

Practical Locally Private Heavy Hitters

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Abstract

We present new practical local differentially private heavy hitters algorithms achieving optimal or near-optimal worst-case error and running time – **TreeHist** and **Bitstogram**. In both algorithms, server running time is $\tilde{O}(n)$ and user running time is $\tilde{O}(1)$, hence improving on the prior state-of-the-art result of Bassily and Smith [STOC 2015] requiring $O(n^{5/2})$ server time and $O(n^{3/2})$ user time. With a typically large number of participants in local algorithms (n in the millions), this reduction in time complexity, in particular at the user side, is crucial for making locally private heavy hitters algorithms usable in practice. We implemented Algorithm **TreeHist** to verify our theoretical analysis and compared its performance with the performance of Google’s RAPPOR code.

Keywords: Differential privacy, local differential privacy, heavy hitters, histograms, sketching.

1. Introduction

We revisit the problem of computing heavy hitters with local differential privacy. Such computations have already been implemented to provide organizations with valuable information about their user base while providing users with the strong the strong guarantee that their privacy would be preserved even if the organization is subpoenaed for the entire information seen during an execution. Two prominent examples are Google’s use of RAPPOR in the Chrome browser (Erlingsson et al., 2014) and Apple’s use of differential privacy in iOS-10 (Thakurta et al., 2017). These tools are used for learning new words typed by users and identifying frequently used emojis and frequently accessed websites.

Differential privacy in the local model. Differential privacy (Dwork et al., 2006) provides a framework for rigorously analyzing privacy risk and hence can help organization

mitigate users’ privacy concerns as it ensures that what is learned about any individual user would be (almost) the same whether the user’s information is used as input to an analysis or not.

Differentially private algorithms work in two main modalities: curator and local. The *curator model* assumes a trusted centralized curator that collects all the personal information and then analyzes it. In contrast, the *local model* does not involve a central repository. Instead, each piece of personal information is randomized by its provider to protect privacy even if all information provided to the analysis is revealed. Holding a central repository of personal information can become a liability to organizations in the face of security breaches, employee misconduct, subpoenas, etc. This makes locally private computations attractive for implementation. Indeed in the last few years Google and Apple have announced successful deployments of locally private analyses (Erlingsson et al., 2014; Thakurta et al., 2017).

Challenges of the local model. A disadvantage of the local model is that it requires introducing noise at a significantly higher level than what is required in the curator model. Furthermore, some tasks which are possible in the curator model are impossible in the local model (Dwork et al., 2006; Kasiviswanathan et al., 2011; Chan et al., 2012). To see the effect of noise, consider estimating the number of HIV positives in a given population of n participants. In the curated model, it suffices to add Laplace noise of magnitude $O_\epsilon(1)$, i.e., independent of n (Dwork et al., 2006). In contrast, a lower bound of $\Omega_\epsilon(\sqrt{n})$ is known for the local model (Chan et al., 2012). A higher noise level implies that the number of participants n needs to be large (maybe in the millions for a reasonable choice of ϵ). An important consequence is that, to be practical, locally private algorithms must exhibit low time, space, and communication complexity, especially at the user side. This is the problem addressed in our work.

Heavy hitters and histograms in the local model. Assume each of n users holds an element x_i taken from a domain of size d . A histogram of this data lists (an estimate of) the multiplicity of each domain element. When d is large, a succinct representation of the histogram is desired, either in the form of a *frequency oracle* – a data structure allowing to approximate the multiplicity of any domain element – or *heavy hitters* – listing the multiplicity of most frequent domain elements, implicitly considering the multiplicity of other domain elements as zero. The problem of computing histograms with differential privacy has attracted significant attention both in the curator model (Dwork et al., 2006; Beimel et al., 2016; Bun et al., 2015) and the local model (Hsu et al., 2012; Erlingsson et al., 2014; Bassily and Smith, 2015). Of relevance is the work in (Mishra and Sandler, 2006).

We briefly report on the state of the art heavy hitters algorithms of Bassily and Smith (2015) and Thakurta et al. (2017), which are most relevant for the current work. Bassily and Smith provide matching lower and upper bounds of $\Theta(\sqrt{n \log(d)}/\epsilon)$ on the worst-case error of local heavy hitters algorithms. Their local algorithm exhibits optimal communication but a rather high time complexity: Server running time is $O(n^{5/2})$ and, crucially, user running time is $O(n^{3/2})$ – complexity that severely hampers the practicality of this algorithm.¹ The construction by Thakurta et al. is a heuristic with no bounds on server running time and

1. If public randomness or efficient Private Information Retrieval is assumed, then the user running time in (Bassily and Smith, 2015) can be improved to $O(n)$.

accuracy (we refer to the maximal difference between the true and estimated multiplicities of a data element as the *error* or the *accuracy* of the protocol).² User computation time is $\tilde{O}(1)$, a significant improvement over (Bassily and Smith, 2015).

Our contributions. The focus of this work is on the design of locally private heavy hitters algorithms with near optimal error, keeping time, space, and communication complexity minimal. We provide two new constructions of heavy hitters algorithms **TreeHist** and **Bitstogram**. These algorithms achieve similar performance but apply different techniques. We implemented Algorithm **TreeHist** and provide measurements in comparison with RAPPOR (Erlingsson et al., 2014) (the only currently available implementation for local histograms). Our measurements are performed with a setting that is favorable to RAPPOR (i.e., a small input domain), yet they indicate that Algorithm **TreeHist** performs better than RAPPOR in terms of noise level.

Table 1 details various performance parameters of **TreeHist** and **Bitstogram**, and the reader can check in the table that these are similar up to small factors which we ignore for the rest of this paragraph. Comparing with (Bassily and Smith, 2015), we improve time complexity both at the server (reduced from $O(n^{5/2})$ to $\tilde{O}(n)$) and at the user (reduced from $O(n^{3/2})$ to $O(\max(\log n, \log d)^2)$). Comparing with (Thakurta et al., 2017), we get provable bounds on the server running time and worst-case error. Note that Algorithm **Bitstogram** achieves optimal worst-case error whereas Algorithm **TreeHist** is almost optimal, by a factor of $\sqrt{\log(n)}$.

Performance metric	TreeHist	Bitstogram
Server time (modular multiplications)	$\tilde{O}(n)$	$\tilde{O}(n)$
User time (modular multiplications)	$O(\max(\log n, \log d)^2)$	$O(\max(\log n, \log d)^2)$
Server processing memory	$\tilde{O}(\sqrt{n})$	$\tilde{O}(\sqrt{n})$
User memory	$O(\max(\log d, \log n))$	$O(\max(\log d, \log n))$
Communication/user	$O(1)$	$O(1)$
Worst-case Error	$O(\sqrt{n \log(n) \log(d)})$	$O(\sqrt{n \log(d)})$

Table 1: Performance of our protocols. Dependency on the privacy parameter ϵ and failure probability β is omitted.

Elements of the constructions. Main details of our constructions are presented in sections 3.1 and 3.2. These are complemented with the detailed descriptions and analyses in the appendix. Both our algorithms make use of frequency oracles – data structures that allow estimating various counts.

Algorithm **TreeHist** identifies heavy-hitters and estimates their frequencies by scanning the levels of a binary prefix tree whose leaves correspond to dictionary items. The recovery of the heavy hitters is in a bit-by-bit manner. As the algorithm progresses down the tree it prunes all the nodes that *cannot* be prefixes of heavy hitters, hence leaving $\tilde{O}(\sqrt{n})$ nodes in every depth. This is done by making queries to a frequency oracle. Once the algorithm reaches the final level of the tree it identifies the list of heavy hitters. It then invokes the

2. The underlying construction in (Thakurta et al., 2017) is of a frequency oracle.

frequency oracle once more on those particular items to obtain more accurate estimates for their frequencies.

Algorithm `Bitstogram` hashes the input domain into a domain of size roughly \sqrt{n} . The observation behind this algorithm is that if a heavy hitter x does not collide with other heavy hitters then $(h(x), x_i)$ would have a significantly higher count than $(h(x), \neg x_i)$ for every i , where x_i is the i th bit of x . This allows recovering all bits of x in parallel given an appropriate frequency oracle.

Followup work. Our work presents locally-private heavy-hitters algorithms with error rates that depend optimally on the number of users, the size of the domain, and the privacy parameter. Following our work, Bun et al. (2018) presented an improved protocol with error rate that depend optimally also on the failure probability. Their protocol is a refined version of our `Bitstogram` protocol.

2. Background and Preliminaries

2.1. Definitions and Notation

Dictionary and user items: Let $\mathcal{V} = [d]$. We consider a set of n users, where each user $i \in [n]$ holds an item $v_i \in \mathcal{V}$. We will use v_i to refer to the binary representation of v_i when it is clear from the context.

Item frequency (multiplicity): For each item $v \in \mathcal{V}$, we define the frequency $f(v)$ of v as the number of users holding v , namely,

$$f(v) \triangleq |\{i \in [n] : v_i = v\}|.$$

Frequency oracle: A frequency oracle is a data structure together with an algorithm that, for any given $v \in \mathcal{V}$, allows computing an estimate $\hat{f}(v)$ of the frequency $f(v)$.

Heavy hitters (succinct histogram): A succinct histogram is a data structure that provides a (short) list of items $(\hat{v}_1, \dots, \hat{v}_k)$, called the *heavy hitters*, together with estimates for their frequencies $(\hat{f}(\hat{v}_j) : j \in [k])$. The frequencies of the items not in the list are implicitly estimated as $\hat{f}(v) = 0$. We measure the error in a succinct histogram by the ℓ_∞ distance between the estimated and true frequencies, $\max_{v \in [d]} |\hat{f}(v) - f(v)|$. We will also consider the maximum error restricted to the items in the list $(\hat{v}_1, \dots, \hat{v}_k)$, that is, $\max_{j \in [k]} |\hat{f}(\hat{v}_j) - f(\hat{v}_j)|$. If a succinct histogram aims to provide ℓ_∞ error η , the list does not need to contain more than $O(1/\eta)$ items since items with estimated frequencies below η may be omitted from the list.

2.2. Local Differential Privacy

In the local model, an algorithm $\mathcal{A} : \mathcal{V} \rightarrow \mathcal{Z}$ accesses the database $\mathbf{v} = (v_1, \dots, v_n) \in \mathcal{V}^n$ only via an oracle that, given index $i \in [n]$, runs a randomized algorithm (local randomizer) $\mathcal{R} : \mathcal{V} \rightarrow \tilde{\mathcal{Z}}$ on input v_i and returns $\mathcal{R}(v_i)$ to \mathcal{A} .

Definition 2.1 (Local differential privacy Dwork et al. (2006); Evfimievski et al. (2003); Kasiviswanathan et al. (2011)) *An algorithm satisfies ϵ -local differential privacy (LDP) if it accesses the database $\mathbf{v} = (v_1, \dots, v_n) \in \mathcal{V}^n$ only via invocations of a local*

randomizer \mathcal{R} and if for all $i \in [n]$, if $\mathcal{R}^{(1)}, \dots, \mathcal{R}^{(k)}$ denote the algorithm's invocations of \mathcal{R} on the data sample v_i , then the algorithm $\mathcal{A}(\cdot) \triangleq (\mathcal{R}^{(1)}(\cdot), \mathcal{R}^{(2)}(\cdot), \dots, \mathcal{R}^{(k)}(\cdot))$ is ϵ -differentially private. That is, if for any pair of inputs \mathbf{v}, \mathbf{v}' that differ on a single input, and for all $\mathcal{S} \subseteq \text{Range}(\mathcal{A})$,

$$\Pr[\mathcal{A}(v) \in \mathcal{S}] \leq e^\epsilon \cdot \Pr[\mathcal{A}(v') \in \mathcal{S}].$$

2.3. Count Sketch and Hadamard Transform

Count sketch (Charikar et al., 2002) together with the Hadamard transform form the basis of our differentially private construction outlined in Section 3.1 and discussed in detail in Section 5.

Count sketch (Charikar et al., 2002) is a sketching algorithm for finding frequent elements in a data stream. Let $\mathcal{V} = [d]$ be a domain of data elements, and $\mathbf{v} = (v_1, \dots, v_n)$ be a stream of data elements. Count sketch ensures that for any given $v \in \mathcal{V}$, using a data structure of size $m = O\left(\frac{n \log(1/\beta)}{k}\right)$ one can ensure that with probability at least $1 - \beta$ the estimated frequency of v is within k of the true frequency, and the estimate is *unbiased*. The algorithm works as follows: First, pick t pairs of hash functions $(h_i : \mathcal{V} \rightarrow [m], g_i : \mathcal{V} \rightarrow \{-1, 1\})$, and set a matrix $\mathbf{M} = \{0\}^{t \times m}$. Second, with every data sample v_i , populate the matrix as follows: $\forall j \in [t], \mathbf{M}[j, h_j(v_i)] \leftarrow \mathbf{M}[j, h_j(v_i)] + g_j(v_i)$. Finally, to estimate the frequency of an element $v \in \mathcal{V}$, compute $\text{median} \{\mathbf{M}[1, h_1(v)] \cdot g_1(v), \dots, \mathbf{M}[t, h_t(v)] \cdot g_t(v)\}$.

Hadamard transform: We use Hadamard transform followed by sampling in order to compress our data transmission from the client to the server and to reduce the space requirements of our protocol. Hadamard transform of a vector $w \in \mathbb{R}^m$ is obtained via multiplying with the Hadamard transform matrix $\mathbf{H}_m \in \{-\frac{1}{\sqrt{m}}, \frac{1}{\sqrt{m}}\}^{m \times m}$ defined recursively as $\mathbf{H}_m = \frac{1}{\sqrt{2}} \begin{bmatrix} \mathbf{H}_{m/2} & \mathbf{H}_{m/2} \\ \mathbf{H}_{m/2} & -\mathbf{H}_{m/2} \end{bmatrix}$, and $\mathbf{H}_1 = [1]$. Two main properties of Hadamard transform we use are: i) it is a dense basis transformation, i.e. the columns of the Hadamard matrix form a basis, and each entry of $\sqrt{m} \cdot \mathbf{H}_m$ is in $\{-1, 1\}$, and ii) any entry (i, j) in $\sqrt{m} \cdot \mathbf{H}_m$ can be computed in $O(\log m)$ time.

2.4. Error correction codes

We will use error correction codes in order to reduce the error of (some of) our constructions outlined in Section 3.2 and discussed in detail in Section 6.

Definition 2.2 A binary (n, k) -code is a pair of mappings (Enc, Dec) where $\text{Enc} : \{0, 1\}^k \rightarrow \{0, 1\}^n$, and $\text{Dec} : \{0, 1\}^n \rightarrow \{0, 1\}^k$. The code is ζ -decodable if for every $x \in \{0, 1\}^k$ and every $y \in \{0, 1\}^n$ whose Hamming distance to $\text{Enc}(x)$ is at most ζn we have that $\text{Dec}(y) = x$.

For any constant $0 < \zeta < 1/4$, there is a construction of a ζ -decodable (n, k) -code, where $n = O(k)$, and furthermore, Enc and Dec run in time $O(n)$. See, e.g., (Guruswami, 2001).

2.5. Tools from Probability

2.5.1. PAIRWISE k -WISE INDEPENDENCE

The definitions given here are taken from (Vadhan, 2012).

Definition 1 *A family (i.e., multiset) of functions $\mathcal{H} = \{h : [N] \rightarrow [M]\}$ is pairwise independent if the following two conditions hold when H is a function chosen uniformly at random from \mathcal{H} :*

1. $\forall x \in [N]$, the random variable $H(x)$ is uniformly distributed in $[M]$.
2. $\forall x_1 \neq x_2 \in [N]$, the random variables $H(x_1)$ and $H(x_2)$ are independent.

We will use the following tail bound on sums of k -wise independent random variables.

Lemma 2.3 (Bellare and Rompel (1994)) *Let $\lambda \geq 6$ be an even integer. Suppose X_1, \dots, X_n are k -wise independent random variables taking values in $[0, 1]$. Let $X = X_1 + \dots + X_n$ and $\mu = \mathbb{E}[X]$, and let $\alpha > 0$. Then,*

$$\Pr[|X - \mu| \geq \alpha] \leq \left(\frac{nk}{\alpha^2}\right)^{k/2}.$$

2.5.2. THE POISSON APPROXIMATION

We will use the following useful facts about the Poisson approximation. When throwing n balls into R bins, the distribution of the number of balls in a given bin is $\text{Bin}(n, 1/R)$. As the Poisson distribution is the limit distribution of the binomial distribution, the distribution of the number of balls in a given bin is approximately $\text{Pois}(n/R)$. In fact, in some cases we could approximate the *joint distribution* of the number of balls in all the bins by assuming the load at each bin is an independent Poisson random variable with mean n/R .

Theorem 2.4 (e.g., Mitzenmacher and Upfal (2005)) *Suppose that n balls are thrown into R bins independently and uniformly at random, and let X_i be the number of balls in the i^{th} bin, where $1 \leq i \leq R$. Let Y_1, \dots, Y_R be independent Poisson random variables with mean n/R . Let $f(x_1, \dots, x_R)$ be a nonnegative function. Then,*

$$\mathbb{E}[f(X_1, \dots, X_R)] \leq e^{\sqrt{n}} \mathbb{E}[f(Y_1, \dots, Y_R)].$$

In particular, the theorem states that any event that takes place with probability p in the Poisson case, takes place with probability at most $pe^{\sqrt{n}}$ in the exact case (this follows by letting f be the indicator function of that event).

We will also use the following bounds for the tail probabilities of a Poisson random variable:

Theorem 2.5 (Alon and Spencer (1992)) *Let X have Poisson distribution with mean μ . For $0 \leq \alpha \leq 1$,*

$$\begin{aligned} \Pr[X \leq \mu(1 - \alpha)] &\leq e^{-\alpha^2 \mu/2} \\ \Pr[X \geq \mu(1 + \alpha)] &\leq e^{-\alpha^2 \mu/2}. \end{aligned}$$

3. Our Algorithms

In this section we give an informal description of our algorithms, and highlight some of the ideas behind our constructions.

3.1. The TreeHist Protocol

We briefly give an overview of our construction that is based on a compressed, noisy version of the count sketch. To maintain clarity of the main ideas, we give here a high-level description of our construction. We refer to Section 5 for a detailed description of this construction.

