

# Zero-One Laws for Sliding Windows and Universal Sketches

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## Abstract

Given a stream of data, a typical approach in streaming algorithms is to design a sophisticated algorithm with small memory that computes a specific statistic over the streaming data. Usually, if one wants to compute a different statistic after the stream is gone, it is impossible. But what if we want to compute a different statistic after the fact? In this paper, we consider the following fascinating possibility: can we collect some small amount of specific data during the stream that is “universal,” i.e., where we do not know anything about the statistics we will want to later compute, other than the guarantee that had we known the statistic ahead of time, it would have been possible to do so with small memory? This is indeed what we introduce (and show) in this paper with matching upper and lower bounds: we show that it is possible to collect universal statistics of polylogarithmic size, and prove that these universal statistics allow us after the fact to compute all other statistics that are computable with similar amounts of memory. We show that this is indeed possible, both for the standard unbounded streaming model and the sliding window streaming model.

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## 1 Introduction

With the vast amount of data being generated today, algorithms for data streams continue to play an important role for many practical applications. As the data being generated continues to grow at a staggering rate, streaming algorithms are increasingly becoming more important as a practical tool to analyze and make sense of all the information. Data streams have recently received a lot of attention with good reason, as evidenced by their wide array of applications. In particular, applications for streaming algorithms which operate over input that arrives on the fly and use a small amount of memory are numerous, ranging from monitoring packets flowing across a network to analyzing patterns in DNA sequences. In practice, such applications generate vast amounts of data in a very short period of time, so it is infeasible to store all this information. This presents a pressing question: when is it possible to avoid storing all the information while still providing approximate solutions with good theoretical guarantees?

Typically, algorithms are developed for data streams in the unbounded model, where some statistic is maintained over the entire history of the stream. For certain applications, it is useful to only compute such statistics over recent data. For instance, we may wish to analyze stock market transactions in a particular timeframe or monitor packets transmitted over a network in the last hour to identify suspicious activity. This framework is known as the sliding window model, where we maintain statistics over the current window of size at most  $N$ , which slides as time progresses. In the sequence-based model, exactly one element arrives and expires from the window per time step. In the timestamp-based model, any number of elements may arrive or expire. Clearly, the timestamp-based model is more general.

In a landmark paper that influenced the streaming field, the work of Alon, Matias and Szegegy [3] studied the following fundamental framework. For a universe  $U = \{1, \dots, n\}$  and an input stream (i.e., a sequence of integers drawn from  $U$ ), let  $M = (m_1, \dots, m_n)$  be the vector where  $m_i$  denotes the frequency that element  $i \in U$  appears in the stream. At any point in time, the paper of [3] showed how to approximate various frequency moments in sublinear space. Informally, for the  $k^{\text{th}}$  frequency moment  $F_k = \sum_{i \in U} m_i^k$ , it was shown that  $F_0, F_1$ , and  $F_2$  can be approximated in polylogarithmic space, while for  $k > 2$ , an upper bound of  $O^*(n^{1-1/k})$  was shown (the notation  $O^*(f(n))$  hides polylogarithmic factors). In addition, a lower bound of  $\Omega(n^{1-5/k})$  was shown for every  $k \geq 6$ . As discussed in [3], such frequency functions are very important in practice and have many applications in databases, as they indicate the degree to which the data is skewed. The fundamental work of Indyk and Woodruff [24] showed how to compute  $F_k$  for  $k > 2$  in space  $O^*(n^{1-2/k})$ , which was the first optimal result for such frequency moments. They reduced the problem of computing  $F_k$  to computing heavy hitters, and indeed our construction builds on their methods. Recently, Li, Nguyễn, and Woodruff [28] showed that any one-pass streaming algorithm that approximates an arbitrary function in the turnstile model can be implemented via linear sketches. Our work is related, as our algorithms are based on linear sketches of [3].

