mathbf{D}(P_V\|Q'_V) = 0$). By the definition of relative entropy,

$$\begin{aligned}
&\mathop{\mathbf{E}}_{A,B,V}[\mathbf{D}(P_V\|Q_V) - \mathbf{D}(P_V\|Q'_V)] \\
&= \mathop{\mathbf{E}}_{A,B,V}\left[p_V\left(\log\left(\frac{p_V}{q_V}\right) - \log\left(\frac{p_V}{q'_V}\right)\right) + (1-p_V)\left(\log\left(\frac{1-p_V}{1-q_V}\right) - \log\left(\frac{1-p_V}{1-q'_V}\right)\right)\right] \\
&= \mathop{\mathbf{E}}_{A,B,V}\left[p_V \log\left(\frac{q'_V}{q_V}\right) + (1-p_V)\log\left(\frac{1-q'_V}{1-q_V}\right)\right]. \qquad (26)
\end{aligned}$$

Assume that the first player owns the vertices in layer $d$. The case that the second player owns the vertices in layer $d$ is analogous. Consider the first summand in Equation (26). It holds that,

$$\mathop{\mathbf{E}}_{A,B,V}\left[p_V \log\left(\frac{q'_V}{q_V}\right)\right] = \mathop{\mathbf{E}}_{B,V}\left[\mathop{\mathbf{E}}_{A}\left[\left(p_V \log\left(\frac{q'_V}{q_V}\right)\right)\middle| B, V\right]\right].$$

By the definition of $q'_V$, for fixed $B, V$, it holds that $q'_V = \mathbf{E}_A[p_V|B, V]$. Since $q'_V$ and $q_V$

are functions of $B$ and $V$, when we condition on $B$ and $V$, $q'_V$ and $q_V$ are fixed. Therefore, conditioned on $B$ and $V$, the term $\log\left(\frac{q'_V}{q_V}\right)$ is independent of $A$. We get that,

$$\begin{aligned}\mathop{\mathbf{E}}_{B,V}\left[\mathop{\mathbf{E}}_{A}\left[\left(p_V \log\left(\frac{q'_V}{q_V}\right)\right)\middle| B,V\right]\right] &= \mathop{\mathbf{E}}_{B,V}\left[\mathop{\mathbf{E}}_{A}\left[p_V|B,V\right]\log\left(\frac{q'_V}{q_V}\right)\right] \\ &= \mathop{\mathbf{E}}_{B,V}\left[q'_V \log\left(\frac{q'_V}{q_V}\right)\right].\end{aligned}$$

In the same way, we get that the second summand in Equation (26) is

$$\mathop{\mathbf{E}}_{A,B,V}\left[(1-p_V)\log\left(\frac{1-q'_V}{1-q_V}\right)\right] = \mathop{\mathbf{E}}_{B,V}\left[(1-q'_V)\log\left(\frac{1-q'_V}{1-q_V}\right)\right].$$

Put together it holds that,

$$\mathop{\mathbf{E}}_{A,B,V}\left[\mathbf{D}(P_V\|Q_V) - \mathbf{D}(P_V\|Q'_V)\right] = \mathop{\mathbf{E}}_{B,V}\left[\mathbf{D}(Q'_V\|Q_V)\right] \geq 0,$$

since the divergence is non-negative. This is true for every layer $d$ in the tree. Therefore, summing over all layers, we get that

$$\mathop{\mathbf{E}}_{A,B}[\mathbf{D}(\mathcal{T}'_\pi)] \leq \mathop{\mathbf{E}}_{A,B}[\mathbf{D}(\mathcal{T}_\pi)].$$

□

The following lemma relates the information cost of $\pi$ to the expected divergence cost of $\mathcal{T}_\pi$. It was shown in [BR11] (see Lemma 5.3 therein) that $IC_\mu(\pi) = \mathbf{E}[\mathbf{D}(\mathcal{T}'_\pi)]$. Together with Lemma 16 we get:

**Lemma 17.** *For every protocol $\pi$ and distributions $Q_v$ known to the player who doesn't own $v$, as above, it holds that*

$$IC_\mu(\pi) = \mathbf{E}[\mathbf{D}(\mathcal{T}'_\pi)] \leq \mathbf{E}[\mathbf{D}(\mathcal{T}_\pi)],$$