We first introduce some objects and public parameters that will be used in the construction:

Prefixes: For a binary string v , we will use $v[1 : \ell]$ to denote the ℓ -bit prefix of v . Let $\bar{\mathcal{V}} = \{v \in \{0, 1\}^\ell \text{ for some } \ell \in [\log d]\}$. Note that elements of $\bar{\mathcal{V}}$ arranged in a binary prefix tree of depth $\log d$, where the nodes at level ℓ of the tree represent all binary strings of length ℓ . The items of the dictionary \mathcal{V} represent the bottommost level of that tree. Note that $|\bar{\mathcal{V}}| = 2d$. We will use \perp to denote an empty string. For a $v \in \{0, 1\}^\ell$ and $b \in \{0, 1\}$, let $v\|b$ denote the $\ell + 1$ -bit string resulting from appending the bit b to v . For a binary string v , we define $\text{Child}(v) \triangleq \{v\|0, v\|1\}$, that is, the set containing the two children of v in the prefix tree. Similarly, for a set of strings \mathcal{U} , we define $\text{ChildSet}(\mathcal{U}) \triangleq \{v : v \in \text{Child}(u) \text{ for some } u \in \mathcal{U}\}$.

Hashes: Let t, m be positive integers to be specified later. We will consider a set of t pairs of hash functions $\{(h_1, g_1), \dots, (h_t, g_t)\}$, where for each $i \in [t]$, $h_i : \bar{\mathcal{V}} \rightarrow [m]$ and $g_i : \bar{\mathcal{V}} \rightarrow \{-1, +1\}$ are independently and uniformly chosen pairwise independent hash functions.

Basis matrix: Let $\mathbf{W} \in \{-1, +1\}^{m \times m}$ be $\sqrt{m} \cdot \mathbf{H}_m$ where \mathbf{H}_m is the Hadamard transform matrix of size m . As will be shown later, we will be making operations over the entries of this matrix. It is important to note that we do not need to store this matrix. The value of any entry in this matrix can be computed in $O(\log m)$ bit operations given the (row, column) index of that entry. In particular, suppose we want to compute the value of the entry $W_{i,j}$ located at the i -th row and j -th column. Let $(i_0, i_1, \dots, i_{\log m-1})$ and $(j_0, j_1, \dots, j_{\log m-1})$ denote the bit representation of i and j , respectively. Then, $W_{i,j} = (-1)^{\sum_{k=0}^{\log m-1} i_k j_k}$.

Global parameters: The total number of users n , the size of the Hadamard matrix m , the number of hash pairs t , the privacy parameter ϵ , the confidence parameter β , and the hash functions $\{(h_1, g_1), \dots, (h_t, g_t)\}$ are assumed to be public information. We set $t = O(\log(n/\beta))$ and $m = O\left(\sqrt{\frac{n}{\log(n/\beta)}}\right)$.

Public randomness: In addition to the t hash pairs $\{(h_1, g_1), \dots, (h_t, g_t)\}$, we assume that the server creates a random partition $\Pi : [n] \rightarrow [\log d] \times [t]$ that assigns to each user $i \in [n]$ a random pair $(\ell_i, j_i) \leftarrow [\log(d)] \times [t]$, and another random function $\mathcal{Q} : [n] \rightarrow [m]$

that assigns³ to each user i a uniformly random index $r_i \leftarrow [m]$. We assume that such random indices ℓ_i, j_i, r_i are shared between the server and each user.

First, we describe the two main modules of our protocol.

3.1.1. A LOCAL RANDOMIZER: LocalRnd

For each $i \in [n]$, user i runs her own independent copy of a local randomizer, denoted as **LocalRnd**, to generate her private report. **LocalRnd** of user i starts by acquiring the index triple $(\ell_i, j_i, r_i) \leftarrow [\log d] \times [t] \times [m]$ from public randomness. For each user, **LocalRnd** is invoked twice in the full protocol: once during the first phase of the protocol (called the pruning phase) where the high-frequency items (*heavy hitters*) are identified, and a second time during the final phase (the estimation phase) to enable the protocol to get better estimates for the frequencies of the heavy hitters. In the first invocation, **LocalRnd** of user i performs its computation on the ℓ_i -th prefix of the item v_i of user i , whereas in the second invocation, it performs the computation on the entire user's string v_i .

Apart from this, in both invocations, **LocalRnd** follows similar steps. It first selects the hash pair (h_{j_i}, g_{j_i}) , computes $c_i = h_{j_i}(v_i[1 : \tilde{\ell}])$ (where $\tilde{\ell} = \ell_i$ in the first invocation and $\tilde{\ell} = \log d$ in the second invocation, and $v_i[1 : \tilde{\ell}]$ is the $\tilde{\ell}$ -th prefix of v_i), then it computes a bit $x_i = g_{j_i}(v_i[1 : \tilde{\ell}]) \cdot W_{r_i, c_i}$ (where $W_{r, c}$ denotes the (r, c) entry of the basis matrix **W**). Finally, to guarantee ϵ -local differential privacy, it generates a randomized response y_i based on x_i (i.e., $y_i = x_i$ with probability $e^{\epsilon/2}/(1 + e^{\epsilon/2})$ and $y_i = -x_i$ with probability $1/(1 + e^{\epsilon/2})$), which is sent to the server.

Our local randomizer can be thought of as a transformed, compressed (via sampling), and randomized version of the count sketch (Charikar et al., 2002). In particular, we can think of **LocalRnd** as follows. It starts off with similar steps to the standard count sketch algorithm, but then deviates from it as it applies Hadamard transform to the user's signal, then samples one bit from the result. By doing so, we can achieve significant savings in space and communication without sacrificing accuracy.

3.1.2. A FREQUENCY ORACLE: FreqOracle

Suppose we want to allow the server to estimate the frequencies of some *given* subset $\widehat{\mathcal{V}} \subseteq \{0, 1\}^\ell$ for some given $\ell \in [\log d]$ based on the noisy users' reports. We give a protocol, denoted as **FreqOracle**, for accomplishing this task.

For each queried item $\hat{v} \in \widehat{\mathcal{V}}$ and for each hash index $j \in [t]$, **FreqOracle** computes $c = h_j(\hat{v})$, then collects the noisy reports of a collection of users $\mathcal{I}_{\ell, j}$ that contains every user i whose pair of prefix and hash indices (ℓ_i, j_i) match (ℓ, j) . Next, it estimates the inverse Hadamard transform of the compressed and noisy signal of each user in $\mathcal{I}_{\ell, j}$. In particular, for each $i \in \mathcal{I}_{\ell, j}$, it computes $y_i W_{r_i, c}$ which can be described as a multiplication between $y_i \mathbf{e}_{r_i}$ (where \mathbf{e}_{r_i} is the indicator vector with 1 at the r_i -th position) and the scaled Hadamard matrix **W**, followed by selecting the c -th entry of the resulting vector. This brings us back to the standard count sketch representation. It then sums all the results and multiplies the outcome by $g_j(\hat{v})$ to obtain an estimate $\hat{f}_j(\hat{v})$ for the frequency of \hat{v} .

3. We could have grouped Π and \mathcal{Q} into one random function mapping $[n]$ to $[\log d] \times [t] \times [m]$, however, we prefer to split them for clarity of exposition as each source of randomness will be used for a different role.

As in the count sketch algorithm, this is done for every $j \in [t]$, then `FreqOracle` obtains a high-confidence estimate by computing the median of all the t frequency estimates.

3.1.3. THE PROTOCOL: `TreeHist`

The protocol is easier to describe via operations over nodes of the prefix tree $\bar{\mathcal{V}}$ of depth $\log d$ (described earlier). The protocol runs through two main phases: the pruning (or, scanning) phase, and the final estimation phase.

In the pruning phase, the protocol scans the levels of the prefix tree starting from the top level (that contains just 0 and 1) to the bottom level (that contains all items of the dictionary). For a given node at level $\ell \in [\log d]$, using `FreqOracle` as a subroutine, the protocol gets an estimate for the frequency of the corresponding ℓ -bit prefix. For any $\ell \in [\log(d) - 1]$, before the protocol moves to level $\ell + 1$ of the tree, it prunes all the nodes in level ℓ that *cannot* be prefixes of actual heavy hitters (high-frequency items in the dictionary). Then, as it moves to level $\ell + 1$, the protocol considers only the children of the surviving nodes in level ℓ . The construction guarantees that, with high probability, the number of surviving nodes in each level cannot exceed $O\left(\sqrt{\frac{n}{\log(d)\log(n)}}\right)$. Hence, the total number of nodes queried by the protocol (i.e., submitted to `FreqOracle`) is at most $O\left(\sqrt{\frac{n \log(d)}{\log(n)}}\right)$.

In the second and final phase, after reaching the final level of the tree, the protocol would have already identified a list of the candidate heavy hitters, however, their estimated frequencies may not be as accurate as we desire due to the large variance caused by the random partitioning of users across all the levels of the tree. Hence, it invokes the frequency oracle once more on those particular items, and this time, the sampling variance is reduced as the set of users is partitioned only across the t hash pairs (rather than across $\log(d) \times t$ bins as in the pruning phase). By doing this, the server obtains more accurate estimates for the frequencies of the identified heavy hitters. The privacy and accuracy guarantees are stated below. The full details are given in Section 5.

3.1.4. PRIVACY AND UTILITY GUARANTEES

Theorem 3.1 *Protocol `TreeHist` is ϵ -local differentially private.*

Theorem 3.2 *There is a number $\eta = O\left(\sqrt{n \log(n/\beta) \log(d)}\right)/\epsilon$ such that with probability at least $1 - \beta$ (over the choice of all public and private randomness), the output list of the `TreeHist` protocol satisfies the following properties:*

1. *it contains all items $v \in \mathcal{V}$ with true frequencies above 3η .*
2. *it does not contain any item $v \in \mathcal{V}$ with true frequency below η .*
3. *Every frequency estimate in the output list is accurate up to an error $\leq O\left(\sqrt{n \log(n/\beta)}\right)/\epsilon$*

3.2. The Bitstogram Protocol

We now present a simplified description of our second protocol, that captures most of the ideas. Any informalities made hereafter are removed in the full description of the protocol (Section 6).

First Step: Frequency Oracle. Recall that a frequency oracle is a protocol that, after communicating with the users, outputs a *data structure* capable of approximating the frequency of every domain element $v \in \mathcal{V}$. So, if we were to allow the server to have linear runtime in the domain size $|\mathcal{V}| = d$, then a frequency oracle would suffice for computing histograms. As we are interested in protocols with a significantly lower runtime, we will only use a frequency oracle as a subroutine, and query it only for (roughly) \sqrt{n} elements.

Let $Z \in \{\pm 1\}^{d \times n}$ be a matrix chosen uniformly at random, and assume that Z is publicly known.⁴ That is, for every domain element $v \in \mathcal{V}$ and every user $j \in [n]$, we have a random bit $Z[v, j] \in \{\pm 1\}$. As Z is publicly known, every user j can identify its corresponding bit $Z[v_j, j]$, where $v_j \in \mathcal{V}$ is the input of user j . Now consider a protocol in which users send randomized responses of their corresponding bits. That is, user j sends $y_j = Z[v_j, j]$ w.p. $\frac{1}{2} + \frac{\epsilon}{2}$ and sends $y_j = -Z[v_j, j]$ w.p. $\frac{1}{2} - \frac{\epsilon}{2}$. We can now estimate the frequency of every domain element $v \in \mathcal{V}$ as

$$a(v) = \frac{1}{\epsilon} \cdot \sum_{j \in [n]} y_j \cdot Z[v, j].$$

To see that $a(v)$ is accurate, observe that $a(v)$ is the sum of n independent random variables (one for every user). For the users j holding the input v (that is, $v_j = v$) we will have that $\frac{1}{\epsilon} \mathbb{E}[y_j \cdot Z[v, j]] = 1$. For the other users we will have that y_j and $Z[v, j]$ are independent, and hence $\mathbb{E}[y_j \cdot Z[v, j]] = \mathbb{E}[y_j] \cdot \mathbb{E}[Z[v, j]] = 0$. That is, $a(v)$ can be expressed as the sum of n independent random variables: $f(v)$ variables with expectation 1, and $(n - f(v))$ variables with expectation 0. The fact that $a(v)$ is an accurate estimation for $f(v)$ now follows from the Hoeffding bound.

Lemma 3.3 (Algorithm Hashtogram) *Let $\epsilon \leq 1$. Algorithm Hashtogram satisfies ϵ -LDP. Furthermore, with probability at least $1 - \beta$, algorithm Hashtogram answers every query $v \in \mathcal{V}$ with $a(v)$ satisfying: $|a(v) - f(v)| \leq O\left(\frac{1}{\epsilon} \cdot \sqrt{n \log\left(\frac{nd}{\beta}\right)}\right)$.*

Second Step: Identifying Heavy-Hitters. Let us assume that we have a frequency oracle protocol with worst-case error τ . We now want to use our frequency oracle in order to construct a protocol that operates on two steps: First, it identifies a small set of potential “heavy-hitters”, i.e., domain elements that appear in the database at least 2τ times. Afterwards, it uses the frequency oracle to estimate the frequencies of those potential heavy elements.⁵

Let $h : \mathcal{V} \rightarrow [T]$ be a (publicly known) random hash function, mapping domain elements into $[T]$, where T will be set later.⁶ We will now use h in order to identify the heavy-hitters.

4. As we later explain, Z has a short description, as it need not be uniform.

5. Event though we describe the protocol as having two steps, the necessary communication for these steps can be done in parallel, and hence, our protocol will have only 1 round of communication.

6. As with the matrix Z , the hash function h can have a short description length.

To that end, let $v^* \in \mathcal{V}$ denote such a heavy-hitter, appearing at least 2τ times in the database S , and denote $t^* = h(v^*)$. Assuming that T is big enough, w.h.p. we will have that v^* is the only input element (from S) that is mapped (by h) into the hash value t^* . Assuming that this is indeed the case, we will now identify v^* bit by bit.

For $\ell \in [\log d]$, denote $S_\ell = (h(v_j), v_{j,\ell})_{j \in [n]}$, where $v_{j,\ell}$ is bit ℓ of v_j . That is, S_ℓ is a database over the domain $([T] \times \{0, 1\})$, where the row corresponding to user j is $(h(v_j), v_{j,\ell})$. Observe that every user can compute her own row locally. As v^* is a heavy-hitter, for every $\ell \in [\log d]$ we have that (t^*, v_ℓ^*) appears in S_ℓ at least 2τ times, where v_ℓ^* is bit ℓ of v^* . On the other hand, as we assumed that v^* is the only input element that is mapped into t^* we get that $(t^*, 1 - v_\ell^*)$ does not appear in S_ℓ at all. Recall that our frequency oracle has error at most τ , and hence, we can use it to accurately determine the bits of v^* .

To make things more concrete, consider the protocol that for every hash value $t \in [T]$, for every coordinate $\ell \in [\log d]$, and for every bit $b \in \{0, 1\}$, obtains an estimation (using the frequency oracle) for the multiplicity of (t, b) in S_ℓ (so there are $\log d$ invocations of the frequency oracle, and a total of $2T \log d$ estimations). Now, for every $t \in [T]$ let us define \hat{v}_t where bit ℓ of \hat{v}_t is the bit b s.t. (t, b) is more frequent than $(t, 1 - b)$ in S_ℓ . By the above discussion, we will have that $\hat{v}_{t^*} = v^*$. That is, the protocol identifies a set of $T \log d$ domain elements, containing all of the heavy-hitters. The frequency of the identified heavy-hitters can then be estimated using the frequency oracle.

Remark 2 *As should be clear from the above discussion, it suffices to take $T \gtrsim n^2$, as this will ensure that there are no collisions among different input elements. As we only care about collisions between “heavy-hitters” (appearing in S at least \sqrt{n} times), it would suffice to take $T \gtrsim n$ to ensure that w.h.p. there are no collisions between heavy-hitters. In fact, we could even take $T \gtrsim \sqrt{n}$, which would ensure that a heavy-hitter v^* has no collisions with constant probability, and then to amplify our confidence using repetitions.*

Lemma 3.4 (Algorithm Bitstogram) *Let $\epsilon \leq 1$. Algorithm Bitstogram satisfies ϵ -LDP. Furthermore, the algorithm returns a list L of length $\tilde{O}(\sqrt{n})$ satisfying:*

1. *With probability $1 - \beta$, for every $(v, a) \in L$ we have that $|a - f(v)| \leq O\left(\frac{1}{\epsilon} \sqrt{n \log(n/\beta)}\right)$.*
2. *W.p. $1 - \beta$, for every $v \in \mathcal{V}$ s.t. $f(v) \geq O\left(\frac{1}{\epsilon} \sqrt{n \log(d/\beta) \log(\frac{1}{\beta})}\right)$, we have that v is in L .*

4. Detailed Experimental Results

In this section we discuss implementation details of our algorithms mentioned in Section 5.⁷ The main objective of this section is to emphasize the empirical efficacy of our algorithms along with the theoretical optimality in terms of error, space, time and communication. Thakurta et al. (2017) recently claimed space optimality for a similar problem, but a formal analysis (or empirical evidence) was not provided. Our experiments corroborate both the analytical bounds in our current work, and in (Thakurta et al., 2017). Our experiments are performed on a macOS-Sierra 10.12 system (in Python 2.7) with 3.3Ghz (Intel Core i5) and 16GB of DDR-3 RAM.

⁷ The experiments are performed without the Hadamard compression during data transmission.

4.1. Private Frequency Oracle

In this experiment, the objective is to test the efficacy of our algorithm in estimating the frequencies of a known set of dictionary of user items, under local differential privacy. We estimate the error in estimation while varying the size of the data set n , changing the privacy parameter ϵ . (See Section 2.1 for a refresher on the notation.)

Figure 1 shows results on a synthetic data set with the domain size of hundred (i.e., $d = 100$) drawn from a power law distribution with power of 15. The default parameters used in Figure 1 are: *number of data samples* (n): 10 million, range of the hash function (m): \sqrt{n} , number of hash functions (t): 285, and the privacy parameter $\epsilon = 2.0$. For the hash functions, we used the prefix bits of SHA-256. The estimated frequency is scaled by the number of samples to normalize the result, and each experiment is averaged over *ten runs*. The bars for **True** refers to the the true frequencies, and the bars for **Priv** corresponds to the differentially private frequencies. The **pctle** corresponds to the frequency of a domain element at the corresponding percentile in the frequency distribution of the data set. *Observations:* i) The plots corroborate the fact that the frequency oracle is indeed *unbiased*. The average frequency estimate (over ten runs) for each percentile is within one standard deviation of the corresponding true estimate. ii) The error in the estimates go down significantly as the number of samples are increased or the privacy parameter ϵ is increased.

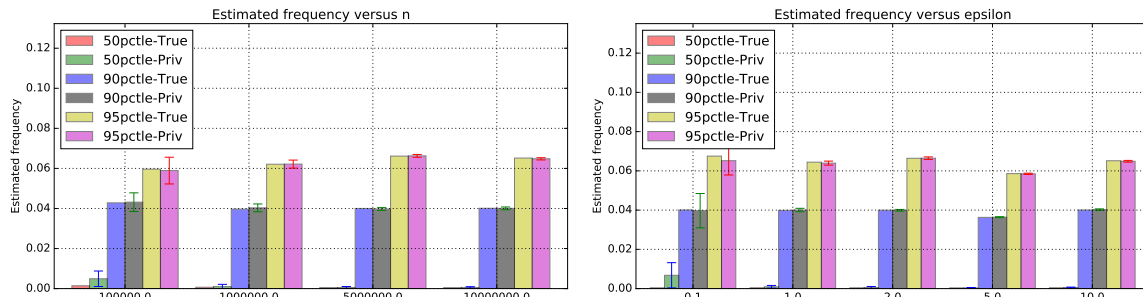


Figure 1: Frequency vs number of samples (n) and privacy parameter (ϵ) on the synthetic data set.