Such works have opened a line of research that is still extremely relevant today. In particular, what other types of frequency-based functions admit efficient solutions in the streaming setting, and which functions are inherently difficult to approximate? In our paper, we strive to answer this question for frequency-based, monotonically increasing functions in the sliding window model. We make progress on two significant, open problems outlined in [2] by Nelson and [1] by Sohler. Specifically, we are the first to formalize the notion of universality for streaming over sliding windows (since the sliding window model is more general than the standard unbounded model, our universality result is also the first such

contribution in the unbounded model). Our main result is the construction of a universal algorithm in the timestamp-based sliding window model for a broad class of functions. That is, we define a class of functions and design a single streaming algorithm that produces a data structure with the following guarantee. When querying the data structure with a function  $G$  taken from the class, our algorithm approximates  $\sum_{i=1}^n G(m_i)$  without knowing  $G$  in advance (here,  $m_i$  denotes the frequency that element  $i$  appears in the window). This is precisely the notion of universality that we develop in our paper, and it is an important step forward towards resolving the problem in [2].

Along the way, we design a zero-one law for a broader class of monotonically increasing functions  $G$  which are zero at the origin that specifies when  $\sum_{i=1}^n G(m_i)$  can be approximated with high probability in one pass, using polylogarithmic memory. If  $G$  satisfies the conditions specified by the test, then given the function  $G$  we construct an explicit, general algorithm that is able to approximate the summation to within a  $(1 \pm \epsilon)$ -factor using polylogarithmic memory. If the function  $G$  does not pass the test, then we provide a lower bound which proves it is impossible to do so. This result generalizes the work of [9] to the sliding window setting, and makes important progress towards understanding the question posed in [1].

## 1.1 Contributions and Techniques

Our contributions in this paper make progress on two important problems:

1. We are the first to formally define the notion of universality in the streaming setting. We define a large class of functions  $\mathcal{U}$  such that, for the entire class, we design a single, universal algorithm for data streams in the sliding window model which maintains a data structure with the following guarantee. When the data structure is queried with any function  $G \in \mathcal{U}$ , it outputs a  $(1 \pm \epsilon)$ -approximation of  $\sum_{i=1}^n G(m_i)$  without knowing  $G$  in advance (note that the choice of  $G$  can change). Our algorithm uses polylogarithmic memory, makes one pass over the stream, and succeeds with high probability.
2. We give a complete, algebraic characterization for the class of tractable functions over sliding windows. We define a broader set of functions  $\mathcal{T}$  such that, for any non-decreasing function  $G$  with  $G(0) = 0$ , if  $G \in \mathcal{T}$ , then we have an algorithm that gives a  $(1 \pm \epsilon)$ -approximation to  $\sum_{i=1}^n G(m_i)$ , uses polylogarithmic space, makes one pass over the stream, and succeeds with high probability. Moreover, if  $G \notin \mathcal{T}$ , we give a lower bound which shows that super-polylogarithmic memory is necessary to approximate  $\sum_{i=1}^n G(m_i)$  with high probability. This extends the work of [9] to the sliding window model.

Our algorithms work in the timestamp-based sliding window model and maintain the sum approximately for every window. The value  $\epsilon$  can depend on  $n$  and  $N$ , so that the approximation improves as either parameter increases. Our construction is very general, applying to many functions using the same techniques. In stark contrast, streaming algorithms typically depend specifically on the function to be approximated (e.g.,  $F_2$  [3, 22] and  $F_0$  [3, 13, 17]). The problems we study have been open for several years, and our construction and proofs are non-trivial. Surprisingly, despite us using existing techniques, their solutions have remained elusive.

For our main result, item 1, it is useful to understand our techniques for solving item 2. When designing the correct zero-one law for tractable functions, a natural place to begin is to understand whether the predicate from [9] suffices for designing an algorithm in the sliding window model. As it turns out, there are some functions which are tractable in the unbounded model but not the sliding window model, and hence the predicate is insufficient. Part of the novelty of our techniques is the identification of an extra smoothing assumption

for the class of tractable functions over sliding windows.

If a function does not satisfy our smoothing assumption, we show a super-polylogarithmic lower bound, inspired by the proof of [15]. For our positive result, we observe that the sliding window model adds extra error terms relative to the unbounded model, which our smoothness condition can bound. We also draw on the methods of [9, 10, 24] by finding heavy elements according to the function  $G$ , and then reducing the sum problem to the heavy elements problem. Our work sheds light on the question posed in [1], by exhibiting a strict separation result between the unbounded and sliding window models. A function which serves as a witness to this separation (i.e., tractable in the unbounded model as defined in [9] but not in the sliding window model) is a monotonically increasing, piecewise linear function that alternates between being constant and linearly increasing. The function can be seen as a linear approximation to  $\log(x)$ .