*where the expectation is over the sampling of the inputs according to $\mu$ and over the randomness.*

# 8 Information Upper Bound

In this section we prove Theorem 2. Let $(x, y) \in \mathrm{supp}(\mu)$ be an input pair to the bursting noise game. Consider the following protocol $\pi'$ for the bursting noise game. Starting from the root until reaching a leaf, at every vertex $v$, if the first player owns $v$, she sends the bit $x_v$ with probability 0.9, and the bit $1 - x_v$ with probability 0.1. Similarly, if the second player owns $v$, she sends the bit $y_v$ with probability 0.9, and the bit $1-y_v$ with probability 0.1. Both players continue to the child of $v$ that is indicated by the communicated bit. When they reach a leaf they output that leaf. By the Chernoff bound, the probability that the players

output a leaf that is not typical with respect to the noisy multi-layer is at most $2^{-\Omega(w)}$. That is, the error probability of $\pi'$ is exponentially small in $k$.

The information cost of the protocol $\pi'$ is too large. The reason is that if the protocol reaches a non-typical vertex at the end of the noisy multi-layer (with respect to the noisy multi-layer), an event that occurs with probability exponentially small in $k$, then the rest of the protocol reveals to each player $\Omega\left((c-i)w\right)$ bits of information about the input of the other player, in expectation (as all the vertices below a non-typical vertex are noisy), and note that $\Omega\left((c-i)w\right)$ is double exponentially large (for almost all $i$). Thus, in expectation, the information revealed to each player about the input of the other player is double exponential in $k$.

For that reason, we consider a variant of the protocol $\pi'$, called $\pi$. Informally speaking, the protocol $\pi$ operates like $\pi'$ but aborts if too much information about the inputs is revealed. Recall that in every round of the protocol $\pi'$, the players are at a vertex $v$ of $\mathcal{T}$ and the player who owns $v$ sends a bit $b_v$ indicating one of $v$'s children. In the new protocol $\pi$, after receiving that bit, the receiving party sends a bit $a_v$ indicating whether they should abort the protocol, where $a_v = 1$ stands for abort and $a_v = 0$ stands for continue. If a bit $a_v = 1$, indicating an abort, was sent, the protocol terminates and both players output an arbitrary leaf of the tree $\mathcal{T}$. It remains to specify how the receiving party, without loss of generality the second player, decides whether to abort or continue, that is, how she determines the value of $a_v$.

To determine whether to abort, the second player considers the last $\ell = 2^{100k}$ vertices $v_1, \ldots, v_\ell$, reached by the protocol and owned by the first player, and the corresponding bits $b_{v_1}, \ldots, b_{v_\ell}$ that were sent by the first player (if less than $\ell$ bits were sent by the first player so far, then the second player does not abort). For every $j \in [\ell]$, the second player compares $b_{v_j}$ and $y_{v_j}$. The second player decides to abort and sends $a_v = 1$ if and only if less than $0.8\ell$ of these pairs are equal (otherwise the second player sends $a_v = 0$).

The following claim shows that the probability that $\pi$ aborts is exponentially small in $k$. If $\pi$ does not abort, it gives the same output as $\pi'$. We conclude that the error probability of $\pi$ is exponentially small in $k$.

**Claim 18.** *Let $(x, y) \in \mathrm{supp}(\mu)$ be an input pair to the bursting noise game. The protocol $\pi$ aborts with probability at most $2^{-10k}$ on the input $(x, y)$.*

*Proof.* Fix $(x, y) \in \mathrm{supp}(\mu_i)$ for some $i \in [c]$. Let $E$ be the event that the protocol $\pi$ reaches a non-typical vertex after multi-layer $i$ (with respect to multi-layer $i$). By the Chernoff bound, the event $E$ occurs with probability at most $2^{-100k}$, as $w = 2^{100}k$. Let $A$ be the event that the protocol $\pi$ aborts. Assume that $E$ does not occur. By the Chernoff bound, the probability of aborting after each round is at most $2^{-2^{50k}}$, as $\ell = 2^{100k}$ and since if $E$ does not occur then $x_v$ and $y_v$ can only differ for at most $w$ vertices reached by the protocol $\pi$. By the union bound, the probability of abort (conditioned on $\neg E$) is at most $cw \cdot 2^{-2^{50k}} < 2^{-100k}$. Therefore,