In Figure 2 we show the result of changing the range of the hash function (m). The observation is that the results are seemingly *insensitive* to the range of the hash function.

We also ran the same experiment (Figure 3) on a real data set drawn uniformly at random the NLTK Brown corpus (NLT). The data set we created has $n = 10$ million samples drawn i.i.d. from the corpus with replacement, and the system parameters are the same default parameters described earlier. In this plot, the rank corresponds to the rank of a domain element in the distribution of frequencies in the data set. The observation is here is also consistent with that of Figure 1.

Comparison to RAPPOR (Erlingsson et al., 2014): Here we compare ourselves to the only other system (RAPPOR project from GOOGLE) for the private frequency estimation problem whose code is publicly available. We took the snapshot of their code

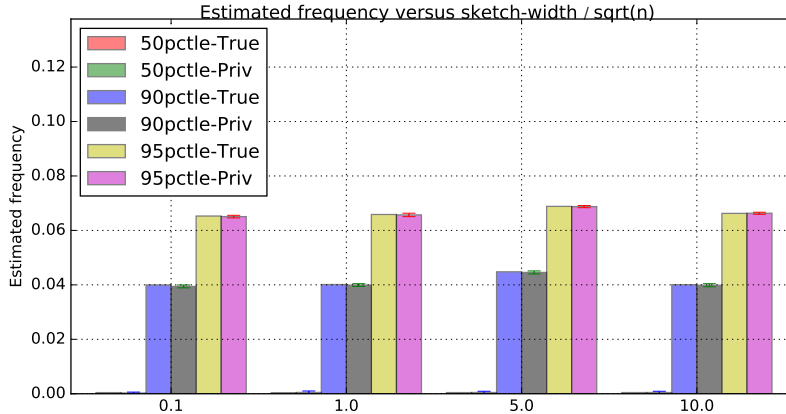


Figure 2: Frequency vs sketch width (m) on the synthetic data set.

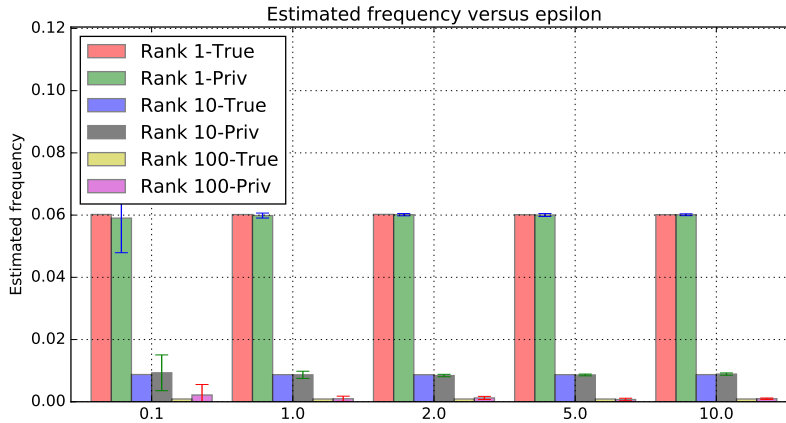


Figure 3: Frequency vs privacy (ϵ) on the NLTK-Brown corpus.

base (<https://github.com/google/rappor>) on May 9th, 2017. In order to perform a fair comparison, we tested our algorithm against one of their demo experiments available (*Demo3* using the `demo.sh` script). We used the privacy parameter $\epsilon = \ln(3)$, the number of data samples $n = 1$ million, and the data set to be the same data set generated by the `demo.sh` script. In Figure 4 we observe that for higher frequencies both RAPPOR and our algorithm perform similarly. However, in lower frequency regimes, the RAPPOR estimates are zero most of the times, while our estimates are closer to the true estimates. *N.B.* We do not claim that our algorithm would outperform the RAPPOR system on all problem instances. However, our current experiment does highlight the need to perform an *at-scale* comparison between the two algorithms.

4.2. Private Heavy-hitters

In this section, we take on the harder task of identifying the heavy hitters, rather than estimating the frequencies of domain elements. We run our experiments on the same two

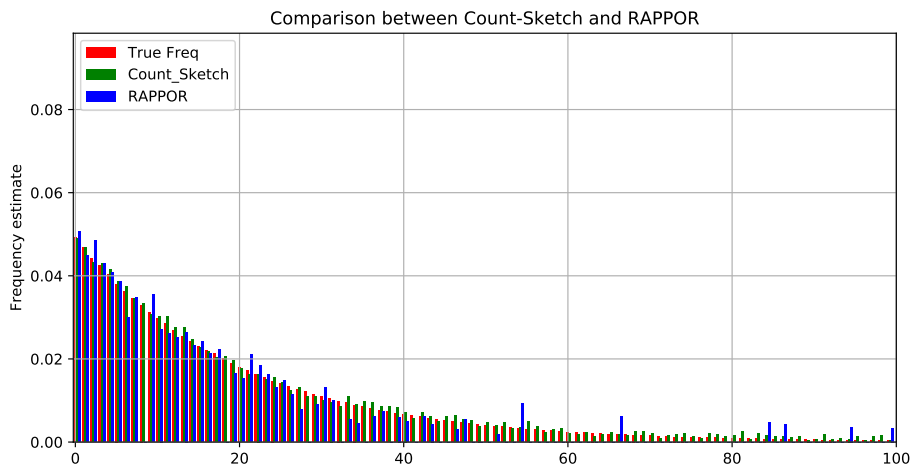


Figure 4: Frequency vs privacy (ϵ) on the Demo 3 experiment from RAPPOR

data sets described earlier with the same setting of parameters, except now we assume that we do not know the domain. As a part of our algorithm design, we assume that every element in the domain is from the English alphabet set [a-z] and are of length exactly equal to six. If they are of larger length, we truncate before entering them in the data set, and if they are of smaller length we tag a \perp at the end. We generate the domain elements for the synthetic data set by first generating the frequency histogram based on the power law distribution described earlier, and then assign random strings of length eight to each bin of the histogram. That becomes our data set. For the NLTK Brown corpus (NLT), we sample $n = 10$ million samples with replacement from the corpus to form our data set. We set a threshold of $15 \cdot \sqrt{n}$ as the threshold for being a heavy hitter. We measure the efficacy of our system by measuring the *precision* and *recall*. Figures 5 and 6 show the true data distribution for the synthetic and the NLTK data set.

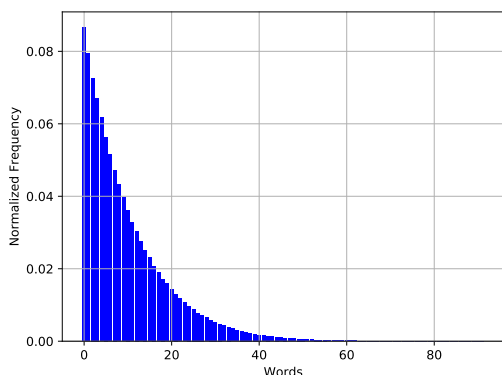


Figure 5: Distr. for Synthetic

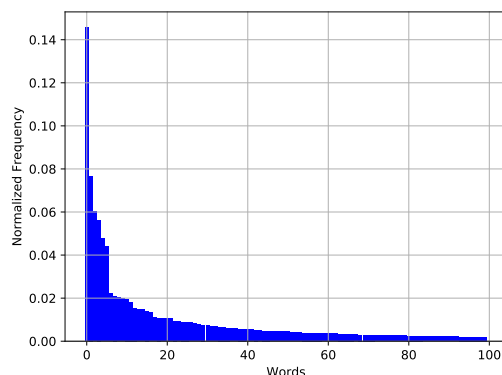


Figure 6: Distr. for Brown (top 100 words)

In Table 4.2 we state our corresponding precision and recall parameters. Our recall numbers are much better than the precision numbers, primarily because of the large number of negative examples (3×10^8 examples). In practice, if there are *false-positives*, they can be easily pruned using domain expertise. For example, if we are trying to identify new words which users are typing in English (app, 2016), then using the domain expertise of English, a set of false positives can be easily ruled out by inspecting the list of heavy hitters output by the algorithm. Further, notice that since we are working with domain elements with size six characters, a brute force algorithm would require 26^6 queries to the frequency oracle, which would be computationally (near) infeasible. While there are other algorithms for finding heavy-hitters (Bassily and Smith, 2015; Hsu et al., 2012), either they do not provide any theoretical guarantee for the utility (Erlingsson et al., 2014; Fanti et al., 2015; Thakurta et al., 2017), or there does not exist a scalable and efficient implementation for them. Our work scores well on both these aspects.

Data set	# of unique words	Precision	Recall
Synthetic	93	0.36 ($\sigma = 0.05$)	0.95 ($\sigma = 0.03$)
NLTK Brown corpus	25991	0.24 ($\sigma = 0.04$)	0.86 ($\sigma = 0.05$)

Table 2: Private Heavy-hitters with threshold= $15\sqrt{n}$. Here σ corresponds to the standard deviation.

5. Locally Private Heavy-hitters via Count-Sketch: The TreeHist Protocol

We start with a detailed description of our construction described at a high level in Section 3.1.

We refer to Section 3.1 for definitions and public parameters that will be used in the construction, namely, *prefixes*, *hashes*, *the basis matrix*, *global parameters*, and *public randomness*. We restate below our public parameters and, when applicable, their specific settings.

Global parameters: We will assume that total number of users n , the size of the Hadamard matrix m , the number of hash pairs t , the privacy parameter ϵ , and the confidence parameter β are public parameters, and hence they will not be explicitly provided as inputs to the algorithms. For integer parameters, we will implicitly assume that results are rounded to the nearest integer. We set $t = 110 \log(n/\beta)$ and $m = 48 \sqrt{\frac{n}{\log(n/\beta)}}$. The hash functions $\{(h_1, g_1), \dots, (h_t, g_t)\}$ will also be assumed to be public information (this is $O(\log(d) \log(n/\beta))$ bits of shared randomness).

Public randomness: In addition to the t hash pairs $\{(h_1, g_1), \dots, (h_t, g_t)\}$, we assume that the server creates a random partition $\Pi : [n] \rightarrow [\log d] \times [t]$ over the set of users, that is, a each user i gets a random pair $(\ell_i, j_i) \leftarrow [\log(d)] \times [t]$ that represents the index of one of $\log(d) \times t$ “buckets.” Moreover, the server uses another random function $\mathcal{Q} : [n] \rightarrow [m]$ that assigns to each user i a uniformly random index $r_i \leftarrow [m]$. We assume that such random indices ℓ_i, j_i, r_i are shared between the server and each user. For each $\ell \in [\log d]$ and each $j \in [t]$, we define $\mathcal{I}_{\ell,j} \triangleq \{i : \ell_i = \ell, j_i = j\}$ and $\mathcal{I}_j \triangleq \{i : j_i = j\} = \cup_{\ell \in [\log d]} \mathcal{I}_{\ell,j}$.

5.1. A Local Randomizer: LocalRnd

For each $i \in [n]$, user i runs her own independent copy of Algorithm 1 below, referred to as **LocalRnd**, to generate her private report. We note that **LocalRnd** takes a flag $\text{Final} \in \{0, 1\}$ as an input. The role of this input will become clear when we discuss the full protocol. In a nutshell, the flag is used to distinguish between two invocations of **LocalRnd**. In particular, the local randomizer of each user is invoked twice in the full protocol: once during the first phase of the protocol (called the pruning phase) where the high-frequency items (*heavy hitters*) are identified, and a second time during the final phase (the final estimation phase) to enable the protocol to get better estimates for the frequencies of the heavy hitters.

Connection to count sketch and Hadamard transform: Our local randomizer can be thought of as a transformed, compressed (via sampling), and randomized version of the count sketch. Up to Step 5 in **LocalRnd** (Algorithm 1), our algorithm follows the standard count sketch algorithm (Charikar et al., 2002). Starting from Step 6, we start to deviate from the standard count sketch as we apply Hadamard transform to the user’s signal, then sample one bit from the result. Indeed Step 6 can be thought of as a composition of two operations: first, we multiply the indicator vector $\mathbf{e}_{c_i} \in \{0, 1\}^m$ by the scaled Hadamard matrix \mathbf{W} (this is equivalent to selecting \mathbf{w}_{c_i} : the c_i -th column of \mathbf{W}), then we randomly sample one entry from \mathbf{w}_{c_i} . By doing so, we can achieve significant savings in space and communication without sacrificing accuracy.

Algorithm 1 LocalRnd

Input: User i input: $v_i \in \mathcal{V}$, Flag: $\text{Final} \in \{0, 1\}$.

- 1: Using shared randomness, get random indices $(\ell_i, j_i) \leftarrow [\log d] \times [t]$ and $r_i \leftarrow [m]$.
- 2: **if** $\text{Final} = 0$ **then**
- 3: Set $s_i := g_{j_i}(v_i[1 : \ell_i])$ and $c_i := h_{j_i}(v_i[1 : \ell_i])$. $\{v[1 : \ell]$ denotes the ℓ -bit prefix of v .}
- 4: **else**
- 5: Set $s_i := g_{j_i}(v_i)$ and $c_i := h_{j_i}(v_i)$
- 6: Compute $x_i := s_i \cdot W_{r_i, c_i}$
 $\{W_{r, c}$ denotes the sign of the (r, c) entry of \mathbf{H}_m (Hadamard matrix of size m).}
- 7: Generate a randomized version y_i of the bit x_i :

$$y_i = \begin{cases} x_i & \text{w.p. } \frac{e^{\epsilon/2}}{e^{\epsilon/2} + 1} \\ -x_i & \text{w.p. } \frac{1}{e^{\epsilon/2} + 1} \end{cases}$$

- 8: **return** y_i .
-

The output of **LocalRnd** is 1 bit per invocation. Hence, during the entire span of the protocol, each user sends 2 bits to the server. Here, we will assume the user’s identity (i.e., the index $i \in [n]$) associated with each report is known to the server (e.g., from a higher layer in the communication protocol stack).

5.2. A Frequency Oracle: FreqOracle

Before describing our protocol for identifying the heavy hitters and estimating their frequencies, we first discuss a protocol for a simpler task. Suppose we want to allow the server to estimate the frequencies of some *given* subset $\widehat{\mathcal{V}} \subseteq \{0, 1\}^\ell$ for some given $\ell \in [\log d]$ based on the users' reports. Algorithm 2 describes a protocol, denoted as **FreqOracle**, for accomplishing this task. Note that in this protocol, we assume that all items whose frequencies are in question are given as inputs.

For each queried item $\hat{v} \in \widehat{\mathcal{V}}$ and for each hash index $j \in [t]$, **FreqOracle** starts by collecting the noisy reports of the collection $\mathcal{I}_{\ell,j}$ of users where each user $i \in \mathcal{I}_{\ell,j}$ is assigned a pair of prefix and hash indices (ℓ_i, j_i) that matches (ℓ, j) . Next, it estimates the inverse Hadamard transform of the compressed and noisy signal of each user in $\mathcal{I}_{\ell,j}$. It then sums all the results and multiplies the outcome by $g_j(\hat{v})$ to obtain an estimate $\hat{f}_j(\hat{v})$ for the frequency of \hat{v} . As in the count sketch algorithm, this is done for every $j \in [t]$, then **FreqOracle** obtains a high-confidence estimate for the frequency of \hat{v} by computing the median of all the t frequency estimates.

Inverse transform and back to count sketch: We note here that Steps 7 in **FreqOracle** (Algorithm 2) can be described as an inverse Hadamard transform of the users' compressed and noisy signals. In particular, each term $y_i W_{r_i,c}$ inside the sum can be described as a multiplication between $y_i \mathbf{e}_{r_i}$ (where \mathbf{e}_{r_i} is the indicator vector with 1 at the r_i -th position) and the scaled Hadamard matrix \mathbf{W} , followed by picking the c -th entry of the resulting vector. This brings us back to the standard count sketch representation (Charikar et al., 2002). The **FreqOracle** protocol then proceeds to process the frequency estimates in the same way a count sketch does. Indeed, Step 8 is a standard step in count sketch algorithm where a high-confidence frequency estimate is obtained via the median technique. Hence, we attain the functionality of the count sketch algorithm with much less space and communication by transforming the users' signals to the Fourier domain, compressing signals via sampling, and then transforming them back at the server.

5.3. Succinct Histogram via a Tree Aggregation Protocol: TreeHist

We now describe our protocol (Algorithm 3), denoted as **TreeHist**, that outputs a succinct histogram, that is, it outputs a list of *heavy hitters* together with estimates for their frequencies. W.l.o.g., we will assume that $n > \log(n/\beta) \log(d)$ since otherwise, we cannot guarantee less than trivial error (i.e., an error of order n). A high-level description of the protocol is given in Section 3.1.3.

The protocol is comprised of two main phases: the pruning phase, and the final estimation phase. We note that the binary flag **Final** is used to distinguish between these two phases. In particular, depending on the value of this flag, the partition of users and the scaling factor passed to **FreqOracle** will differ. More precisely, when **Final** = 0 (the pruning phase), for each prefix length ℓ , the partition of users is given by the collection $\{\mathcal{I}_{\ell,j} : j \in [t]\}$ and the scaling factor is set as $\gamma = t \log d$. When **Final** = 1 (the final phase), the partition of users is $\{\mathcal{I}_j : j \in [t]\}$ and the scaling factor is $\gamma = t$.

Algorithm 2 FreqOracle

Input: Prefix length: $\ell \in [\log d]$, a subset of ℓ -bit prefixes $\widehat{\mathcal{V}} \subseteq \{0, 1\}^\ell$, collection of t disjoint subsets of users: $\{\tilde{\mathcal{I}}_j : j \in [t]\}$, oracle access to items of relevant users: $(v_i : i \in \cup_{j \in [t]} \tilde{\mathcal{I}}_j)$, scaling factor: γ , Flag: Final $\in \{0, 1\}$. $\{a_\epsilon \triangleq \frac{e^{\epsilon/2} + 1}{e^{\epsilon/2} - 1}$ is assumed to be a global constant seen by FreqOracle.}

- 1: **for** $\hat{v} \in \widehat{\mathcal{V}}$ **do**
- 2: **for** Hash index $j = 1$ to t **do**
- 3: Set $s := g_j(\hat{v})$ and $c := h_j(\hat{v})$.
- 4: **for** Users $i \in \tilde{\mathcal{I}}_j$ **do**
- 5: Get user- i 's 1-bit report: $y_i = \text{LocalRnd}(v_i, \text{Final})$
- 6: Get user- i 's random index $r_i = \mathcal{Q}(i)$ using public randomness.
- 7: Compute the j -th estimate of the frequency of \hat{v} : $\hat{f}_j(\hat{v}) := \gamma \cdot a_\epsilon \sum_{i \in \tilde{\mathcal{I}}_j} y_i \cdot s \cdot W_{r_i, c}$.
- 8: Compute final estimate for the frequency of \hat{v} : $\hat{f}(\hat{v}) := \text{Median} \left(\{ \hat{f}_j(\hat{v}) : j \in [t] \} \right)$.
- 9: FreqList := $\left\{ \left(\hat{v}, \hat{f}(\hat{v}) \right) : \hat{v} \in \widehat{\mathcal{V}} \right\}$.
- 10: **return** FreqList.