To obtain our main result, we observe that one can remove the assumption from our initial constructions that  $G$  is given up front (so that all applications of  $G$  happen at the end of the window). However, some technical issues arise, as our construction relies on some parameters of  $G$  that stem from our zero-one law. To address these issues, we parameterize our class of functions  $\mathcal{U}$  by a constant, allowing us to build a single algorithm to handle the entire parameterized class.

## 1.2 Related Work

The paper of Braverman and Ostrovsky [9] is the most closely related to our paper. We extend their result from the unbounded model to the timestamp-based sliding window model (by formalizing a new characterization of tractable functions) and by designing a universal algorithm for a large class of functions. Our results build on [9, 10, 24].

Approximating frequency moments and  $L_p$  norms has many applications, and there are indeed a vast number of papers on the subject. Compared to such works, we make minimal assumptions and our results are extremely broad, as we design general algorithms that can not only handle frequency moments, but other functions as well. Flajolet and Martin [17] gave an algorithm to approximate  $F_0$  (i.e., counting distinct elements), and Alon, Matias, and Szegedy [3] showed how to approximate  $F_k$  for  $0 \leq k \leq 2$  using polylogarithmic memory, while for  $k > 2$  they showed how to approximate  $F_k$  using  $O^*(n^{1-1/k})$  memory. They also showed an  $\Omega(n^{1-5/k})$  lower bound for  $k \geq 6$ . Indyk [22] used stable distributions to approximate  $L_p$  norms for  $p \in (0, 2]$ . Indyk and Woodruff [24] gave the first optimal algorithm for  $F_k$  ( $k > 2$ ), where an  $O^*(n^{1-2/k})$  upper bound was developed. In a followup work, Bhuvanagiri, Ganguly, Kesh, and Saha [6] improved the space by polylogarithmic factors. Bar-Yossef, Jayram, Kumar, and Sivakumar [4] gave an  $\Omega(n^{1-(2+\epsilon)/k})$  lower bound, which was improved to  $\Omega(n^{1-2/k})$  by Chakrabarti, Khot, and Sun [11] for any one-pass streaming algorithm. The literature is vast, and other results for such functions include [5, 12, 13, 16, 18, 23, 26, 27, 31].

There is also a vast literature in streaming for sliding windows. In their foundational paper, Datar, Gionis, Indyk, and Motwani [15] gave a general technique called exponential histograms that allows many fundamental statistics to be computed in optimal space, including count, sum of positive integers, average, and the  $L_p$  norm for  $p \in [1, 2]$ . Gibbons and Tirthapura [19] made improvements for the sum and count problem with algorithms that are optimal in space and time. Braverman and Ostrovsky [8] gave a general framework for a large class of smooth functions, which include the  $L_p$  norm for  $p > 0$ . Our work complements their results, as the functions they studied need not be frequency based. Many works have studied frequency estimation and frequent item identification, including [7, 14, 20, 21, 25, 30, 32]. Many of our constructions rely on computing frequent elements, but we must do so under a broad

class of functions.

### 1.3 Roadmap

In Section 2, we describe notation used throughout this paper, give some definitions, and formalize the main problems we study. In Section 3, we give a lower bound for functions that are not tractable (i.e., we show the “zero” part of our zero-one law) and we give an algorithm for any tractable function (i.e., we show the “one” part of our zero-one law). Finally, in Section 4, we show the main result of this paper by giving a universal streaming algorithm.

## 2 Notation and Problem Definition

We have a universe of  $n$  elements  $[n]$ , where  $[n] = \{1, \dots, n\}$ , and an integer  $N$ . A stream  $D(n, N)$  is a (possibly infinite) sequence of integers  $a_1, a_2, \dots$ , each from the universe  $[n]$ , where  $N$  is an upper bound on the size of the sliding window. Specifically, at each time step, there is a current window  $W$  that contains *active* elements, where  $|W| \leq N$ . The window  $W$  contains the most recent elements of the stream, and elements which no longer belong in the window are *expired*. We use the timestamp-based model for sliding windows (i.e., any number of elements from the stream may enter or leave the window at each time step). We denote the frequency vector by  $M(W)$ , where  $M(W) = (m_1, \dots, m_n)$  and each  $m_i$  is the frequency of element  $i \in [n]$  in window  $W$  (i.e.,  $m_i = |\{j \mid a_j = i \wedge j \text{ is active}\}|$ ). For the window  $W$ , the  $k^{\text{th}}$  frequency moment  $F_k(M(W)) = \sum_{i=1}^n m_i^k$ . For a vector  $V = (v_1, \dots, v_n)$ , we let  $|V|$  be the  $L_1$  norm of  $V$ , namely  $|V| = \sum_i |v_i|$ . For a vector  $V = (v_1, \dots, v_n)$  and a function  $f$ , we define the  $f$ -Vector as  $f(V) = (f(v_1), \dots, f(v_n))$ . We say that  $x$  is a  $(1 \pm \epsilon)$ -approximation of  $y$  if  $(1 - \epsilon)y \leq x \leq (1 + \epsilon)y$ . We define  $O^*(f(n, N)) = O(\log^{O(1)}(nN)f(n, N))$ . We say a probability  $p$  is negligible if  $p = O^*(\frac{1}{nN})$ . Consider the following problem:

► **Problem 1 ( $G$ -Sum).** Let  $G : \mathbb{R} \rightarrow \mathbb{R}$  be an arbitrary non-decreasing function such that  $G(0) = 0$ . For any stream  $D(n, N)$ , any  $k$ , and any  $\epsilon = \Omega(1/\log^k(nN))$ , output a  $(1 \pm \epsilon)$ -approximation of  $\sum_{i=1}^n G(m_i)$  for the current window  $W$ .

We first give some definitions which will be useful throughout the paper and help us define our notion of tractability, beginning with the local jump:

► **Definition 1 (Local Jump).**  $\forall \epsilon > 0, x \in \mathbb{N}$ , we define the local jump  $\pi_\epsilon(x)$  as

$$\min \{x, \min \{z \in \mathbb{N} \mid G(x+z) > (1+\epsilon)G(x) \vee G(x-z) < (1-\epsilon)G(x)\}\}.$$

That is,  $\pi_\epsilon(x)$  is essentially the minimum amount needed to cause  $G$  to jump by a  $(1 \pm \epsilon)$ -factor by shifting either to the left or right of  $x$ .

► **Definition 2 (Heavy Element).** For a vector  $V = (v_1, \dots, v_n)$ , parameter  $d > 0$ , and function  $f$ , we say an element  $i$  is  $(f, d)$ -heavy with respect to the vector  $V$  if  $f(v_i) > d \sum_{j \neq i} f(v_j)$ .

► **Definition 3 (Residual Second Moment).** If there is an  $(F_2, 1)$ -heavy element  $v_i$  with respect to  $V = (v_1, \dots, v_n)$ , we define the *residual second moment* as  $F_2^{\text{res}}(V) = F_2(V) - v_i^2 = \sum_{j \neq i} v_j^2$ .

► **Definition 4 (Sampled Substream).** For a stream  $D(n, N)$  and function  $H : [n] \rightarrow \{0, 1\}$ , we denote by  $D_H$  the *sampled substream* of  $D$  consisting of all elements that are mapped to 1 by the function  $H$ . More formally,  $D_H = D \cap H^{-1}(1)$ .

We analogously define  $W_H$  to be the corresponding window for the sampled substream  $D_H$ . We are now ready to define our zero-one law.

► **Definition 5 (Tractability).** We say a function  $G$  is tractable if  $G(1) > 0$  and:

$$\forall k \exists N_0, t \forall x, y \in \mathbb{N}^+ \forall R \in \mathbb{R}^+ \forall \epsilon : \quad \left( R > N_0, \frac{G(x)}{G(y)} = R, \epsilon > \frac{1}{\log^k(Rx)} \right) \Rightarrow \left( \left( \frac{\pi_\epsilon(x)}{y} \right)^2 \geq \frac{R}{\log^t(Rx)} \right) \quad \text{and} \quad (1)$$

$$\forall k \exists N_1, r \forall x \geq N_1 \forall \epsilon : \epsilon > \frac{1}{\log^k(x)} \Rightarrow \pi_\epsilon(x) \geq \frac{x}{\log^r(x)}. \quad (2)$$

We let **Tractable** be the set of functions that satisfy the above predicate. We now turn to our universal setting and develop an analogous notion of tractability in the context of universality. It is similar to Definition 5, except we need to upper bound some parameters by a constant.