$$\Pr[A] \leq \Pr[E] + \Pr[A|\neg E] \leq 2 \cdot 2^{-100k}.$$

□

To upper bound the information cost of the protocol $\pi$ we will use Lemma 17. We denote by $\mathcal{T}_\pi$ the binary tree associated with the communication protocol $\pi$, as in Section 7. That is, every vertex $v$ of $\mathcal{T}_\pi$ corresponds to a possible transcript of $\pi$, and the two edges going out of $v$ are labeled by 0 and 1, corresponding to the next bit to be transmitted. The non-leaf vertices of the tree $\mathcal{T}_\pi$ have the following structure: Every non-leaf vertex $v$ in an odd layer of $\mathcal{T}_\pi$ corresponds to a non-leaf vertex of $\mathcal{T}$, the binary tree on which the bursting noise game is played. Since the correspondence is one-to-one, we refer to the vertex in $\mathcal{T}$ corresponding to $v$ also as $v$. The next bit to be transmitted by $\pi$ on the vertex $v$ is $b_v$. For a non-leaf vertex $v$ in an even layer of $\mathcal{T}_\pi$, the next bit to be transmitted by $\pi$ on the vertex $v$ is $a_v$.

As explained in Section 7, every input pair $(x, y) \in \mathrm{supp}(\mu)$ to the bursting noise game, induces a distribution $P_v = (p_v, 1 - p_v)$ for every non-leaf vertex $v$ of the tree $\mathcal{T}_\pi$, where $p_v$ is the probability that the next bit transmitted by the protocol $\pi$ on the vertex $v$ and inputs $x, y$ is 0. Namely, if $v$ is in an odd layer of $\mathcal{T}_\pi$ (and recall that in this case we think of $v$ as both a vertex of $\mathcal{T}_\pi$ and of $\mathcal{T}$), the distribution $P_v$ is the following: In the case that the first player owns $v$ in $\mathcal{T}$, if $x_v = 0$ then $P_v = (0.9, 0.1)$, and if $x_v = 1$ then $P_v = (0.1, 0.9)$. In the case that the second player owns $v$, if $y_v = 0$ then $P_v = (0.9, 0.1)$, and if $y_v = 1$ then $P_v = (0.1, 0.9)$. If $v$ is in an even layer of $\mathcal{T}_\pi$ then $P_v$ is $P_v = (0, 1)$ if the player sending $a_v$ decides to abort, and $P_v = (1, 0)$ if she decides to continue (note that given $x, y, v$, this decision is deterministic).

For every non-leaf vertex $v$ of $\mathcal{T}_\pi$, we define an additional distribution $Q_v = (q_v, 1 - q_v)$ (depending on the input $(x, y)$). We think of every $P_v$ as the "correct" distribution over the two children of $v$. This distribution is known to the player who sends the next bit on the vertex $v$. We think of $Q_v$ as an estimation of $P_v$, based on the knowledge of the player who doesn't send the next bit. For a vertex $v$ in an odd layer of $\mathcal{T}_\pi$ (and recall that in this case we think of $v$ as both a vertex of $\mathcal{T}_\pi$ and of $\mathcal{T}$), the distribution $Q_v$ is the following: In the case that the first player owns $v$ in $\mathcal{T}$, if $y_v = 0$ then $Q_v = (0.9, 0.1)$, and if $y_v = 1$ then $Q_v = (0.1, 0.9)$. In the case that the second player owns $v$, if $x_v = 0$ then $Q_v = (0.9, 0.1)$, and if $x_v = 1$ then $Q_v = (0.1, 0.9)$. If $v$ is in an even layer of $\mathcal{T}_\pi$ then $Q_v = (1 - \frac{1}{cw}, \frac{1}{cw})$.

For the rest of the section, we think of $\mathcal{T}_\pi$ as the tree $\mathcal{T}_\pi$ together with the distributions $P_v$ and $Q_v$, for every vertex $v$ in the tree $\mathcal{T}_\pi$.