Algorithm 3 TreeHist The Full Protocol

Input: Oracle access to users' items $(v_i \in \mathcal{V} : i \in [n])$.

- 1: Set $\eta := 147 \sqrt{n \log(n/\beta) \log(d)}/\epsilon$ and $a_\epsilon := \frac{e^{\epsilon/2} + 1}{e^{\epsilon/2} - 1}$.
- 2: Set $\gamma := t \log d$ {This setting will change at the final stage of the protocol.}
- 3: Initialize Prefixes = $\{\perp\}$.
- 4: **for** Prefix length $\ell = 1$ to $\log d$ **do**
- 5: Get the collection $\{\mathcal{I}_{\ell, j} : j \in [t]\}$ using the random partition Π . {See ‘‘Public randomness’’ above.}
- 6: $\left\{ \left(\hat{v}, \hat{f}(\hat{v}) \right) : \hat{v} \in \text{ChildSet}(\text{Prefixes}) \right\} =$
 FreqOracle $\left(\ell, \text{ChildSet}(\text{Prefixes}), \{\mathcal{I}_{\ell, j} : j \in [t]\}, (v_i : i \in \cup_{j \in [t]} \mathcal{I}_{\ell, j}), \gamma, \text{Final} = 0 \right)$.
- 7: Initialize NewPrefixes = \emptyset .
- 8: **for** $v \in \text{ChildSet}(\text{Prefixes})$ **do**
- 9: **if** $\hat{f}(\hat{v}) \geq 2\eta$ **then**
- 10: Add \hat{v} to NewPrefixes.
- 11: Update Prefixes \leftarrow NewPrefixes.
- 12: Set $\gamma := t$
- 13: SuccHist := FreqOracle $\left(\log d, \text{Prefixes}, \{\mathcal{I}_j : j \in [t]\}, (v_i : i \in [n]), \gamma, \text{Final} = 1 \right)$.
 {Here $\mathcal{I}_j = \cup_{\ell \in [\log d]} \mathcal{I}_{\ell, j}$ as defined earlier in the ‘‘Public randomness’’ paragraph.}
- 14: **return** SuccHist.

5.3.1. RUNNING TIME AND PROCESSING MEMORY OF TreeHist

Running Time: We note that Algorithm FreqOracle (Algorithm 2) is invoked $\log(d) + 1 = O(\log d)$ times by TreeHist (Algorithm 3): once per each level of the tree and another at the final estimation phase at the bottom level of the tree.

The main time-consuming step in FreqOracle is the summation in Step 7. This step is executed $O\left(\sqrt{\frac{n \log n}{\log d}}\right)$ times in every invocation of FreqOracle since there are $O\left(\sqrt{\frac{n}{\log n \log d}}\right)$ nodes queried in every level of the tree and for each node FreqOracle computes $t \approx \log n$ frequency estimates. That is, in total, this step is executed $O(\sqrt{n \log n \log d})$ times in the TreeHist protocol. A direct implementation would involve summing $\approx n$ bits each time this step is executed, and hence would amount for running time of $\approx n^{1.5}$. However, we now show that this can be reduced to $\approx n$.

Consider Step 7 of Algorithm FreqOracle when FreqOracle is invoked by TreeHist at any given prefix level $\ell \in [\log d]$ during the pruning phase. We will not consider the final estimation phase here since its running time is dominated by that of the pruning phase. Note that since the size of the basis matrix \mathbf{W} is $m = O\left(\sqrt{\frac{n}{\log n}}\right)$, there are only $O\left(\sqrt{\frac{n}{\log n}}\right)$ values that the row index r_i can take. Hence, $\hat{f}_j(\hat{v})$ (computed in Step 7) can be expressed as follows:

$$\hat{f}_j(\hat{v}) = \gamma \cdot a_\epsilon \sum_{\kappa \in [m]} \left(\sum_{i \in \mathcal{I}_{\ell, j}: r_i = \kappa} y_i \right) \cdot s \cdot W_{\kappa, c}.$$

Thus, to implement Step 7 of FreqOracle for all items (ℓ, j, \hat{v}) , we first compute $\left(\sum_{i \in \mathcal{I}_{\ell, j}: r_i = \kappa} y_i \right)$ for every $\kappa \in [m]$, $\ell \in [\log d]$, and $j \in [t]$ (this amounts to a running time of $O(n)$ in total). Then, for every value of (ℓ, j, \hat{v}) , computing $\hat{f}_j(\hat{v})$ would require summing $m = O\left(\sqrt{\frac{n}{\log n}}\right)$ numbers. Hence, in total, the running time of TreeHist is $O\left(\sqrt{\frac{n}{\log n \log d}} \cdot m \cdot t \log d\right) = O(n\sqrt{\log d})$.

Processing memory: For the implementation described above, Algorithm FreqOracle maintains $m \cdot t \cdot \log d$ sums of at most n bits each. This would require a processing memory of $O(\sqrt{n} \log^{1.5}(n) \log(d))$ bits. The memory required for all the remaining steps of the TreeHist protocol does not exceed this amount, and hence, the total processing memory required by TreeHist is $O(\sqrt{n} \log^{1.5}(n) \log(d))$ bits.

5.3.2. PRIVACY AND UTILITY GUARANTEES

In this section we provide the privacy and utility guarantees for the TreeHist protocol.

Theorem 5.1 [*ϵ -Local Differential Privacy of the TreeHist Protocol*] *The TreeHist protocol (Algorithm 3) is ϵ -local differentially private.*

Theorem 5.2 [Utility of the TreeHist Protocol] *Let $\eta = 147\sqrt{n \log(n/\beta) \log(d)}/\epsilon$. With probability at least $1 - \beta$ (over the choice of all public and private randomness), the output SuccHist of the TreeHist protocol (Algorithm 3) satisfies the following properties:*

1. SuccHist contains all items $v \in \mathcal{V}$ with true frequencies above 3η .
2. SuccHist does not contain any item $v \in \mathcal{V}$ with true frequency below η .
3. Every frequency estimate in SuccHist is accurate up to an error $\leq 147\frac{\sqrt{n \log(n/\beta)}}{\epsilon}$

We defer the proofs of Theorems 5.1 and 5.2 to Appendix A.

The following lemma states the error guarantees of the Algorithm FreqOracle used by the TreeHist protocol. Note that FreqOracle is invoked by the TreeHist protocol during both pruning and final estimation phases of the protocol. This lemma is central to our proof of Theorem 5.2.

In the following lemma, we will use $\{\tilde{\mathcal{I}}_j\}$ to generically denote the collection of users subsets passed to FreqOracle. As described early in Section 5.3, the form of this collection depends on whether the protocol is in the pruning phase (the flag Final = 0) or the final estimation phase (Final = 1). In particular, for any prefix length $\ell \in [\log d]$, note that when FreqOracle is invoked by TreeHist in the pruning phase (i.e., Final = 0), the collection $\{\tilde{\mathcal{I}}_j : j \in [t]\} = \{\mathcal{I}_{\ell,j} : j \in [t]\}$ and the scaling factor $\gamma = t \log d$, whereas when FreqOracle is invoked by TreeHist in the final phase (i.e., Final = 1), the collection $\{\tilde{\mathcal{I}}_j : j \in [t]\} = \{\mathcal{I}_j : j \in [t]\}$ and $\gamma = t$. We will use the generic notation $\{\tilde{\mathcal{I}}_j : j \in [t]\}$ and γ since the same proof works for both cases.

Lemma 5.3 *Let $\beta \in (0, 1)$. Let the number of hash pairs $t \geq 110 \log(n/\beta)$, and the size of Hadamard basis matrix $m \geq 48\sqrt{\frac{n}{\log(n/\beta)}}$. Consider algorithm FreqOracle (Algorithm 2) as invoked by the TreeHist protocol (Algorithm 3). For any $\ell \in [\log d]$, suppose that FreqOracle is invoked on inputs: ℓ , subset $\hat{\mathcal{V}} \subseteq \{0, 1\}^\ell$ of size $|\hat{\mathcal{V}}| \leq \sqrt{n}$, collection of users' subsets $\{\tilde{\mathcal{I}}_j : j \in [t]\}$, and scaling factor γ . Assuming that $n \geq 48t \log(d)$, then, with probability $1 - \frac{\beta}{\log d}$, we have:*

$$\forall v \in \hat{\mathcal{V}} : |f(v) - \hat{f}(v)| \leq 14\sqrt{n\gamma}/\epsilon.$$

where $\hat{f}(v)$ is the estimate of FreqOracle for the true frequency $f(v)$ of the item $v \in \hat{\mathcal{V}}$.

Proof Fix $\beta \in (0, 1)$, $\ell \in [\log d]$, $t \geq 110 \log(n/\beta)$, $m \geq 48\sqrt{\frac{n}{\log(n/\beta)}}$, and a subset $\hat{\mathcal{V}} \subseteq \{0, 1\}^\ell$ of size $|\hat{\mathcal{V}}| \leq \sqrt{n}$.

There are three independent sources of randomness. The first is due to randomness in the collection $\{\tilde{\mathcal{I}}_j : j \in [t]\}$ induced by the random partitioning of users via Π . The second source of randomness is due to the randomness in the choice of the t hash pairs $(h_1, g_1), \dots, (h_t, g_t)$. The third source of randomness is due to the random row indices $r_i, i \in [n]$, generated by \mathcal{Q} , and the randomization for privacy (step 7 in Algorithm 1).

Before we discuss the guarantees we can attain under these sources of randomness, we first introduce some notation. For $j \in [t]$, $v \in \hat{\mathcal{V}}$, we define $f_{\tilde{\mathcal{I}}_j}(v) \triangleq \sum_{i \in \tilde{\mathcal{I}}_j} \mathbf{1}(v_i[1 : \ell] = v)$;

that is, $f_{\tilde{\mathcal{I}}_j}(v)$ is the number of users in $\tilde{\mathcal{I}}_j$ whose the ℓ -prefix of their items is v , and define $n_{\tilde{\mathcal{I}}_j} \triangleq |\tilde{\mathcal{I}}_j|$, i.e., $n_{\tilde{\mathcal{I}}_j}$ is the number of users in $\tilde{\mathcal{I}}_j$. Let $\hat{f}_{\tilde{\mathcal{I}}_1}(v), \dots, \hat{f}_{\tilde{\mathcal{I}}_t}(v)$ be independent Poisson random variables with mean $\frac{f(v)}{\gamma}$, and let $\hat{n}_{\tilde{\mathcal{I}}_1}, \dots, \hat{n}_{\tilde{\mathcal{I}}_t}$ be independent Poisson random variables with mean $\frac{n}{\gamma}$.

For each $v \in \hat{\mathcal{V}}$, let

$$\mathcal{G}_1(v) = \{j \in [t] : |\gamma f_{\tilde{\mathcal{I}}_j}(v) - f(v)| \leq 4\sqrt{n\gamma}, \frac{n}{2} \leq \gamma n_{\tilde{\mathcal{I}}_j} \leq 2n\}.$$

Claim 5.4 *With probability $1 - \frac{\beta}{3 \log d}$ over the randomness in $\{\tilde{\mathcal{I}}_j : j \in [t]\}$, for every $v \in \hat{\mathcal{V}}$, we have $|\mathcal{G}_1(v)| \geq \frac{9}{10}t$.*

Fix $v \in \hat{\mathcal{V}}$ and $j \in [t]$. Consider the Poisson random variables $hf_{\tilde{\mathcal{I}}_j}(v)$ and $\hat{n}_{\tilde{\mathcal{I}}_j}$ with mean $\frac{f(v)}{\gamma}$ and $\frac{n}{\gamma}$, respectively. By Theorem 2.5 (a tail bound for the Poisson distribution), the union bound, and assuming that $n \geq 48\gamma$ (the assumption stated in the lemma), then with probability at least 0.996, we have

$$|\gamma \hat{f}_{\tilde{\mathcal{I}}_j}(v) - f(v)| \leq 4\sqrt{n\gamma}, \quad \frac{n}{2} \leq \gamma \hat{n}_{\tilde{\mathcal{I}}_j} \leq 2n.$$

Now, consider the sequences of independent Poisson random variables $\hat{f}_{\tilde{\mathcal{I}}_1}(v), \dots, \hat{f}_{\tilde{\mathcal{I}}_t}(v)$ and $\hat{n}_{\tilde{\mathcal{I}}_1}, \dots, \hat{n}_{\tilde{\mathcal{I}}_t}$ as defined above. Using the fact that $t \geq 110 \log(n/\beta) \geq 55 \log\left(\frac{3en \log d}{\beta}\right)$, then by Chernoff's bound, with probability at least $1 - \frac{\beta}{3en \log d}$, we have

$$\left| \{j \in [t] : |\gamma \hat{f}_{\tilde{\mathcal{I}}_j}(v) - f(v)| \leq 4\sqrt{n\gamma}, \quad \frac{n}{2} \leq \gamma \hat{n}_{\tilde{\mathcal{I}}_j} \leq 2n \} \right| \geq \frac{9}{10}t.$$

Using Theorem 2.4 (the Poisson approximation), with probability at least $1 - \frac{\beta}{3\sqrt{n} \log(d)}$, we have $|\mathcal{G}_1(v)| \geq \frac{9}{10}t$. Hence, by the fact that $|\hat{\mathcal{V}}| \leq \sqrt{n}$ and the union bound, with probability at least $1 - \frac{\beta}{3 \log d}$, for all $v \in \hat{\mathcal{V}}$, we have $|\mathcal{G}_1(v)| \geq \frac{9}{10}t$.

For the remainder of the proof, we will condition on the event in Claim 5.4.

Let \mathbf{w}_c denote the c -th column of the basis matrix \mathbf{W} . Since the prefix length ℓ is fixed, we will denote $v_i[1 : \ell]$ (the ℓ -prefix of the item of user i) as v_i for brevity.

For each $v \in \hat{\mathcal{V}}$, let

$$\mathcal{G}_2(v) = \left\{ j \in [t] : \left| \gamma \sum_{i \in \tilde{\mathcal{I}}_j} \frac{g_j(v_i) \cdot g_j(v)}{m} \cdot \langle \mathbf{w}_{h_j(v_i)}, \mathbf{w}_{h_j(v)} \rangle - f(v) \right| \leq 8\sqrt{\gamma n} \right\}.$$

Claim 5.5 *Conditioned on the event in Claim 5.4, with probability at least $1 - \frac{\beta}{3 \log d}$ over the choice of the t hash pairs $\{(h_j, g_j) : j \in [t]\}$, for all $v \in \hat{\mathcal{V}}$, we have $|\mathcal{G}_2(v)| \geq \frac{4}{5}t$.*

Fix $v \in \hat{\mathcal{V}}$ and $j \in \mathcal{G}_1(v)$. First consider $\left| \sum_{i \in \tilde{\mathcal{I}}_j} \frac{g_j(v_i) \cdot g_j(v)}{m} \cdot \langle \mathbf{w}_{h_j(v_i)}, \mathbf{w}_{h_j(v)} \rangle - f_{\tilde{\mathcal{I}}_j}(v) \right|$. Note that the columns of \mathbf{W} are orthogonal, and each of them has norm \sqrt{m} . Hence, the

error quantity above comes from those $i \in \tilde{\mathcal{I}}_j$ with $v_i \neq v$ and yet $h_j(v_i) = h_j(v)$. In particular, we can write this error as

$$\left| \sum_{i \in \tilde{\mathcal{I}}_j} z_i g_j(v_i) g_j(v) \right|$$

where $z_i = \mathbf{1}(v_i \neq v, h_j(v_i) = h_j(v))$. Define $\mathcal{V}_{\text{Bad}} = \{\hat{v} \in \hat{\mathcal{V}} : f_{\tilde{\mathcal{I}}_j}(\hat{v}) \geq 2\sqrt{n_{\tilde{\mathcal{I}}_j}}\}$. Note that $|\mathcal{V}_{\text{Bad}}| \leq \sqrt{n_{\tilde{\mathcal{I}}_j}}/2$. Now, observe that

$$\begin{aligned} \mathbb{P}_{h_j, g_j} \left[\left| \sum_{i \in \tilde{\mathcal{I}}_j} z_i g_j(v_i) g_j(v) \right| > 2\sqrt{n_{\tilde{\mathcal{I}}_j}} \right] &\leq \mathbb{P}_{h_j, g_j} [\exists \hat{v} \in \mathcal{V}_{\text{Bad}} : h_j(\hat{v}) = h_j(v)] \\ &+ \mathbb{P}_{h_j, g_j} \left[\left| \sum_{i \in \tilde{\mathcal{I}}_j} z_i g_j(v_i) g_j(v) \right| > 2\sqrt{n_{\tilde{\mathcal{I}}_j}} \mid \forall \hat{v} \in \mathcal{V}_{\text{Bad}}, h_j(\hat{v}) \neq h_j(v) \right] \end{aligned}$$

By the pairwise independence property of h_j and the union bound, the first probability term on the right hand side is bounded from above by $\frac{|\mathcal{V}_{\text{Bad}}|}{m} = \frac{\sqrt{n_{\tilde{\mathcal{I}}_j}}}{2m}$. We now consider the second probability term. Note that the event we conditioned on in the second probability term implies that for every $i \in \tilde{\mathcal{I}}_j$ where $v_i \neq v$, we must have $f_{\tilde{\mathcal{I}}_j}(v_i) < 2\sqrt{n_{\tilde{\mathcal{I}}_j}}$. Hence, conditioned on this event, by the pairwise independence of each of h_j and g_j , we have

$$\begin{aligned} \mathbf{Var} \left[\sum_{i \in \tilde{\mathcal{I}}_j} z_i g_j(v_i) g_j(v) \right] &= \sum_{i \in \tilde{\mathcal{I}}_j} \mathbf{Var} [z_i g_j(v_i)] + \sum_{\substack{i, k \in \tilde{\mathcal{I}}_j: \\ v_i = v_k \neq v}} \mathbb{E} [z_i z_k] \mathbb{E} [g_j(v_i) g_j(v_k)] \\ &\leq n_{\tilde{\mathcal{I}}_j}/m + 2n_{\tilde{\mathcal{I}}_j}^{3/2}/m \leq 3n_{\tilde{\mathcal{I}}_j}^{3/2}/m. \end{aligned}$$