► **Definition 6 (Universal Tractability).** Fix a constant  $C$ . Let  $\mathcal{U}(C)$  denote the set of non-decreasing functions  $G$  where  $G(0) = 0$ ,  $G(1) > 0$ , and:

$$\forall k \leq C \exists N_0, t \leq 10C \forall x, y \in \mathbb{N}^+ \forall R \in \mathbb{R}^+ \forall \epsilon : \quad \left( R > N_0, \frac{G(x)}{G(y)} = R, \epsilon > \frac{1}{\log^k(Rx)} \right) \Rightarrow \left( \left( \frac{\pi_\epsilon(x)}{y} \right)^2 \geq \frac{R}{\log^t(Rx)} \right) \quad \text{and} \quad (3)$$

$$\forall k \leq C \exists N_1, r \leq 10C \forall x \geq N_1 \forall \epsilon : \epsilon > \frac{1}{\log^k(x)} \Rightarrow \pi_\epsilon(x) \geq \frac{x}{\log^r(x)}. \quad (4)$$

Some examples of functions that are tractable in the universal sense include the moments  $x^p$  for  $p \leq 2$ , for which  $\pi_\epsilon(x) = \Omega(\epsilon x)$ , along with other functions such as  $(x+1) \log(x+1)$ .

► **Definition 7 (Universal Core Structure).** A data structure  $S$  is a universal core structure for a fixed vector  $V = (v_1, \dots, v_n)$  with parameters  $\epsilon, \delta, \alpha > 0$ , and a class of functions  $\mathcal{G}$ , where  $G \in \mathcal{G}$  satisfies  $G : \mathbb{R} \rightarrow \mathbb{R}$ , if given any  $G \in \mathcal{G}$ ,  $S$  outputs a set  $T = \{(x_1, j_1), \dots, (x_\ell, j_\ell)\}$  such that with probability at least  $1 - \delta$  we have: 1) For each  $1 \leq i \leq \ell$ ,  $(1 - \epsilon)G(v_{j_i}) \leq x_i \leq (1 + \epsilon)G(v_{j_i})$ , and 2) If there exists  $i$  such that  $v_i$  is  $(G, \alpha)$ -heavy with respect to  $V$ , then  $i \in \{j_1, \dots, j_\ell\}$ .

► **Definition 8 (Universal Core Algorithm).** An algorithm  $\mathcal{A}$  is a universal core algorithm with parameters  $\epsilon, \delta, \alpha > 0$ , and a class of functions  $\mathcal{G}$ , where  $G \in \mathcal{G}$  satisfies  $G : \mathbb{R} \rightarrow \mathbb{R}$ , if, given any stream  $D(n, N)$ ,  $\mathcal{A}$  outputs a universal core structure for the vector  $M(W)$  with parameters  $\epsilon, \delta, \alpha$ , and  $\mathcal{G}$ .

► **Definition 9 (Universal Sum Structure).** A data structure  $S$  is a universal sum structure for a fixed vector  $V = (v_1, \dots, v_n)$  with parameters  $\epsilon, \delta > 0$ , and a class of functions  $\mathcal{G}$ , where  $G \in \mathcal{G}$  satisfies  $G : \mathbb{R} \rightarrow \mathbb{R}$ , if given any  $G \in \mathcal{G}$ ,  $S$  outputs a value  $x$  such that with probability at least  $1 - \delta$  we have:  $(1 - \epsilon) \sum_{i=1}^n G(v_i) \leq x \leq (1 + \epsilon) \sum_{i=1}^n G(v_i)$ .

► **Definition 10 (Universal Sum Algorithm).** An algorithm  $\mathcal{A}$  is a universal sum algorithm with parameters  $\epsilon, \delta > 0$ , and a class of functions  $\mathcal{G}$ , where  $G \in \mathcal{G}$  satisfies  $G : \mathbb{R} \rightarrow \mathbb{R}$ , if, given any stream  $D(n, N)$ ,  $\mathcal{A}$  outputs a universal sum structure for the vector  $M(W)$  with parameters  $\epsilon, \delta$ , and  $\mathcal{G}$ .

In this paper, our main result is the proof of the following theorem:

► **Theorem 1.** Fix a constant  $C$  and let  $\mathcal{U}(C)$  be the universally tractable set from Definition 6. There is a universal sum algorithm that has parameters  $\epsilon = \Omega(1/\log^k(nN))$  (for  $0 \leq k \leq C$ ),  $\delta = 0.3$ , and  $\mathcal{G} = \mathcal{U}(C)$ , uses polylogarithmic space in  $n$  and  $N$ , and makes a single pass over the input stream  $D(n, N)$ .