**Proposition 19.** *It holds that*

$$\mathbf{D}(\mathcal{T}_\pi) = O(k).$$

*Proof.* Fix $(x, y) \in \mathrm{supp}(\mu_i)$ for some $i \in [c]$. By Equation (25),

$$\mathbf{D}(\mathcal{T}_\pi) = \sum_v \tilde{p}_v \cdot \mathbf{D}(P_v \| Q_v),$$

where $\tilde{p}_v$ is the probability that the protocol $\pi$ reaches the vertex $v$ on input $(x, y)$. We will bound the last sum separately for vertices $v$ in odd layers and for vertices $v$ in even layers.

We first sum over vertices in even layers. For every vertex $v$ in an even layer of $\mathcal{T}_\pi$, if $P_v = (0,1)$ (protocol aborts) we have $\mathbf{D}(P_v \| Q_v) = \log(cw)$, and if $P_v = (1,0)$ (protocol continues) we have $\mathbf{D}(P_v \| Q_v) = \log\left(\frac{1}{1-1/cw}\right) = \log\left(1 + \frac{1}{cw-1}\right) < \frac{2}{cw}$. By Claim 18, the probability that $\pi$ aborts is at most $2^{-10k}$. Therefore, the sum in Equation (25) taken over vertices in even layers is at most $cw \cdot \frac{2}{cw} + 2^{-10k} \cdot \log(cw) \leq 3$, as for each of the $cw$ even layers, the probability of reaching a vertex in this layer is at most 1.

We next sum over vertices in odd layers. Recall that each such vertex corresponds to a vertex in $\mathcal{T}$. Let $v$ be a vertex in an odd layer of $\mathcal{T}_\pi$. If $v$ corresponds to a non-noisy vertex in $\mathcal{T}$ we have $\mathbf{D}(P_v \| Q_v) = 0$, and if $v$ corresponds to a noisy vertex in $\mathcal{T}$ we have $\mathbf{D}(P_v \| Q_v) \leq 4$. Recall that $i$ is the noisy multi-layer. Then,

1. The vertices above multi-layer $i$ in $\mathcal{T}$ add nothing to the divergence cost.
2. Multi-layer $i$ of $\mathcal{T}$ adds $O(w)$ to the divergence cost.
3. If $i < c$: Let $v$ be the vertex in layer $i^* + w$ of $\mathcal{T}$ that the players reach during the execution of the protocol $\pi$. If $v$ is a typical vertex with respect to multi-layer $i$, the vertices below $v$ add nothing to the divergence cost. If $v$ is a non-typical vertex, the protocol aborts after at most $4\ell$ rounds in expectation. Since the probability that $v$ is a non-typical vertex with respect to multi-layer $i$ is at most $2^{-1000k}$ (as $w = 2^{100}k$), the expected divergence cost added by this case is at most $2^{-1000k} \cdot 4\ell \cdot 4 \leq 1$.

Together, the total divergence cost is $O(w) = O(k)$, as claimed. □

By Proposition 19 and Lemma 17 we get that $IC_\mu(\pi) \leq O(k)$.


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# A Information Theoretic Lemmas

**Lemma (Lemma 8 restated, Shearer-Like Inequality for Relative Entropy).** *Let $P, Q : \Omega_1 \times \cdots \times \Omega_M \to [0,1]$ be two distributions, such that $Q$ is a product distribution, i.e., for every $j \in [M]$, there exists $Q_j : \Omega_j \to [0,1]$, such that $Q(x_1, \ldots, x_M) = \prod_{j\in[M]} Q_j(x_j)$. Let $T = \{T_i\}_{i\in I}$ be a collection of subsets of $[M]$, such that each element of $[M]$ appears in at most $\frac{1}{K}$ fraction of the members of $T$. For $A \subseteq [M]$, let $P_A$ and $Q_A$ be the marginal distributions of $A$ in the distributions $P$ and $Q$ (respectively). Then,*

$$K \cdot \mathop{\mathbf{E}}_{i\in_R I}[\mathbf{D}(P_{T_i} \| Q_{T_i})] \leq \mathbf{D}(P\|Q).$$