Hence, by using Chebyshev's inequality, the second probability term is bounded by $\frac{3\sqrt{n_{\tilde{\mathcal{I}}_j}}}{4m}$. Hence, we have

$$\mathbb{P}_{h_j, g_j} \left[\left| \sum_{i \in \tilde{\mathcal{I}}_j} z_i g_j(v_i) g_j(v) \right| > 2\sqrt{n_{\tilde{\mathcal{I}}_j}} \right] \leq \frac{5\sqrt{n_{\tilde{\mathcal{I}}_j}}}{4m} \leq \frac{\sqrt{\log(n/\beta)}}{25\sqrt{\gamma}} \leq \frac{1}{250},$$

where the last inequality follows from the fact that $j \in \mathcal{G}_1(v)$ and the fact that $m \geq 48\sqrt{\frac{n}{\log(n/\beta)}}$. Thus, with probability at least 0.996, we have

$$\left| \sum_{i \in \tilde{\mathcal{I}}_j} \frac{g_j(v_i) \cdot g_j(v)}{m} \cdot \langle \mathbf{w}_{h_j(v_i)}, \mathbf{w}_{h_j(v)} \rangle - f_{\tilde{\mathcal{I}}_j}(v) \right| \leq 2\sqrt{n_{\tilde{\mathcal{I}}_j}}.$$

Now, as we conditioned on the event in Claim 5.4, this implies that with probability at least 0.996, we have

$$\begin{aligned} & \left| \gamma \sum_{i \in \tilde{\mathcal{I}}_j} \frac{g_j(v_i) \cdot g_j(v)}{m} \cdot \langle \mathbf{w}_{h_j(v_i)}, \mathbf{w}_{h_j(v)} \rangle - f(v) \right| \\ & \leq \left| \gamma \sum_{i \in \tilde{\mathcal{I}}_j} \frac{g_j(v_i) \cdot g_j(v)}{m} \cdot \langle \mathbf{w}_{h_j(v_i)}, \mathbf{w}_{h_j(v)} \rangle - \gamma f_{\tilde{\mathcal{I}}_j}(v) \right| + |\gamma f_{\tilde{\mathcal{I}}_j}(v) - f(v)| \\ & \leq 8\sqrt{\gamma n}. \end{aligned}$$

Conditioned on the event of Claim 5.4, $|\mathcal{G}_1(v)| \geq \frac{9}{10}t \geq 99 \log(n/\beta) \geq 49 \log\left(\frac{3\sqrt{n} \log d}{\beta}\right)$. We note also that the sums $\left\{ \sum_{i \in \tilde{\mathcal{I}}_j} \frac{g_j(v_i) \cdot g_j(v)}{m} \cdot \langle \mathbf{w}_{h_j(v_i)}, \mathbf{w}_{h_j(v)} \rangle : j \in [t] \right\}$ are independent. Thus, conditioned on the event of Claim 5.4, by Chernoff's bound, with probability at least $1 - \frac{\beta}{3\sqrt{n} \log d}$, we have $|\mathcal{G}_2(v)| \geq \frac{22}{25}|\mathcal{G}_1(v)| \geq \frac{4}{5}t$. Hence, by the fact that $|\hat{\mathcal{V}}| \leq \sqrt{n}$ and the union bound, with probability at least $1 - \frac{\beta}{3 \log d}$, for all $v \in \hat{\mathcal{V}}$, we have $|\mathcal{G}_2(v)| \geq \frac{4}{5}t$. We continue the proof while conditioning on this event as well.

For every $i \in [n]$ let $r_i \leftarrow [m]$ be the row index chosen uniformly at random for user i using the random function \mathcal{Q} , and let y_i denote the randomized bit generated by user i in step 7 of Algorithm 1. As was denoted in algorithm `FreqOracle`, for each $v \in \hat{\mathcal{V}}$, let $\hat{f}_j(v) = \gamma a_\epsilon \sum_{i \in \tilde{\mathcal{I}}_j} y_i g_j(v) W_{r_i, h_j(v)}$ denote the j -th frequency estimate of `FreqOracle` for $v \in \hat{\mathcal{V}}$.

For each $v \in \hat{\mathcal{V}}$, let

$$\mathcal{G}_3(v) = \left\{ j \in [t] : |\hat{f}_j(v) - f(v)| \leq 14\sqrt{n\gamma}/\epsilon \right\}.$$

Claim 5.6 *Conditioned on the events in Claims 5.4 and 5.5, with probability at least $1 - \frac{\beta}{3 \log d}$ over the randomness in $\{(r_i, y_i) : i \in \tilde{\mathcal{I}}_j, j \in [t]\}$, for all $v \in \hat{\mathcal{V}}$, we have $|\mathcal{G}_3(v)| \geq \frac{7}{10}t$*

Fix $v \in \hat{\mathcal{V}}$ and $j \in \mathcal{G}_2(v)$. Note that, conditioned on any realization for $\tilde{\mathcal{I}}_j$, each term in the sum

$$\sum_{i \in \tilde{\mathcal{I}}_j} (a_\epsilon y_i g_j(v) W_{r_i, h_j(v)} - \frac{g_j(v_i) g_j(v)}{m} \langle \mathbf{w}_{h_j(v_i)}, \mathbf{w}_{h_j(v)} \rangle)$$

is independent, zero mean random variable whose support length is bounded by $a_\epsilon + 1 = O(\frac{1}{\epsilon})$. Hence, by Chernoff's bound, with probability at least 0.99, we have

$$\left| \sum_{i \in \tilde{\mathcal{I}}_j} (a_\epsilon y_i g_j(v) W_{r_i, h_j(v)} - \frac{g_j(v_i) g_j(v)}{m} \langle \mathbf{w}_{h_j(v_i)}, \mathbf{w}_{h_j(v)} \rangle) \right| \leq 4\sqrt{n_{\tilde{\mathcal{I}}_j}}/\epsilon.$$

Thus, conditioned on the events in Claims 5.4 and 5.5, with probability at least 0.99, we have

$$\left| \gamma \sum_{i \in \tilde{\mathcal{I}}_j} a_\epsilon y_i g_j(v) W_{r_i, h_j(v)} - f(v) \right| \leq 14\sqrt{\gamma n}/\epsilon,$$

i.e., $|\hat{f}_j(v) - f(v)| \leq 14\sqrt{n\gamma}/\epsilon$.

Conditioned on the event of Claim 5.5, $|\mathcal{G}_2(v)| \geq \frac{4}{5}t \geq 88 \log(n/\beta) \geq 44 \log\left(\frac{3\sqrt{n} \log d}{\beta}\right)$. We note also that the above sums for $j = 1, \dots, t$ are independent. Thus, conditioned on the event of Claim 5.5, by Chernoff's bound, with probability at least $1 - \frac{\beta}{3\sqrt{n} \log d}$, we have $|\mathcal{G}_3(v)| \geq \frac{22}{25}|\mathcal{G}_2(v)| \geq \frac{88}{125}t \geq \frac{7}{10}t$. Hence, by the fact that $|\hat{\mathcal{V}}| \leq \sqrt{n}$ and the union bound, with probability at least $1 - \frac{\beta}{3 \log d}$, for all $v \in \hat{\mathcal{V}}$, we have $|\mathcal{G}_3(v)| \geq \frac{7}{10}t$.

By combining Claims 5.4, 5.5, and 5.6, we conclude that with probability at least $1 - \frac{\beta}{\log d}$, for all $v \in \hat{\mathcal{V}}$, we have

$$\left| \{j \in [t] : |\hat{f}_j(v) - f(v)| \leq 14\sqrt{n\gamma}/\epsilon\} \right| > \frac{1}{2}t.$$

Since for any item v , the final frequency estimate $\hat{f}(v)$ generated by `FreqOracle` is the median of $\hat{f}_1(v), \dots, \hat{f}_t(v)$, then the above implies that with probability at least $1 - \frac{\beta}{\log d}$, for all $v \in \hat{\mathcal{V}}$, we have $|\hat{f}(v) - f(v)| \leq 14\sqrt{n\gamma}/\epsilon$. This completes the proof of the lemma. ■

6. Locally Private Heavy-hitters bit-by-bit: The Bitstogram Protocol

In this section we provide the full details for our `Bitstogram` protocol. We will use the following notation. Let $S \in \mathcal{V}^n$ be a database, which may be distributed across n users (each holding one row). For $v \in \mathcal{V}$, we will be interested in estimating the the multiplicity of v in S , i.e., $f_S(v) = |\{v_i \in S : v_i = v\}|$. As we explained in Section 3.2, we provide both an alternative frequency oracle protocol, as well as an alternative protocol for identifying the heavy-hitters.

6.1. Warmup: A Simple Protocol for Heavy-Hitters

For readability, we first present a simplification of our protocol that captures most of the ideas presented in Section 3.2. This simplification does not achieve the optimal error rate, space complexity, and time complexity that we aim for. We will later modify this construction in order to reduce these complexities.

6.1.1. FREQUENCY ORACLE

Our protocols use the simple local randomizer \mathcal{R} (Algorithm 4), where every user holds one bit, and flips it with probability $1/(e^\epsilon + 1)$.

Algorithm 4 \mathcal{R} : Basic Randomizer

Inputs: $x \in \{\pm 1\}$, and privacy parameter ϵ .

1. Generate and return a random bit $z = \begin{cases} x & \text{w.p. } e^\epsilon/(e^\epsilon + 1) \\ -x & \text{w.p. } 1/(e^\epsilon + 1) \end{cases}$
-

Let $Z \in \{\pm 1\}^{d \times n}$ be a matrix chosen uniformly at random, and assume that Z is publicly known. That is, for every domain element $v \in \mathcal{V}$ and every user $j \in [n]$, we have a random bit $Z[v, j] \in \{\pm 1\}$. As Z is publicly known, every user j can identify its corresponding bit $Z[v_j, j]$, where $v_j \in \mathcal{V}$ is the input of user j . We now present a simple protocol, **ExplicitHist**, in which users send randomized responses of their corresponding bits in this matrix. As we show, this allows the server to estimate the frequency of every domain element.

Algorithm 5 ExplicitHist

Public randomness: Uniformly random matrix $Z \in \{\pm 1\}^{d \times n}$.

Setting: Each player $j \in [n]$ holds a value $v_j \in \mathcal{V}$. Define $S = (v_1, \dots, v_n)$.

Define $\tilde{S} = (\tilde{v}_1, \dots, \tilde{v}_n)$ where $\tilde{v}_j = Z[v_j, j]$.

Input: Oracle access to \tilde{S} .

1. For $j \in [n]$, get user- j 's 1-bit report: $y_j \leftarrow \mathcal{R}(\tilde{v}_j)$.
 2. On input $v \in \mathcal{V}$, return $a(v) = \frac{e^\epsilon + 1}{e^\epsilon - 1} \cdot \sum_{j \in [n]} y_j \cdot Z[v, j]$, and wait for the next input.
-

Lemma 6.1 *Let $\epsilon \leq 1$, and fix a subset $V \subseteq \mathcal{V}$ of size $d' \leq d$. With probability at least $1 - \beta$, algorithm **ExplicitHist** answers every $v \in V$ with $a(v)$ satisfying:*

$$|a(v) - f_S(v)| \leq \frac{3}{\epsilon} \sqrt{n \cdot \ln(4d'/\beta)}.$$

Proof Fix $v \in V$, and denote $c(v) = \sum_{j \in [n]} y_j \cdot Z[v, j]$, and recall that algorithm **ExplicitHist** answers the query v with $a(v) = \frac{e^\epsilon + 1}{e^\epsilon - 1} \cdot c(v)$. We start by analyzing the expectation of $c(v)$:

$$\begin{aligned} \mathbb{E}[c(v)] &= \sum_{j \in [n]} \mathbb{E}[y_j \cdot Z[v, j]] = \sum_{j \in [n]: v_j = v} \mathbb{E}[y_j \cdot Z[v, j]] + \sum_{j \in [n]: v_j \neq v} \mathbb{E}[y_j \cdot Z[v, j]] \\ &= \sum_{j \in [n]: v_j = v} \mathbb{E}[y_j \cdot Z[v, j]] + \sum_{j \in [n]: v_j \neq v} \mathbb{E}[y_j] \cdot \mathbb{E}[Z[v, j]] = f_S(v) \cdot \frac{e^\epsilon - 1}{e^\epsilon + 1}. \end{aligned}$$

That is, $c(v)$ can be expressed as two sums of ± 1 independent random variables: $f_S(v)$ variables with expectation $\frac{e^\epsilon - 1}{e^\epsilon + 1}$, and $(n - f_S(v))$ variables with expectation 0. Using the Hoeffding bound, with probability at least $1 - \frac{\beta}{d'}$ we have that $|c(v) - \frac{e^\epsilon - 1}{e^\epsilon + 1} \cdot f_S(v)| \leq \sqrt{n \cdot \ln(4d'/\beta)}$. That is, $|a(v) - f_S(v)| \leq \frac{e^\epsilon + 1}{e^\epsilon - 1} \cdot \sqrt{n \cdot \ln(4d'/\beta)}$. Using the union bound, this holds simultaneously for every $v \in V$ with probability at least $1 - \beta$. \blacksquare

Observation 6.2 *For the analysis above it suffices that, for every $j \in [n]$, the entries of column j of Z are only pairwise independent. Furthermore, appealing to Lemma 2.3 (concentration of k -wise independent random variables) instead of the Hoeffding bound, it suffices that, for every $v \in \mathcal{V}$, the entries of row v of Z are only k -wise independent, for $k = 3 \ln(d/\beta)$.*

6.1.2. A SIMPLE HEAVY HITTERS PROTOCOL

We next give a (simplified version) of our construction that uses a frequency oracle protocol with worst-case error τ in order to identify the heavy-hitters. Let $h : \mathcal{V} \rightarrow [T]$ be a (publicly known) random hash function, mapping domain elements into $[T]$, where T will be set later. We will use h in order to identify the heavy-hitters. To that end, let $v^* \in \mathcal{V}$ denote such a heavy-hitter, appearing at least 2τ times in the database S , and denote $t^* = h(v^*)$. Assuming that T is big enough, w.h.p. we will have that v^* is the only input element (from S) that is mapped (by h) into the hash value t^* . Assuming that this is indeed the case, we will now identify v^* bit by bit.

For $\ell \in [\log d]$, denote $S_\ell = (h(v_j), v_{j,\ell})_{j \in [n]}$, where $v_{j,\ell}$ is bit ℓ of v_j . That is, S_ℓ is a database over the domain $([T] \times \{0, 1\})$, where the row corresponding to user j is $(h(v_j), v_{j,\ell})$. Observe that every user can compute her own row locally. As v^* is a heavy-hitter, for every $\ell \in [\log d]$ we have that (t^*, v_ℓ^*) appears in S_ℓ at least 2τ times, where v_ℓ^* is bit ℓ of v^* . On the other hand, as we assumed that v^* is the only input element that is mapped into t^* we get that $(t^*, 1 - v_\ell^*)$ does not appear in S_ℓ at all. Recall that our frequency oracle has error at most τ , and hence, we can use it to accurately determine the bits of v^* . The details are given in algorithm `SuccinctHist`.

Algorithm 6 `SuccinctHist`

Public randomness: Random hash function $h : \mathcal{V} \rightarrow [T]$.

Random partition of $[n]$ into $\log d$ subsets $I_1, \dots, I_{\log d}$.

Setting: Each player $j \in [n]$ holds a value $v_j \in \mathcal{V}$. Define $S = (v_1, \dots, v_n)$.

For $\ell \in [\log d]$, let $S_\ell = (h(v_j), v_{j,\ell})_{j \in I_\ell}$, where $v_{j,\ell}$ is bit ℓ of v_j .

That is, S_ℓ is a database over the domain $[T] \times \{0, 1\}$.

Input: Oracle access to S and S_ℓ for $\ell \in [\log d]$.

1. For $\ell \in [\log d]$, use `ExplicitHist`(S_ℓ) with $\frac{\epsilon}{2}$ to get $a_\ell(t, b)$ for all $(t, b) \in [T] \times \{0, 1\}$.
 2. For $t \in [T]$, define $\hat{v}_t \in \mathcal{V}$, where bit ℓ of \hat{v}_t is $\hat{v}_{t,\ell} = \arg\max\{a_\ell(t, 0), a_\ell(t, 1)\}$.
 3. Use `ExplicitHist`(S) with privacy parameter $\frac{\epsilon}{2}$ to obtain $a(\hat{v}_t)$ for all $t \in [T]$.
 4. Return list $L = \{(\hat{v}_t, a(\hat{v}_t)) : t \in [T]\}$.
-

Lemma 6.3 *Let $\epsilon \leq 1$, denote $w \triangleq 32 \log(d) \log(16 \log d) + \frac{48}{\epsilon} \sqrt{2n \log d \cdot \ln(64 \log d)}$, and set $T = \frac{32n}{w}$. Algorithm `SuccinctHist` returns a list L of length T satisfying:*

1. *With probability $1 - \beta$, for every $(v, a) \in L$ we have that $|a - f_S(v)| \leq \frac{6}{\epsilon} \sqrt{n \cdot \ln(4T/\beta)}$.*
2. *For every $v \in \mathcal{V}$ s.t. $f_S(v) \geq w$, with probability $1/2$ we have that v is in L .*

Remark 3 *In this simplified version of our protocol, we successfully recover a heavy-hitter only if we recover (exactly) all of its bits. As will be illustrated in Section 6.2.2, this requirement can be relaxed using an error correction code s.t. in order to recover a heavy-hitter v^* it suffices to recover correctly only part of its bits.*

Proof Item 1 of the lemma follows directly from Lemma 6.1. We now prove item 2. Assuming that $n \geq 12 \log(d) \log(12 \log d)$, by the Chernoff bound, with probability at least $7/8$ (over partitioning $[n]$ into subsets $I_1, \dots, I_{\log d}$), for every $\ell \in [\log d]$ we have that $\frac{n}{2 \log d} \leq |I_\ell| \leq \frac{2n}{\log d}$. We continue the analysis assuming that this is the case.

Fix $v^* \in \mathcal{V}$ s.t. $f_S(v^*) \geq w$, and consider the following good event (over sampling h):

$$\mathbf{Event } E_1 : |\{v \in S : v \neq v^* \text{ and } h(v) = h(v^*)\}| \leq w/4.$$

Event E_1 states that v^* is mapped (by the hash function h) into a cell without too many collisions with different input elements. Denote $t^* = h(v^*)$. While the multiplicity of v^* in S is at least w , event E_1 states that other than v^* there are at most $w/4$ elements which are mapped into t^* . That is, v^* dominates the cell t^* . We first show that if E_1 occurs, then w.h.p. v^* is in the list L .