We can reduce the constant failure probability to inverse polynomial via standard methods. To formalize our other main result, we define the following class:

► **Definition 11 (STREAM-POLYLOG).** We say function  $G \in \text{STREAM-POLYLOG}$  if  $\forall k = O(1)$ ,  $\exists t = O(1)$  and an algorithm  $\mathcal{A}$  such that for any universe size  $n$ , window size  $N$ ,  $\epsilon \geq 1/\log^k(nN)$ , and stream  $D(n, N)$ : 1)  $\mathcal{A}$  makes one pass over  $D$ , 2)  $\mathcal{A}$  uses  $O(\log^t(nN))$  space, and 3) For any window  $W$ ,  $\mathcal{A}$  maintains a  $(1 \pm \epsilon)$ -approximation of  $|G(M(W))|$  except with probability at most 0.3.

Note that the constant error probability can be made to be as small as an inverse polynomial by standard techniques. Our other main result is the following:

► **Theorem 2.** Let  $G$  be a non-decreasing function such that  $G(0) = 0$ . Then we have  $G \in \text{STREAM-POLYLOG} \iff G \in \text{Tractable}$ .

### 3 A Characterization for Tractable Functions

In this section, we prove Theorem 2 by first giving a lower bound for non-tractable functions. We first show a deterministic lower bound for any algorithm that approximates  $G$ -Sum. Our technique is inspired by the lower bound proof in [15] for estimating the number of 1's for sliding windows.

► **Theorem 3.** Let  $G$  be a function such that  $G \notin \text{Tractable}$ . Then, any deterministic algorithm that solves the  $G$ -Sum problem with relative error  $\epsilon' = 1/\log^b(nN)$  (for some constant  $b$ ) must use space at least  $\Omega(\log^a(nN))$ , where  $a$  is arbitrarily large.

**Proof.** We construct a set of input streams such that, for any pair of data streams in the set, the algorithm must distinguish between these two inputs at some point as the window slides. Therefore, the space of the algorithm must be at least logarithmic in the size of this set.

Since  $G \notin \text{Tractable}$ , in Definition 5, either Predicate (1) or Predicate (2) does not hold. If Predicate (1) is not true, then the lower bound from [9] applies and the theorem follows. Hence, we assume that Predicate (2) does not hold, which implies the following:  $\exists k, \forall r, N_1, \exists x \geq N_1, \epsilon : \epsilon > \frac{1}{\log^k(x)} \wedge \pi_\epsilon(x) < \frac{x}{\log^r(x)}$ . Let  $k$  be given, and let  $r$  be arbitrarily large. This negation implies that there are infinitely many increasing points  $x_1, x_2, x_3, \dots$  and corresponding values  $\epsilon_1, \epsilon_2, \epsilon_3, \dots$ , where  $\epsilon_i > \frac{1}{\log^k(x_i)}$  and  $\pi_{\epsilon_i}(x_i) < \frac{x_i}{\log^r(x_i)}$ .

Surprisingly, we construct our lower bound with a universe of size  $n = 1$ , namely  $U = \{1\}$ . For each  $x_i$ , we construct a set of streams with a fixed, upper bounded window size of  $N = x_i$ , and argue that the algorithm must use memory at least  $\log^r(x_i)$  (note that, as the  $x_i$  are monotonically increasing, our lower bound applies for asymptotically large  $N$ ). We assume without loss of generality that  $G(x_i - \pi_{\epsilon_i}(x_i)) < (1 - \epsilon_i)G(x_i)$ . Our constructed streams are defined as follows. For each  $N = x_i$ , note that our window consists of elements which have arrived in the past  $x_i$  time steps. For the first  $x_i$  time steps, we construct many streams by choosing  $\lfloor \frac{x_i}{\pi_{\epsilon_i}(x_i)} \rfloor$  of these time steps (each choice defining a different stream). For each chosen time step, we insert  $\pi_{\epsilon_i}(x_i)$  1's into the stream, and for each time step that is not chosen, we insert zero elements. For technical reasons, we pad the last time step  $x_i$  in the first window with  $x_i - \pi_{\epsilon_i}(x_i) \lfloor \frac{x_i}{\pi_{\epsilon_i}(x_i)} \rfloor$  1's. Note that the number of elements in the first window at time  $x_i$  is  $\pi_{\epsilon_i}(x_i) \lfloor \frac{x_i}{\pi_{\epsilon_i}(x_i)} \rfloor + (x_i - \pi_{\epsilon_i}(x_i) \lfloor \frac{x_i}{\pi_{\epsilon_i}(x_i)} \rfloor) = x_i$ . We insert nothing at time step  $x_i + 1$ . For the remaining time steps  $x_i + 2, \dots, 2x_i - 1$ , we simply repeat the first  $x_i - 2$  time steps of the stream (i.e., if time step  $t$  was chosen in the first  $x_i$  time steps,  $1 \leq t \leq x_i - 2$ , then we insert  $\pi_{\epsilon_i}(x_i)$  1's at time step  $x_i + t + 1$ ).