*Proof.* For the proof of this lemma, it will be convenient to use the notion of conditional relative entropy using the notation of [CT06] (see Section 2.5 of [CT06]). For $A \subseteq [M]$, and $(x_1, \ldots, x_M) \in \Omega_1 \times \ldots \times \Omega_M$, we denote $x_A = \{x_j : j \in A\}$. In this notation, we need to prove that

$$K \cdot \mathop{\mathbf{E}}_{i \in_R I} \left[ \mathbf{D}\left(P(x_{T_i}) \| Q(x_{T_i})\right) \right] \leq \mathbf{D}\left(P(x_{[M]}) \| Q(x_{[M]})\right).$$

Define $x_{<j} = \{x_\ell : \ell < j\}$ and $x_{T_i,<j} = \{x_\ell : \ell \in T_i, \ell < j\}$. By the chain rule for relative entropy (see Section 2.5 in [CT06]),

$$\begin{aligned}
\mathbf{D}\left(P(x_{[M]}) \| Q(x_{[M]})\right) &= \sum_{j \in [M]} \mathbf{D}\left(P(x_j|x_{<j}) \| Q(x_j|x_{<j})\right), \\
\mathbf{D}\left(P(x_{T_i}) \| Q(x_{T_i})\right) &= \sum_{j \in T_i} \mathbf{D}\left(P(x_j|x_{T_i,<j}) \| Q(x_j|x_{T_i,<j})\right).
\end{aligned}$$

Since $Q$ is a product distribution, conditioning can only increase the relative entropy (see, for example, Lemma 2.5.3 in [Gra90]). In particular, for every $j \in T_i$ it holds that

$$\mathbf{D}\left(P(x_j|x_{T_i,<j}) \| Q(x_j|x_{T_i,<j})\right) \leq \mathbf{D}\left(P(x_j|x_{<j}) \| Q(x_j|x_{<j})\right).$$

Therefore,

$$\mathbf{D}\left(P(x_{T_i}) \| Q(x_{T_i})\right) \leq \sum_{j \in T_i} \mathbf{D}\left(P(x_j|x_{<j}) \| Q(x_j|x_{<j})\right).$$

Summing over all $i \in I$ we get that

$$\sum_{i \in I} \mathbf{D}\left(P(x_{T_i}) \| Q(x_{T_i})\right) \leq \sum_{i \in I} \sum_{j \in T_i} \mathbf{D}\left(P(x_j|x_{<j}) \| Q(x_j|x_{<j})\right). \tag{27}$$

For every $j \in [M]$, the term $\mathbf{D}\left(P(x_j|x_{<j}) \| Q(x_j|x_{<j})\right)$ appears on the right-hand side of Equation (27) at most $\frac{|I|}{K}$ times. Therefore,

$$\begin{aligned}
\sum_{i \in I} \mathbf{D}\left(P(x_{T_i}) \| Q(x_{T_i})\right) &\leq \frac{|I|}{K} \cdot \sum_{j \in [M]} \mathbf{D}\left(P(x_j|x_{<j}) \| Q(x_j|x_{<j})\right) \\
&= \frac{|I|}{K} \cdot \mathbf{D}\left(P(x_{[M]}) \| Q(x_{[M]})\right).
\end{aligned}$$

Dividing by $\frac{|I|}{K}$ we get that the claim holds. □

**Lemma 20 (Lemma 5.11 in [KR13]).** *Let $\mu : \Omega \to [0,1]$ be a distribution satisfying $I = \mathbf{I}(\mu) \leq 0.01$. Let $\mathcal{A} \subseteq \Omega$ be the set of elements with $\mu(x) < \frac{1}{|\Omega|}$. Denote*

$$I^{neg}(\mu) = -\sum_{x \in \mathcal{A}} \mu(x) \log(|\Omega| \mu(x)).$$

*Then,*

$$I^{neg}(\mu) < 4 I^{0.25} \log\left(\tfrac{1}{I^{0.25}}\right) < 4 I^{0.1}.$$

**Lemma 21** (**Lemma 5.12 in [KR13]**). *Let $\mu : \Omega \to [0,1]$ be a distribution satisfying $I = \mathbf{I}(\mu) \leq 0.01$. Let $\mathcal{A} \subseteq \Omega$ be the set of elements with $\mu(x) \geq \frac{2}{|\Omega|}$. Then,*

$$\mu(\mathcal{A}) < 4I^{0.25} \log\left(\tfrac{1}{I^{0.25}}\right) + I < 5I^{0.1}.$$