Asserting that $w \geq 32 \log(d) \log(16 \log d)$, by the Chernoff bound we get that with probability $7/8$ (over partitioning $[n]$ into subsets $I_1, \dots, I_{\log d}$), for every $\ell \in [\log d]$ we have that

$$f_{S_\ell}(h(v^*), v_\ell^*) \geq f_{S_\ell}(h(v^*), 1 - v_\ell^*) + \frac{w}{4 \log d}.$$

If that is the case, then by the properties of algorithm `ExplicitHist`, for

$$w \geq \frac{48}{\epsilon} \sqrt{2n \log d \cdot \ln(64 \log d)},$$

with probability at least $1 - \frac{1}{8 \log d}$ we have that $\hat{v}_{t^*, \ell} = v_\ell^*$, where $t^* = h(v^*)$. Using the union bound, this holds simultaneously for all $\ell \in [\log d]$ with probability at least $7/8$, in which case $v^* = \hat{v}_{t^*}$ is in the list L .

It remains to show that Event E_1 occurs with high probability. To that end, observe that

$$\mathbb{E}_h [|\{v \in S : v \neq v^* \text{ and } h(v) = h(v^*)\}|] = \sum_{v \in S : v \neq v^*} \mathbb{E}_h [\mathbb{1}_{h(v)=h(v^*)}] \leq \frac{n}{T}.$$

Thus, by Markov's inequality, we have that

$$\Pr \left[|\{v \in S : v \neq v^* \text{ and } h(v) = h(v^*)\}| \geq \frac{8n}{T} \right] \leq \frac{1}{8}.$$

Setting $T = \frac{32n}{w}$ completes the proof. ■

6.2. Reducing Space and Time Complexities

In this section we give the full details for algorithm `Bitstogram`.

6.2.1. SPACE AND TIME EFFICIENT FREQUENCY ORACLE

We modify algorithm `ExplicitHist` to obtain our frequency oracle with improved time and space complexity – algorithm `Hashtogram`. Intuitively, this algorithm is obtained by applying algorithm `ExplicitHist` after *hashing* every domain element using a (pairwise independent) hash function (however, to boost accuracy, we use *several* hash functions). This will mean that the random matrix Z we use only has $\approx \sqrt{n}$ rows, instead on $|\mathcal{V}|$ rows. As we will see, this will allow us to efficiently implement the algorithm.

Algorithm 7 Hashtogram

Public randomness: Random partition of $[n]$ into R subsets I_1, \dots, I_R .
 Random hash functions h_1, \dots, h_R mapping \mathcal{V} to $[T]$.
 Uniformly random matrix $Z \in \{\pm 1\}^{T \times n}$.

Setting: Each player $j \in [n]$ holds a value $v_j \in \mathcal{V}$. Define $S = (v_1, \dots, v_n)$.
 Define $\tilde{S} = (\tilde{v}_1, \dots, \tilde{v}_n)$, where $\tilde{v}_j = Z[h_r(v_j), j]$ and h_r is s.t. $j \in I_r$.

Input: Oracle access to \tilde{S} .

1. For $j \in [n]$, get user- j 's 1-bit report: $y_j \leftarrow \mathcal{R}(\tilde{v}_j)$.
 2. For every $(r, t) \in [R] \times [T]$ compute $a_r(t) = \frac{e^\epsilon + 1}{e^\epsilon - 1} \cdot \sum_{j \in I_r} y_j \cdot Z[t, j]$.
 3. On input $v \in \mathcal{V}$, return $a(v) = R \cdot \text{Median}\{a_1(h_1(v)), \dots, a_R(h_R(v))\}$, and wait for the next input.
-

Lemma 6.4 *Algorithm Hashtogram satisfies ϵ -LDP.*

Lemma 6.5 *Let $\epsilon \leq 1$, and fix a subset $V \subseteq \mathcal{V}$ of size $d' \leq d$ to be queried to algorithm Hashtogram. Let algorithm Hashtogram be executed with $R \geq 132 \log(4d'/\beta)$, and $T \geq \epsilon^2 \cdot \log(d'/\beta) + \epsilon \cdot \sqrt{n/\log(d'/\beta)}$, and $n \geq 8R \log(8d'/\beta)$. With probability at least $1 - \beta$, algorithm Hashtogram answers every $v \in V$ with $a(v)$ satisfying:*

$$|a(v) - f_S(v)| \leq \frac{27}{\epsilon} \cdot \sqrt{nR \log\left(\frac{2Rd'}{\beta}\right)}.$$

Observe that the error in the lemma is sub-optimal, as the optimal error behaves like $\frac{1}{\epsilon} \sqrt{n \log d}$. However, we will only use Lemma 6.5 with constant d' and constant β , and hence, will not be effected by this issue. A similar analysis (for the same algorithm) gives better bounds for other settings of parameters. Specifically,

Lemma 6.6 *Let $\epsilon \leq 1$. Fix a subset $V \subseteq \mathcal{V}$ of size $d' \leq d$ to be queried to algorithm Hashtogram. Let algorithm Hashtogram be executed with $R \geq 300 \log(12nd'/\beta)$ and $n \geq 43R$ and $T \geq \epsilon \cdot \sqrt{n/\log(nd'/\beta)}$. With probability at least $1 - \beta$, algorithm Hashtogram answers every $v \in V$ with $a(v)$ satisfying:*

$$|a(v) - f_S(v)| \leq \frac{400}{\epsilon} \cdot \sqrt{n \log\left(\frac{12nd'}{\beta}\right)}.$$

As the analysis of the two lemmas are very similar, we only present the proof of Lemma 6.6. The proof of Lemma 6.5 appears in Section B for completeness.

Proof [Proof of Lemma 6.6] Consider the following good event:

Event E_1 (over sampling h_1, \dots, h_R):

For every query $v^* \in V$ there exists a subset $R_1^{v^*} \subseteq [R]$ of size $|R_1^{v^*}| \geq \frac{7}{8}R$ s.t. for every $r^* \in R_1^{v^*}$ it holds that $|\{v \in S : v \neq v^* \text{ and } h_{r^*}(v) = h_{r^*}(v^*)\}| \leq \frac{16n}{T}$.

Event E_1 states that for at least $7R/8$ of the hash functions, we have that v^* is mapped into a cell without too many collisions with different input elements. Informally, for every single hash function h_r , algorithm **Hashtogram** estimates the number of occurrences of $h_r(v^*)$ in S . Hence, if event E_1 occurs, then most of the estimations result in accurate answers. We start by showing that event E_1 happens with high probability. To that end, fix $v^* \in V$ and fix $r^* \in [R]$. We have that

$$\mathbb{E}_{h_{r^*}} [|\{v \in S : v \neq v^* \text{ and } h_{r^*}(v) = h_{r^*}(v^*)\}|] = \sum_{v \in S : v \neq v^*} \mathbb{E}_{h_{r^*}} [\mathbb{1}_{h_{r^*}(v) = h_{r^*}(v^*)}] \leq \frac{n}{T}.$$

Thus, by Markov's inequality, we have that

$$\Pr_{h_{r^*}} \left[|\{v \in S : v \neq v^* \text{ and } h_{r^*}(v) = h_{r^*}(v^*)\}| \geq \frac{16n}{T} \right] \leq \frac{1}{16}.$$

As the hash functions are independent from each other, for $R \geq 48 \ln(\frac{d'}{\beta})$, by the Chernoff bound we get that with probability at least $1 - \beta/d'$ (over sampling h_1, \dots, h_R) there exists a subset $R_1^{v^*} \subseteq [R]$ of size $|R_1^{v^*}| \geq \frac{7}{8}R$ s.t. for every $r^* \in R_1^{v^*}$ it holds that

$$|\{v \in S : v \neq v^* \text{ and } h_{r^*}(v) = h_{r^*}(v^*)\}| \leq \frac{16n}{T}.$$

Using the union bound over every $v^* \in V$, we have that event E_1 happens with probability at least $1 - \beta$.

Event E_2 (over partitioning $[n]$ into I_1, \dots, I_R):

There exists a subset $R_2 \subseteq [R]$ of size $|R_2| \geq \frac{7}{8}R$ s.t. for every $r \in R_2$ it holds that $\frac{n}{2R} \leq |I_r| \leq \frac{2n}{R}$.

Following Theorem 2.4 (the Poisson approximation), we analyze event E_2 in the Poisson case. To that end, let $\hat{I}_1, \dots, \hat{I}_R$ be independent Poisson random variables with mean n/R . Now fix $r \in R$. Using a tail bound for the Poisson distribution (see Theorem 2.5), assuming that $n \geq 43R$ we have that $\Pr[\frac{n}{2R} \leq \hat{I}_r \leq \frac{2n}{r}] \geq \frac{99}{100}$. As $\hat{I}_1, \dots, \hat{I}_R$ are independent, assuming that $R \geq 300 \ln(\frac{3n}{\beta})$, by the Chernoff bound we get that event E_2 happens with probability at least $1 - \frac{\beta}{3n}$ in the Poisson case. Hence, by Theorem 2.4, event E_2 happens with probability at least $1 - \beta$.⁸

For every $r \in [R]$, let $S_r = (v_j)_{j \in I_r}$ denote a database containing the data of all users j s.t. $j \in I_r$. Also for $v^* \in V$ and $r \in [R]$ denote $S^{r,v^*} \triangleq \{v \in S : h_r(v) = h_r(v^*)\}$. That is, $|S^{r,v^*}|$ is the number of users j s.t. $h_r(v_j) = h_r(v^*)$. Furthermore, for $v^* \in V$ and $r \in [R]$ denote $I_r^{v^*} \triangleq \{v \in S_r : h_r(v) = h_r(v^*)\}$. That is, $|I_r^{v^*}|$ is the number of users j s.t. $j \in I_r$ and $h_r(v_j) = h_r(v^*)$. Observe that $|S^{r,v^*}| \geq f_S(v^*)$ and that $|I_r^{v^*}| \geq f_{S_r}(v^*)$.

8. Alternatively, instead of the requirement on R , it suffices to ensure that $n \geq O(R \log(\frac{R}{\beta}))$. In such a case, by the Chernoff bound, for every fixed $r \in [R]$ we will have that $\frac{n}{2R} \leq |I_r| \leq \frac{2n}{r}$ with probability at least $1 - \frac{\beta}{R}$. Hence, using the union bound, this happens simultaneously for all $r \in [R]$ with probability at least $1 - \beta$. Observe that at any case we already require that $R \geq O(\log(\frac{d'}{\beta}))$ for Event E_1 .

Event E_3 (over partitioning $[n]$ into I_1, \dots, I_R):

For every query $v^* \in V$ there exists a subset $R_3^{v^*} \subseteq [R]$ of size $|R_3^{v^*}| \geq \frac{9}{10}R$ s.t. for every $r^* \in R_3^{v^*}$ it holds that $|R \cdot |I_r^{v^*}| - |S^{r, v^*}| | \leq \sqrt{8Rn}$

We analyze event E_3 in the Poisson case. To that end, fix $v^* \in V$, and let $\hat{I}_1^{v^*}, \dots, \hat{I}_R^{v^*}$ be independent Poisson random variables with mean $\frac{|S^{r, v^*}|}{R}$. Now fix $r \in [R]$. Using a tail bound for the Poisson distribution (see Theorem 2.5), with probability at least $19/20$ we have that

$$\left| R \cdot \hat{I}_r^{v^*} - |S^{r, v^*}| \right| \leq \sqrt{8Rn}. \quad (1)$$

As $\hat{I}_1^{v^*}, \dots, \hat{I}_R^{v^*}$ are independent, assuming that $R \geq 300 \ln(\frac{3nd'}{\beta})$, by the Chernoff bound we get that with probability at least $1 - \frac{\beta}{3nd'}$, Inequality (1) holds for at least $9/10$ choices of $r \in [R]$. By Theorem 2.4 (the Poisson approximation), with probability at least $1 - \frac{\beta}{d'}$, this is also the case for the random variables $|I_r^{v^*}|$. That is, with probability at least $1 - \frac{\beta}{d'}$, there exists a subset $R_3^{v^*} \subseteq [R]$ of size $|R_3^{v^*}| \geq \frac{9}{10}R$ s.t. for every $r^* \in R_3^{v^*}$ it holds that

$$\left| R \cdot |I_{r^*}^{v^*}| - |S^{r^*, v^*}| \right| \leq \sqrt{8Rn}.$$

Using the union bound over every choice of $v^* \in V$, we get that event E_3 happens with probability at least $1 - \beta$.⁹

Event E_4 (over sampling Z and the coins of the local randomizers):

For every query $v^* \in V$ there exists a subset $R_4^{v^*} \subseteq [R]$ of size $|R_4^{v^*}| \geq \frac{7}{8} \cdot \frac{9}{10}R$ s.t. for every $r^* \in R_4^{v^*}$ it holds that $|R \cdot a_{r^*}(h_{r^*}(v^*)) - R \cdot |I_{r^*}^{v^*}| | \leq \frac{e^\epsilon + 1}{e^\epsilon - 1} \cdot \sqrt{11nR}$.

For $v^* \in V$ and $r \in [R]$ denote $c_r(v^*) = \sum_{j \in I_r} y_j \cdot Z[h_r(v^*), j]$, and recall that algorithm **Hashtogram** answers the query $v^* \in \mathcal{V}$ with $a(v^*) = R \cdot \frac{e^\epsilon + 1}{e^\epsilon - 1} \cdot \text{Median}\{c_r(v^*)\}_{r \in [R]}$. Fix $v^* \in V$ and $r \in R_2$ (where $R_2 \subseteq [R]$ is the subset from event E_2). We now analyze the expectation of $c_r(v^*)$:

$$\begin{aligned} \mathbb{E}[c_r(v^*)] &= \sum_{j \in I_r} \mathbb{E}[y_j \cdot Z[h_r(v^*), j]] \\ &= \sum_{j \in I_r: h_r(v_j) = h_r(v^*)} \mathbb{E}[y_j \cdot Z[h_r(v^*), j]] + \sum_{j \in I_r: h_r(v_j) \neq h_r(v^*)} \mathbb{E}[y_j \cdot Z[h_r(v^*), j]] \\ &= \sum_{j \in I_r: h_r(v_j) = h_r(v^*)} \mathbb{E}[y_j \cdot Z[h_r(v^*), j]] + \sum_{j \in I_r: h_r(v_j) \neq h_r(v^*)} \mathbb{E}[y_j] \cdot \mathbb{E}[Z[h_r(v^*), j]] \\ &= |\{v \in S_r : h_r(v) = h_r(v^*)\}| \cdot \frac{e^\epsilon - 1}{e^\epsilon + 1} \triangleq |I_r^{v^*}| \cdot \frac{e^\epsilon - 1}{e^\epsilon + 1} \end{aligned}$$

9. Alternatively, instead of the requirement on R , it suffices to ensure that $n \geq O(R \log(\frac{d'R}{\beta}))$, in which such the analysis follows from the Chernoff bound.

That is, $c_r(v^*)$ can be expressed as two sums of ± 1 independent random variables: $|I_r^{v^*}|$ variables with expectation $\frac{e^\epsilon - 1}{e^\epsilon + 1}$, and $(|I_r| - |I_r^{v^*}|)$ variables with expectation 0 (recall that by event E_2 we have $\frac{n}{2R} \leq |I_r| \leq \frac{2n}{R}$). Using the Hoeffding bound, with probability at least $43/44$ we have that $\left| c_r(v^*) - \frac{e^\epsilon - 1}{e^\epsilon + 1} \cdot |I_r^{v^*}| \right| \leq \sqrt{11n/R}$. That is,

$$\left| R \cdot a_r(h_r(v^*)) - R \cdot |I_r^{v^*}| \right| \leq \frac{e^\epsilon + 1}{e^\epsilon - 1} \cdot \sqrt{11nR}. \quad (2)$$

Fix $v^* \in V$, and observe that the above sums are independent for different values of r . Hence, using the Chernoff bound and asserting that $R \geq 150 \ln(d'/\beta)$, for that fixed $v^* \in M$, with probability at least $1 - \beta/d'$ we have that Inequality (2) holds for at least $7R/8$ choices of $r \in R_1$. Using the union bound, with probability at least $1 - \beta$, this is true for every $v^* \in V$ simultaneously. That is, event E_4 happens with probability at least $1 - \beta$.

We are now ready to complete the proof, by showing that the guarantees of the lemma holds whenever events E_1, E_2, E_3, E_4 occur. Fix $v^* \in V$. Combining events E_3 and E_4 , we get that for every $r \in R_3^{v^*} \cap R_4^{v^*}$

$$\left| R \cdot a_r(h_r(v^*)) - |S^{r,v^*}| \right| \leq \frac{e^\epsilon + 1}{e^\epsilon - 1} \cdot \sqrt{11nR} + \sqrt{8Rn}. \quad (3)$$

Recall that for every $r \in [R]$ we have that $|S^{r,v^*}| \geq f_S(v^*)$. Furthermore, by event E_1 , for every $r \in R_1^{v^*}$ we have that $|S^{r,v^*}| \leq f_S(v^*) + \frac{16n}{T}$. Hence, for every $r \in R_1^{v^*} \cap R_3^{v^*} \cap R_4^{v^*}$ we have that

$$\left| R \cdot a_r(h_r(v^*)) - f_S(v^*) \right| \leq \frac{e^\epsilon + 1}{e^\epsilon - 1} \cdot \sqrt{11nR} + \sqrt{8Rn} + \frac{16n}{T} \triangleq \text{error}(v^*).$$

That is, for every $r \in R_1^{v^*} \cap R_3^{v^*} \cap R_4^{v^*}$ we have that $R \cdot a_r(h_r(v^*))$ is accurate up to $\text{error}(v^*)$. As $|R_1^{v^*} \cap R_3^{v^*} \cap R_4^{v^*}| \geq \frac{9}{16}R$, and as algorithm `Hashtogram` answers v^* with $a(v^*)$ chosen as the median of $\{R \cdot a_r(h_r(v^*))\}$, we get that $|a(v^*) - f_S(v^*)| \leq \text{error}(v^*)$. This holds for every $v^* \in M$. \blacksquare

Processing Memory. Algorithm `Hashtogram` maintains (on step 2) $R \cdot T$ sums of at most n bits. This requires $O(R \cdot T \cdot \log n)$ bits for processing memory.