Now, we argue that for any such pair of constructed streams  $A, B$  which are different, any algorithm with relative error smaller than  $\epsilon' = 1/\log^k(nN)$  must distinguish between



these two inputs. To see this, consider the earliest time  $d$  when the two streams differ (note that  $1 \leq d \leq x_i - 1$ ). Let  $W_A$  be the window for stream  $A$  (similarly, we define  $W_B$  as the window for stream  $B$ ). Let  $c$  be the number of chosen time steps in the first  $d$  time steps of stream  $A$ . Without loss of generality, we assume that time step  $d$  was chosen in stream  $A$  but not in stream  $B$ . Hence, the number of chosen time steps in stream  $B$  up to time  $d$  is  $c - 1$ . Consider the windows at time step  $x_i + d$ . The number of elements in  $W_A$  at this time is given by  $\pi_{\epsilon_i}(x_i) \lfloor \frac{x_i}{\pi_{\epsilon_i}(x_i)} \rfloor - c + (c - 1) + x_i - \pi_{\epsilon_i}(x_i) \lfloor \frac{x_i}{\pi_{\epsilon_i}(x_i)} \rfloor = x_i - \pi_{\epsilon_i}(x_i)$ . Moreover, the number of elements in  $W_B$  is given by  $\pi_{\epsilon_i}(x_i) \lfloor \frac{x_i}{\pi_{\epsilon_i}(x_i)} \rfloor - (c - 1) + (c - 1) + x_i - \pi_{\epsilon_i}(x_i) \lfloor \frac{x_i}{\pi_{\epsilon_i}(x_i)} \rfloor = x_i$ . Hence, the  $G$ -Sum value at time  $x_i + d$  for  $W_A$  is  $G(x - \pi_{\epsilon_i}(x_i)) < (1 - \epsilon_i)G(x_i)$ . As long as the algorithm has relative error  $\epsilon' = 1/\log^k(nN) < \epsilon_i$ , streams  $A$  and  $B$  must be distinguished at some point in time as the window slides.

Thus, the algorithm's memory is lower bounded by the logarithm of the number of constructed streams, of which there are  $\lfloor \frac{x_i}{\pi_{\epsilon_i}(x_i)} \rfloor$  for each  $x_i$ . We have  $\log \left( \left\lfloor \frac{x_i}{\pi_{\epsilon_i}(x_i)} \right\rfloor \right) \geq \lfloor \frac{x_i}{\pi_{\epsilon_i}(x_i)} \rfloor \log(\pi_{\epsilon_i}(x_i)) \geq \frac{\log^r(x_i)}{2} \log(\pi_{\epsilon_i}(x_i))$ . If  $\pi_{\epsilon_i}(x_i) = 1$ , we repeat the proof inserting two 1's at each time step and the proof goes through. Observing that  $r$  can be made arbitrarily large gives the proof.  $\blacktriangleleft$

We now have a randomized lower bound by appealing to Yao's minimax principle [29] and building on top of our deterministic lower bound, similarly to [15] (applying the principle with the uniform distribution suffices).

► **Theorem 4.** Let  $G$  be a function where  $G \notin \text{Tractable}$ . Then, any randomized algorithm that solves  $G$ -Sum with relative error smaller than  $\epsilon' = 1/\log^b(nN)$  for some constant  $b$  and succeeds with at least constant probability  $1 - \delta$  must use memory  $\Omega(\log^a(nN))$ , where  $a$  is arbitrarily large.