Runtime. Observe that a direct implementation of (step 2 of) algorithm `Hashtogram` consists of summing a total of $T \approx \sqrt{n}$ bits per user, and hence results in a runtime of $\approx n^{1.5}$. As we next explain, this can be reduced to $\approx n$. First observe that for the analysis of Lemma 6.6 (specifically, for the analysis of Event E_4), it suffices that, for every $j \in [n]$, the entries of column j of Z are only pairwise independent. That is, each column of Z consists of $T \approx \sqrt{n}$ pairwise independent bits. We can represent such a column using $\log T \approx \log \sqrt{n}$ bits, in which case there are at most $T \approx \sqrt{n}$ choices for the columns of Z (see, e.g., Construction 3.18 of Vadhan (2012)). So, even though the matrix Z contains n columns, it has at most $T \approx \sqrt{n}$ distinct columns. Let us denote those distinct columns as z_1, \dots, z_T , where $z_\gamma[t]$ denotes the bit in position t in this column. We will write $Z[\cdot, j] = z_\gamma$ to indicate that the j^{th} column of Z is z_γ . With this notation, we can restate $a_r(t)$ (computed on step 2 of algorithm `Hashtogram`) as follows.

$$a_r(t) = \frac{e^\epsilon + 1}{e^\epsilon - 1} \cdot \sum_{j \in I_r} y_j \cdot Z[t, j] = \frac{e^\epsilon + 1}{e^\epsilon - 1} \cdot \sum_{1 \leq \gamma \leq T} \left(\sum_{j \in I_r \text{ s.t. } Z[\cdot, j] = z_\gamma} y_j \right) \cdot z_\gamma[t].$$

Thus, we can implement step 2 of algorithm **Hashtogram** by first computing $\left(\sum_{j \in I_r \text{ s.t. } Z[\cdot, j] = z_\gamma} y_j \right)$ for every $1 \leq \gamma \leq T$ and $1 \leq r \leq R$ (this amounts to summing a total of n bits, and can be done in time $\approx n$). Afterwards, for every choice of (r, t) , computing $a_r(t)$ consists of summing $T \approx \sqrt{n}$ elements. Overall, step 2 of the algorithm can be executed in time $\approx R \cdot T \cdot T \approx n \log n$.

6.2.2. THE FULL PROTOCOL

We are now ready to present the full **Bitstogram** protocol. This protocol is obtained by repeating the strategy outlined in algorithm **SuccinctHist** to boost the success probability, and by using an error correction code to reduce the error (as we mentioned in Remark 3).

Algorithm 8 Bitstogram

Tool used: Binary code (Enc, Dec), where $\text{Enc} : \mathcal{V} \rightarrow \mathcal{V}'$, correcting $\frac{1}{8}$ -fraction of errors with constant rate, that is $\log d' = O(\log d)$, where $d = |\mathcal{V}|$ and $d' = |\mathcal{V}'|$.

Public randomness: Random partition of $[n]$ into $R \log d'$ subsets $I_{r,\ell}$ for $(r, \ell) \in [R] \times [\log d']$. Random hash functions h_1, \dots, h_R mapping \mathcal{V}' to $[T]$.

Setting: Each player $j \in [n]$ holds a value $v_j \in \mathcal{V}$. Define $S = (v_1, \dots, v_n)$.

For every $j \in [n]$ denote $c_j = \text{Enc}(v_j) \in \mathcal{V}'$.

For $(r, \ell) \in [R] \times [\log d']$, let $S_{r,\ell} = (h_r(c_j), c_{j,\ell})_{j \in I_{r,\ell}}$, where $c_{j,\ell}$ is bit ℓ of c_j .

That is, $S_{r,\ell}$ is a database over the domain $[T] \times \{0, 1\}$.

Input: Oracle access to S and to $S_{r,\ell}$ for every $(r, \ell) \in [R] \times [\log d']$.

1. For $(r, \ell) \in [R] \times [\log d']$, use **Hashtogram**($S_{r,\ell}$) with $\frac{\epsilon}{2}$ to get $\{a_{r,\ell}(t, b) : (t, b) \in [T] \times \{0, 1\}\}$.
 2. For $(r, t) \in [R] \times [T]$, define $\hat{c}_{r,t} \in \mathcal{V}'$, where bit ℓ of $\hat{c}_{r,t}$ is $\hat{c}_{r,t,\ell} = \text{argmax}\{a_{r,\ell}(t, 0), a_{r,\ell}(t, 1)\}$. That is, if $a_{r,\ell}(t, 0) \geq a_{r,\ell}(t, 1)$ then $\hat{c}_{r,t,\ell} = 0$, and otherwise $\hat{c}_{r,t,\ell} = 1$.
 3. For $(r, t) \in [R] \times [T]$, define $\hat{v}_{r,t} = \text{Dec}(\hat{c}_{r,t}) \in \mathcal{V}$.
 4. Use **Hashtogram**(S) with privacy parameter $\frac{\epsilon}{2}$ to obtain $a(\hat{v}_{r,t})$ for all $(r, t) \in [R] \times [T]$.
 5. Return list $L = \{(\hat{v}_{r,t}, a(\hat{v}_{r,t})) : (r, t) \in [R] \times [T]\}$.
-

Remark 4 *The execution of **Hashtogram** on step 4 is made using the parameters stated in Lemma 6.6, in order to obtain accurate answers for every fixture of n queries with probability $1 - \beta$. The executions of **Hashtogram** on step 1 are made using the parameters stated in Lemma 6.5, in order to obtain accurate answers for every fixture of two queries with probability $255/256$. Observe that every such instantiation of **Hashtogram** is queried $2T$ times, and hence, some of these queries might result in inaccurate answers. Nevertheless,*

as will be made clear later, due to our use of error correction code, these inaccurate answers will not effect the final outcome of the algorithm.

Lemma 6.7 *Algorithm Bitstogram satisfies ϵ -LDP.*

Lemma 6.8 *Let $\epsilon \leq 1$, and assume that $\log d \geq O(\log(n/\beta))$. Set $R = O(\log(1/\beta))$ and $T = O\left(\frac{\epsilon n}{\sqrt{R \log d}}\right)$. Algorithm Bitstogram returns a list L of length $R \cdot T$ satisfying:*

1. *With probability $1 - \beta$, for every $(v, a) \in L$ we have that $|a - f_S(v)| \leq O\left(\frac{1}{\epsilon} \sqrt{n \log(n/\beta)}\right)$.*
2. *With probability $1 - \beta$, for every $v \in \mathcal{V}$ s.t. $f_S(v) \geq O\left(\frac{1}{\epsilon} \sqrt{n \log(d) \log(\frac{1}{\beta})}\right)$, we have that v is in L .*

Remark 5 *The assumption in Lemma 6.8 that $\log d \geq O(\log(n/\beta))$ is without loss of generality, as otherwise the universe size d is small enough to allow the server to run in time linear in d , which makes the problem much easier. Specifically, if $d < \sqrt{n}$ then we can instantiate the frequency oracle of Lemma 6.1, and query it for every domain element. As $d < \sqrt{n}$, this can be executed in time $\approx n$ (the runtime analysis is similar to the one in Section 6.2.1). Otherwise, if $d \geq \sqrt{n}$ then we already have that $\log d \geq O(\log n)$, and (if necessary) we can pad the representation of domain elements to satisfy the assumption that $\log d \geq O(\log(n/\beta))$.*

Proof [Proof of Lemma 6.8] Item 1 of the lemma follows directly from Lemma 6.6. We now prove item 2. Consider the following good event (over sampling h_1, \dots, h_R):

Event E_1 (over sampling h_1, \dots, h_R):

There exists a subset $R_1 \subseteq [R]$ of size $|R_1| \geq \frac{7}{8}R$ s.t. for every $r^* \in R_1$ and for every $v^* \in S$ satisfying $f_S(v^*) \geq \frac{n^{1.5}}{T}$ it holds that $|\{v \in S : v \neq v^* \text{ and } h_{r^*}(\text{Enc}(v)) = h_{r^*}(\text{Enc}(v^*))\}| \leq \frac{16n^{1.5}}{T}$.

We start by showing that event E_1 happens with high probability. To that end, fix $v^* \in S$ and fix $r^* \in [R]$. We have that

$$\begin{aligned} & \mathbb{E}_{h_{r^*}} [|\{v \in S : v \neq v^* \text{ and } h_{r^*}(\text{Enc}(v)) = h_{r^*}(\text{Enc}(v^*))\}|] \\ &= \sum_{v \in S : v \neq v^*} \mathbb{E}_{h_{r^*}} [\mathbb{1}_{h_{r^*}(\text{Enc}(v)) = h_{r^*}(\text{Enc}(v^*))}] \leq \frac{n}{T}. \end{aligned}$$

Thus, by Markov's inequality, we have that

$$\Pr_{h_{r^*}} \left[|\{v \in S : v \neq v^* \text{ and } h_{r^*}(\text{Enc}(v)) = h_{r^*}(\text{Enc}(v^*))\}| \geq \frac{16n^{1.5}}{T} \right] \leq \frac{1}{16\sqrt{n}}.$$

Assuming that $T \leq n$, there could be at most \sqrt{n} “heavy” elements v^* satisfying $f_S(v^*) \geq \frac{n^{1.5}}{T}$. Hence, using the union bound,

$$\Pr_{h_{r^*}} \left[\begin{array}{c} \exists v^* \text{ s.t. } f_S(v^*) \geq \frac{n^{1.5}}{T} \\ \text{and} \\ |\{v \in S : v \neq v^* \text{ and } h_{r^*}(\text{Enc}(v)) = h_{r^*}(\text{Enc}(v^*))\}| \geq \frac{16n^{1.5}}{T} \end{array} \right] \leq \frac{1}{16}.$$

As the hash functions are independent from each other, for $R \geq 48 \ln(\frac{1}{\beta})$, by the Chernoff bound we get that with probability at least $1 - \beta$ (over sampling h_1, \dots, h_R) there exists a subset $R_1 \subseteq [R]$ of size $|R_1| \geq \frac{7}{8}R$ s.t. for every v^* satisfying $f_S(v^*) \geq \frac{n^{1.5}}{T}$ and every $r^* \in R_1$ it holds that

$$|\{v \in S : v \neq v^* \text{ and } h_{r^*}(\text{Enc}(v)) = h_{r^*}(\text{Enc}(v^*))\}| \leq \frac{16n^{1.5}}{T}.$$

That is, event E_1 happens with probability at least $1 - \beta$. We continue the analysis assuming that event E_1 occurs.

Event E_2 (over partitioning $[n]$ into $\{I_{r,\ell}\}$):

There exists a subset $R_2 \subseteq [R]$ of size $|R_2| \geq \frac{7}{8}R$ s.t. for every $r^* \in R_2$, there exist at least $\frac{31}{32} \log d'$ choices for $\ell^* \in [\log d']$ for which $|I_{r^*,\ell^*}| \leq \frac{2n}{R \log d'}$.

Following Theorem 2.4 (the Poisson approximation), we analyze event E_2 in the Poisson case. To that end, let $\tilde{I}_{1,1}, \dots, \tilde{I}_{R, \log d'}$ be independent Poisson random variables with mean $\frac{n}{R \log d'}$. Now fix $(r, \ell) \in [R] \times [\log d']$. Using a tail bound for the Poisson distribution (see Theorem 2.5), assuming that $n \geq 10R \log d'$ we have that $\Pr[\tilde{I}_{r,\ell} \leq \frac{2n}{R \log d'}] \geq \frac{99}{100}$. As $\tilde{I}_{1,1}, \dots, \tilde{I}_{R, \log d'}$ are independent, assuming that $R \log d' \geq 300 \ln(\frac{3n}{\beta})$, by the Chernoff bound we get that with probability at least $1 - \frac{\beta}{3n}$ there are at least $\frac{98}{100} R \log d'$ choices for $(r, \ell) \in [R] \times [\log d']$ s.t. $\tilde{I}_{r,\ell} \leq \frac{2n}{R \log d'}$. Hence, by Theorem 2.4, with probability at least $1 - \beta$ there are at least $\frac{98}{100} R \log d'$ choices for $(r, \ell) \in [R] \times [\log d']$ s.t. $|I_{r,\ell}| \leq \frac{2n}{R \log d'}$. If that is the case, then there must be at least $\frac{7R}{8}$ choices for $r \in [R]$ for which

$$\left| \left\{ \ell \in [\log d'] : |I_{r,\ell}| \leq \frac{2n}{R \log d'} \right\} \right| \geq \frac{31}{32} \log d'.$$

That is, $\Pr[E_2] \geq 1 - \beta$.

Event E_3 (over partitioning $[n]$ into $\{I_{r,\ell}\}$):

For every $v^* \in S$ s.t. $f_S(v^*) \geq \frac{n^{1.5}}{T}$ there exists a subset $R_3^{v^*} \subseteq [R]$ of size $|R_3^{v^*}| \geq \frac{7}{8}R$ s.t. for every $r^* \in R_3^{v^*}$ there exist at least $\frac{31}{32} \log d'$ choices for $\ell^* \in [\log d']$ for which $|\{j \in I_{r^*,\ell^*} : v_j = v^*\}| \geq \frac{f_S(v^*)}{2R \log d'}$.

We analyze event E_3 in the Poisson case. To that end, fix $v^* \in S$ s.t. $f_S(v^*) \geq \frac{n^{1.5}}{T}$, and let $\tilde{I}_{1,1}^{v^*}, \dots, \tilde{I}_{R, \log d'}^{v^*}$ be independent Poisson random variables with mean $\frac{f_S(v^*)}{R \log d'}$. Now fix $(r, \ell) \in [R] \times [\log d']$. Using a tail bound for the Poisson distribution (see Theorem 2.5), assuming that $\frac{n^{1.5}}{T} \geq 37R \log d'$ we have that $\Pr[\tilde{I}_{r, \ell} \geq \frac{f_S(v^*)}{2R \log d'}] \geq \frac{99}{100}$. As $\tilde{I}_{1,1}^{v^*}, \dots, \tilde{I}_{R, \log d'}^{v^*}$ are independent, assuming that $R \log d' \geq 300 \ln(\frac{3n^2}{\beta})$, by the Chernoff bound we get that with probability at least $1 - \frac{\beta}{3n^2}$ there are at least $\frac{98}{100} R \log d'$ choices for $(r, \ell) \in [R] \times [\log d']$ s.t. $\tilde{I}_{r, \ell} \geq \frac{f_S(v^*)}{2R \log d'}$. Hence, by Theorem 2.4, with probability at least $1 - \frac{\beta}{n}$ there are at least $\frac{98}{100} R \log d'$ choices for $(r, \ell) \in [R] \times [\log d']$ s.t. $|\{j \in I_{r^*, \ell^*} : v_j = v^*\}| \geq \frac{f_S(v^*)}{2R \log d'}$. If that is the case, then there must be at least $\frac{7R}{8}$ choices for $r \in [R]$ for which

$$\left| \left\{ \ell \in [\log d'] : |\{j \in I_{r^*, \ell^*} : v_j = v^*\}| \geq \frac{f_S(v^*)}{2R \log d'} \right\} \right| \geq \frac{31}{32} \log d'.$$

Using the union bound, this holds simultaneously for every such v^* with probability at least $1 - \beta$. That is, $\Pr[E_3] \geq 1 - \beta$.

Event E_4 (over partitioning $[n]$ into $\{I_{r, \ell}\}$):

For every $v^* \in S$ s.t. $f_S(v^*) \geq \frac{n^{1.5}}{T}$ there exists a subset $R_4^{v^*}$ of size $|R_4^{v^*}| \geq \frac{7}{8} \cdot \frac{7}{8} R$ s.t. for every $r^* \in R_4^{v^*}$ there exist at least $\frac{31}{32} \log d'$ choices for $\ell^* \in [\log d']$ for which

$$|\{j \in I_{r^*, \ell^*} : v_j \neq v^* \text{ and } h_{r^*}(\text{Enc}(v_j)) = h_{r^*}(\text{Enc}(v^*))\}| \leq \frac{32n^{1.5}}{RT \log d'}.$$

We analyze event E_4 in the Poisson case. To that end, fix $v^* \in S$ s.t. $f_S(v^*) \geq \frac{n^{1.5}}{T}$. For $r \in [R]$ denote

$$\text{Col}_r(v^*) = |\{v \in S : v \neq v^* \text{ and } h_r(\text{Enc}(v)) = h_r(\text{Enc}(v^*))\}|,$$

and recall that by event E_1 , for $r \in R_1$ we have that $\text{Col}_r(v^*) \leq \frac{16n^{1.5}}{T}$.

Let $\tilde{I}_{1,1}^{v^*}, \dots, \tilde{I}_{R, \log d'}^{v^*}$ be independent Poisson random variables with mean $\frac{\text{Col}_r(v^*)}{R \log d'}$. Now fix $(r, \ell) \in R_1 \times [\log d']$. Using a tail bound for the Poisson distribution (see Theorem 2.5), assuming that $T \leq \frac{n^{1.5}}{R \log d'}$ we have that $\Pr[\tilde{I}_{r, \ell} \leq \frac{32n^{1.5}}{RT \log d'}] \geq \frac{99}{100}$. As $\tilde{I}_{1,1}^{v^*}, \dots, \tilde{I}_{R, \log d'}^{v^*}$ are independent, assuming that $|R_1| \log d' \geq 300 \ln(\frac{3n^2}{\beta})$, by the Chernoff bound we get that with probability at least $1 - \frac{\beta}{3n^2}$ there are at least $\frac{98}{100} |R_1| \log d'$ choices for $(r, \ell) \in R_1 \times [\log d']$ s.t. $\tilde{I}_{r, \ell} \leq \frac{32n^{1.5}}{RT \log d'}$. Hence, by Theorem 2.4, with probability at least $1 - \frac{\beta}{n}$ there are at least $\frac{98}{100} |R_1| \log d'$ choices for $(r, \ell) \in R_1 \times [\log d']$ s.t.

$$|\{j \in I_{r^*, \ell^*} : v_j \neq v^* \text{ and } h_{r^*}(\text{Enc}(v_j)) = h_{r^*}(\text{Enc}(v^*))\}| \leq \frac{32n^{1.5}}{RT \log d'}.$$

If that is the case, then there must be at least $\frac{7R_1}{8}$ choices for $r \in R_1$ for which

$$\left| \left\{ \ell \in [\log d'] : |\{j \in I_{r^*, \ell^*} : v_j = v^*\}| \leq \frac{32n^{1.5}}{RT \log d'} \right\} \right| \geq \frac{31}{32} \log d'.$$

Using the union bound, this holds simultaneously for every such v^* with probability at least $1 - \beta$. That is, $\Pr[E_4] \geq 1 - \beta$.

We are now ready to complete the proof. Fix $v^* \in S$ s.t. $f_S(v^*) \geq \frac{264n^{1.5}}{T}$. Let R_1, R_2, R_3^*, R_4^* be the sets from events E_1, E_2, E_3, E_4 , and observe that $|R_1 \cap R_2 \cap R_3^* \cap R_4^*| \geq \frac{3}{8}R$, and furthermore, for every $r^* \in (R_1 \cap R_2 \cap R_3^* \cap R_4^*)$ there exist at least $\frac{29}{32} \log d'$ choices for $\ell \in [\log d']$ for which

- (a) $|\{j \in I_{r^*, \ell} : v_j = v^*\}| \geq \frac{132n^{1.5}}{TR \log d'}$,
- (b) $|\{j \in I_{r^*, \ell} : v_j \neq v^* \text{ and } h_{r^*}(\text{Enc}(v_j)) = h_{r^*}(\text{Enc}(v^*))\}| \leq \frac{32n^{1.5}}{TR \log d'}$,
- (c) $|I_{r^*, \ell}| \leq \frac{2n}{R \log d'}$.

Denote $c^* = \text{Enc}(v^*)$. Observe that by items (a),(b) above, we have that

$$f_{S_{r^*, \ell}}(h_{r^*}(c^*), c_\ell^*) \geq f_{S_{r^*, \ell}}(h_{r^*}(c^*), 1 - c_\ell^*) + \frac{100n^{1.5}}{TR \log d'}.$$

Let us assume that $T \leq \frac{\epsilon n}{326\sqrt{R \log d'}}$. By the properties of algorithm `Hashtogram` (Lemma 6.5), For every r^*, ℓ^* satisfying items (a),(b),(c), algorithm `Hashtogram`(S_{r^*, ℓ^*}) ensures that with probability at least $\frac{255}{256}$ we have $\hat{c}_{r^*, t^*, \ell} = c_\ell^*$, where $t^* = h_{r^*}(v^*)$.

Now fix $r^* \in (R_1 \cap R_2 \cap R_3^* \cap R_4^*)$, and recall that the coins of `Hashtogram`($S_{r^*, \ell}$) are independent for different values of ℓ . Hence, by the Chernoff bound, for $\log d' \geq 850 \log(\frac{n}{\beta})$ we get that

$$\Pr \left[|\{\ell \in [\log d'] : \hat{c}_{r^*, t^*, \ell} = c_\ell^*\}| \geq \frac{9}{10} \log d' \right] \geq 1 - \frac{\beta}{n}.$$

That is, for our fixed v^* there exists an $r^* \in (R_1 \cap R_2 \cap R_3^* \cap R_4^*)$ s.t. with probability at least $1 - \frac{\beta}{n}$ we have that \hat{c}_{r^*, t^*} and $c^* = \text{Enc}(v^*)$ differ on at most $\frac{1}{10}$ -fraction of their bits. Using the union bound, with probability at least $1 - \beta$, this holds simultaneously for every v^* s.t. $f_S(v^*) \geq \frac{264n}{T}$. By the properties of the error correction code, in such a case, for every v^* s.t. $f_S(v^*) \geq \frac{264n^{1.5}}{T}$ we have that $\hat{v}_{r^*, t^*} = \text{Dec}(\hat{c}_{r^*, t^*}) = v^*$, and that v^* is in the list L . \blacksquare

Runtime. On step 1 algorithm `Bitstogram` instantiates $O(R \cdot \log d)$ copies of the frequency oracle `Hashtogram`. Every such instantiation takes time $\approx n$. Afterwards, the algorithm queries the oracles for a total of $O(2RT \log d) \approx n$ queries, each of which takes time $\tilde{O}(1)$. Thus, step 1 takes time $\approx n$ to compute. Steps 2 and 3 loop for every $(r, t) \in [R] \times [T]$, and take time $\approx RT \approx n$ to compute. Finally, step 4 instantiates `Hashtogram` (in time $\approx n$) and queries it for a total of $R \cdot T \approx n$ queries. Overall, algorithm `Bitstogram` runs in time $\approx n$.

Processing Memory. Recall that step 1 of algorithm `Bitstogram` requires querying algorithm `Hashtogram` for a total of $O(2RT \log d) \approx n$ queries. As we are aiming for an algorithm with processing memory $\approx \sqrt{n}$, we cannot store all of the answers in memory simultaneously. To resolve this issue, let us reorganize step 1 of algorithm `Bitstogram` as follows:

- 1a. For every $(r, \ell) \in [R] \times [\log d']$, invoke `Hashtogram` $(S_{r,\ell})$ with $\frac{\epsilon}{2}$.
- 1b. For every $(r, t) \in [R] \times [T]$, for every $\ell \in [\log d']$, query `Hashtogram` $(S_{r,\ell})$ to get $\{a_{r,\ell}(t, b) : b \in \{0, 1\}\}$.

Now, every one of the steps 1b,2,3,4 of the algorithm `Bitstogram` contains a loop over every choice of $(r, t) \in [R] \times [T]$, and we can group all of these steps together into one loop over $(r, t) \in [R] \times [T]$. If an iteration of this loop results in a value $a(\hat{v}_{r,t}) \leq \sqrt{n}$, we can simply ignore it (recall that the frequencies of domain elements that are not in the list L are estimated as zero, and that \sqrt{n} is less than the guaranteed bound on the error of the protocol, so this step does not affect our error bounds). As there could be at most \sqrt{n} elements with frequencies at least $\approx \sqrt{n}$, the necessary processing memory is only $\approx \sqrt{n}$.

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A. Missing Proofs from Section 5

A.1. Proof of Theorem 5.1

It is easy to see that each invocation of `LocalRnd` is $\epsilon/2$ -locally differentially private since conditioned on any realization of ℓ_i, j_i, r_i , for any pair of possible input items $v_i, v'_i \in \mathcal{V}$ to `LocalRnd` and any output bit b generated in step 7 of Algorithm 1, we have

$$\mathbb{P}[y_i = b \mid v_i] \leq e^{\epsilon/2} \mathbb{P}[y_i = b \mid v'_i]$$

Note that `LocalRnd` is invoked only twice for each user: once when `Final = 0` (the scanning/pruning phase of `TreeHist`) and another time when `Final = 1` (the final phase of `TreeHist`). Thus, it follows that protocol `TreeHist` is ϵ -differentially private.

A.2. Proof of Theorem 5.2

Let $\beta \in (0, 1)$ and η as defined in the theorem statement. Consider the pruning phase of `TreeHist`, that is, Steps 1 to 11 in Algorithm 3. Let γ be as set in Step 2. In this phase, `TreeHist` invokes `FreqOracle` once in every iteration of the outer `for` loop (over the levels of the tree) with the flag `Final = 0`. Consider any such iteration ℓ . Suppose, for now, that the size of `ChildSet(Prefixes)` passed to `FreqOracle` in that iteration is at most $2n/\eta$. (We will show that with probability at least $1 - \beta$ this condition is satisfied at all levels ℓ of the tree, i.e., it's a loop invariant). By invoking Lemma 5.3 with $\widehat{\mathcal{V}} = \text{ChildSet(Prefixes)}$, $\{\tilde{\mathcal{I}}_j : j \in [t]\} = \{\mathcal{I}_{\ell,j} : j \in [t]\}$, and $\gamma = t \log d = 110 \log(n/\beta) \log d$, we have that with probability at least $1 - \beta/\log(d)$, for every $\hat{v} \in \text{ChildSet(Prefixes)} : f(\hat{v}) > 3\eta$, `FreqOracle` gives an estimate $\hat{f}(\hat{v}) \geq 2\eta$, and for every $\hat{v} \in \text{ChildSet(Prefixes)} : f(\hat{v}) \leq \eta$, `FreqOracle` gives an estimate $\hat{f}(\hat{v}) < 2\eta$. Hence, Step 9 implies, that with probability at least $1 - \beta/\log d$, all $\hat{v} \in \text{ChildSet(Prefixes)}$ with true frequencies $f(\hat{v}) \geq 3\eta$ will proceed to the next iteration $\ell + 1$ and all those $\hat{v} \in \text{ChildSet(Prefixes)}$ with true frequencies $f(\hat{v}) < \eta$ will be pruned out. Since the number of nodes \hat{v} with true frequency $f(\hat{v}) \geq \eta$ cannot be more than n/η , then the number of surviving nodes in the next iteration $\ell + 1$ cannot be more than $2n/\eta$. Hence, this condition will be satisfied in the next iteration, and we can proceed in the same fashion. Note that when $\ell = 1$, the condition is trivially satisfied since there are only $2 < 2n/\eta$ nodes at that level. This induction argument shows that with probability at least $1 - \beta$, for every level $\ell \in [\log d]$, the surviving nodes at level ℓ correspond to prefixes whose true frequencies are not below η and include all prefixes whose true frequencies are above 3η . In particular,

with probability at least $1 - \beta$, all items in `SuccHist` satisfy these properties. This covers the proof of items 1 and 2 of Theorem 5.2.

Now, consider the final phase of `TreeHist`, that is, Steps 12 to 14 Algorithm 3. Let γ be as set in Step 12. In this phase, `TreeHist` invokes `FreqOracle` on the surviving nodes at the final level of the tree (the last update of `Prefixes`) and with input flag `Final` = 1. Now, by invoking Lemma 5.3 with $\hat{V} = \text{Prefixes}$, $\{\tilde{\mathcal{L}}_j : j \in [t]\} = \{\mathcal{L}_j : j \in [t]\}$, and $\gamma = t = 110 \log(n/\beta)$, we have that with probability at least $1 - \beta/\log(d)$, for every $\hat{v} \in \text{Prefixes}$, $|\hat{f}(\hat{v}) - f(\hat{v})| \leq 14\sqrt{nt}/\epsilon = O\left(\frac{\sqrt{n \log(n/\beta)}}{\epsilon}\right)$. This proves item 3 of the theorem.

B. Missing Proofs from Section 6

B.1. Proof of Lemma 6.5

Consider the following good event:

Event E_1 (over sampling h_1, \dots, h_R):

For every query $v^* \in V$ there exists a subset $R_1^{v^*} \subseteq [R]$ of size $|R_1^{v^*}| \geq \frac{7}{8}R$ s.t. for every $r^* \in R_1^{v^*}$ it holds that $|\{v \in S : v \neq v^* \text{ and } h_{r^*}(v) = h_{r^*}(v^*)\}| \leq \frac{16n}{T}$.

Event E_1 states that for at least $7R/8$ of the hash functions, we have that v^* is mapped into a cell without too many collisions with different input elements. Informally, for every single hash function h_r , algorithm `HashHist` estimates the number of occurrences of $h_r(v^*)$ in S . Hence, if event E_1 occurs, then most of the estimations result in accurate answers. We start by showing that event E_1 happens with high probability. To that end, fix $v^* \in V$ and fix $r^* \in [R]$. We have that

$$\mathbb{E}_{h_{r^*}} [|\{x \in S : v \neq v^* \text{ and } h_{r^*}(v) = h_{r^*}(v^*)\}|] = \sum_{v \in S: v \neq v^*} \mathbb{E}_{h_{r^*}} [\mathbb{1}_{h_{r^*}(v) = h_{r^*}(v^*)}] \leq \frac{n}{T}.$$

Thus, by Markov's inequality, we have that

$$\Pr_{h_{r^*}} \left[|\{v \in S : v \neq v^* \text{ and } h_{r^*}(v) = h_{r^*}(v^*)\}| \geq \frac{16n}{T} \right] \leq \frac{1}{16}.$$

As the hash functions are independent from each other, for $R \geq 48 \ln(\frac{d'}{\beta})$, by the Chernoff bound we get that with probability at least $1 - \beta/d'$ (over sampling h_1, \dots, h_R) there exists a subset $R_1^{v^*} \subseteq [R]$ of size $|R_1^{v^*}| \geq \frac{7}{8}R$ s.t. for every $r^* \in R_1^{v^*}$ it holds that

$$|\{v \in S : v \neq v^* \text{ and } h_{r^*}(v) = h_{r^*}(v^*)\}| \leq \frac{16n}{T}.$$

Using the union bound, we have that event E_1 happens with probability at least $1 - \beta$. We continue the analysis assuming that event E_1 occurs.

For every $r \in [R]$, let $S_r = (v_j)_{j \in I_r}$ denote a database containing the data of all users j s.t. $j \in I_r$. Also for $v^* \in V$ and $r \in [R]$ denote $|S^{r,v^*}| \triangleq |\{v \in S : h_r(v) = h_r(v^*)\}|$. That is, $|S^{r,v^*}|$ is the number of users j s.t. $h_r(v_j) = h_r(v^*)$. Furthermore, for $v^* \in V$ and $r \in [R]$ denote $|I_r^{v^*}| \triangleq |\{v \in S_r : h_r(v) = h_r(v^*)\}|$. That is, $|I_r^{v^*}|$ is the number of users j s.t. $j \in I_r$ and $h_r(v_j) = h_r(v^*)$. Observe that $|S^{r,v^*}| \geq f_S(v^*)$ and that $|I_r^{v^*}| \geq f_{S_r}(v^*)$.

Fix $v^* \in V$. By the Chernoff bound, with probability at least $1 - \beta/d'$ (over partitioning $[n]$ into subsets I_1, \dots, I_R), for every $r \in [R]$ we have that

$$\left| R \cdot |I_r^{v^*}| - |S^{r,v^*}| \right| \leq \sqrt{3R \cdot |S^{r,v^*}| \cdot \log\left(\frac{2Rd'}{\beta}\right)}. \quad (4)$$

Using the union bound this holds simultaneously for every $v^* \in V$ and $r \in [R]$ with probability at least $1 - \beta$. Also, assuming that $n \geq 8R \log(2R/\beta)$, by the Chernoff bound, with probability at least $1 - \beta$ (over partitioning $[n]$ into subsets I_1, \dots, I_R), for every $r \in [R]$ we have that $\frac{n}{2R} \leq |I_r| \leq \frac{2n}{R}$. We continue the analysis assuming that this is the case, and that Inequality (4) holds.

Event E_2 (over sampling Z and the coins of the local randomizers):

For every query $v^* \in V$ there exists a subset $R_2^{v^*} \subseteq [R]$ of size $|R_2^{v^*}| \geq \frac{7}{8}R$ s.t. for every $r^* \in R_2^{v^*}$ it holds that $\left| R \cdot a_{r^*}(h_{r^*}(v^*)) - R \cdot |I_{r^*}^{v^*}| \right| \leq \frac{e^\epsilon + 1}{e^\epsilon - 1} \cdot \sqrt{11nR}$.

For $v^* \in V$ and $r \in [R]$ denote $c_r(v^*) = \sum_{j \in I_r} y_j \cdot Z[h_r(v^*), j]$, and recall that algorithm **Hashtogram** answers the query v^* with $a(v^*) = R \cdot \frac{e^\epsilon + 1}{e^\epsilon - 1} \cdot \text{Median}_{r \in [R]} \{c_r(v^*)\}$. Fix $v^* \in V$ and $r \in [R]$. We now analyze the expectation of $c_r(v^*)$:

$$\begin{aligned} \mathbb{E}[c(v^*)] &= \sum_{j \in I_r} \mathbb{E}[y_j \cdot Z[h_r(v^*), j]] \\ &= \sum_{j \in I_r: h_r(v_j) = h_r(v^*)} \mathbb{E}[y_j \cdot Z[h_r(v^*), j]] + \sum_{j \in I_r: h_r(v_j) \neq h_r(v^*)} \mathbb{E}[y_j \cdot Z[h_r(v^*), j]] \\ &= \sum_{j \in I_r: h_r(v_j) = h_r(v^*)} \mathbb{E}[y_j \cdot Z[h_r(v^*), j]] + \sum_{j \in I_r: h_r(v_j) \neq h_r(v^*)} \mathbb{E}[y_j] \cdot \mathbb{E}[Z[h_r(v^*), j]] \\ &= |\{v \in S_r : h_r(v) = h_r(v^*)\}| \cdot \frac{e^\epsilon - 1}{e^\epsilon + 1} \triangleq |I_r^{v^*}| \cdot \frac{e^\epsilon - 1}{e^\epsilon + 1} \end{aligned}$$

That is, $c_r(v)$ can be expressed as two sums of ± 1 independent random variables: $|I_r^{v^*}|$ variables with expectation $\frac{e^\epsilon - 1}{e^\epsilon + 1}$, and $(|I_r| - |I_r^{v^*}|)$ variables with expectation 0 (recall that $\frac{n}{2R} \leq |I_r| \leq \frac{2n}{R}$). Using the Hoeffding bound, with probability at least $43/44$ we have that $\left| c_r(v^*) - \frac{e^\epsilon - 1}{e^\epsilon + 1} \cdot |I_r^{v^*}| \right| \leq \sqrt{11n/R}$. That is,

$$\left| R \cdot a_r(h_r(v^*)) - R \cdot |I_r^{v^*}| \right| \leq \frac{e^\epsilon + 1}{e^\epsilon - 1} \cdot \sqrt{11nR}. \quad (5)$$

Fix $v^* \in V$, and observe that the above sums are independent for different values of r . Hence, using the Chernoff bound and asserting that $R \geq 132 \ln(d'/\beta)$, for that fixed $v^* \in V$, with probability at least $1 - \beta/d'$ we have that Inequality (5) holds for at least $7R/8$ choices of $r \in [R]$. Using the union bound, with probability at least $1 - \beta$, this is true for every $v^* \in V$ simultaneously. That is, event E_2 happens with probability at least $1 - \beta$. We continue the analysis assuming that event E_2 occurs. For every $v^* \in V$ we denote $R_3^{v^*} = R_1^{v^*} \cap R_2^{v^*}$.

Combining event E_2 with Inequality (4), we get that for every $r \in R_2^{v^*}$

$$\left| R \cdot a_r(h_r(v^*)) - |S^{r,v^*}| \right| \leq \frac{e^\epsilon + 1}{e^\epsilon - 1} \cdot \sqrt{11nR} + \sqrt{3R \cdot |S^{r,v^*}| \cdot \log\left(\frac{2Rd'}{\beta}\right)}. \quad (6)$$

Recall that for every $v^* \in V$ and every $r \in [R]$ we have that $|S^{r,v^*}| \geq f_S(v^*)$. Furthermore, for every $v^* \in V$ and every $r \in R_1^{v^*}$ we have that $|S^{r,v^*}| \leq f_S(v^*) + \frac{16n}{T}$. Hence, for every $v^* \in V$ and every $r \in R_3^{v^*}$ we have that

$$\begin{aligned} |R \cdot a_r(h_r(v^*)) - f_S(v^*)| &\leq \frac{e^\epsilon + 1}{e^\epsilon - 1} \cdot \sqrt{11nR} + \sqrt{3R \cdot \left(f_S(v^*) + \frac{16n}{T}\right) \cdot \log\left(\frac{2Rd'}{\beta}\right) + \frac{16n}{T}} \\ &\triangleq \text{error}(v^*). \end{aligned}$$

That is, for every $r \in R_3^{v^*}$ we have that $R \cdot a_r(h_r(v^*))$ is accurate up to $\text{error}(v^*)$. As $|R_3^{v^*}| \geq \frac{3}{4}R$, and as algorithm **Hashtogram** answers v^* with $a(v^*)$ chosen as the median of $\{R \cdot a_r(h_r(v^*))\}$, we get that $|a(v^*) - f_S(v^*)| \leq \text{error}(v^*)$ for every $v^* \in V$.