We now complete the proof of Theorem 2 by first approximating heavy elements (note that we reduce the  $G$ -Sum problem to the following problem):

► **Problem 2 ( $G$ -Core).** We have a stream  $D(n, N)$  and parameters  $\epsilon, \delta > 0$ . For each window  $W$ , with probability at least  $1 - \delta$ , maintain a set  $S = \{g'_1, \dots, g'_\ell\}$  such that  $\ell = O^*(1)$  and there exists a set of indices  $\{j_1, \dots, j_\ell\}$  where  $(1 - \epsilon)G(m_{j_i}) \leq g'_i \leq (1 + \epsilon)G(m_{j_i})$  for each  $1 \leq i \leq \ell$ . If there is a  $(G, 1)$ -heavy element  $m_i$  with respect to  $M(W)$ , then  $i \in \{j_1, \dots, j_\ell\}$ .

We begin solving the above problem via the following lemma (taken from [9]).

► **Lemma 1.** Let  $V = (v_1, \dots, v_n)$  be a vector with non-negative entries of dimension  $n$  and  $H$  be a pairwise independent random vector of dimension  $n$  with entries  $h_i \in \{0, 1\}$  such that  $P(h_i = 1) = P(h_i = 0) = \frac{1}{2}$ . Denote by  $H'$  the vector with entries  $1 - h_i$ . Let  $K > 10^4$  be a constant, and let  $X = \langle V, H \rangle$  and  $Y = \langle V, H' \rangle$ . If there is an  $(F_1, K)$ -heavy element  $v_i$  with respect to  $V$ , then:  $P((X > KY) \vee (Y > KX)) = 1$ . If there is no  $(F_1, \frac{K}{10^4})$ -heavy element with respect to  $V$ , then:  $P((X > KY) \vee (Y > KX)) \leq \frac{1}{2}$ .

We now give some lemmas related to how approximating values can affect the function  $G$ .

► **Lemma 2.** Let  $0 < \epsilon \leq \frac{1}{2}$ , and let  $x, u, v, y \geq 0$  satisfy  $|x - u| \leq 0.1\pi_\epsilon(x)$  and  $v, y < 0.1\pi_\epsilon(x)$ , where  $\pi_\epsilon(x) > 1$ . Then  $(1 - 4\epsilon)G(u + v + y) \leq G(u) \leq (1 + 4\epsilon)G(u - v - y)$ .

**Proof.** First, we note that  $u + v + y \leq x + 0.1\pi_\epsilon(x) + v + y \leq x + 0.3\pi_\epsilon(x) \leq x + \pi_\epsilon(x) - 1$  (recalling  $\pi_\epsilon(x) > 1$ ). We can similarly get that  $u - v - y \geq x - (\pi_\epsilon(x) - 1)$ . Hence, we get that  $(1 - \epsilon)G(x) \leq G(x - (\pi_\epsilon(x) - 1)) \leq G(u - v - y) \leq G(u) \leq G(u + v + y) \leq G(x + (\pi_\epsilon(x) - 1)) \leq (1 + \epsilon)G(x)$ .

We conclude by noting:  $(1 + 4\epsilon)G(u - v - y) \geq (1 + 4\epsilon)(1 - \epsilon)G(x) \geq \frac{(1 + 4\epsilon)(1 - \epsilon)}{1 + \epsilon}G(u) \geq G(u)$ . Similarly, we get  $(1 - 4\epsilon)G(u + v + y) \leq (1 - 4\epsilon)(1 + \epsilon)G(x) \leq \frac{(1 - 4\epsilon)(1 + \epsilon)}{1 - \epsilon}G(u) \leq G(u)$ .  $\blacktriangleleft$



► **Lemma 3.** Let  $x, u, v, y \geq 0$  be such that  $|x - u| \leq v + y$ , and let  $0 < \epsilon < 1$ . If  $(1 - \epsilon)G(u + v + y) \leq G(u) \leq (1 + \epsilon)G(u - v - y)$ , then  $(1 - \epsilon)G(x) \leq G(u) \leq (1 + \epsilon)G(x)$ .

**Proof.** We have  $(1 - \epsilon)G(x) \leq (1 - \epsilon)G(u + v + y) \leq G(u) \leq (1 + \epsilon)G(u - v - y) \leq (1 + \epsilon)G(x)$ . ◀

We now give a useful subroutine over sliding windows which we use in our main algorithm for approximating heavy elements and prove its correctness (there is a similar algorithm and proof in [9], though it must be adapted to the sliding window setting).

```

1 for  $i = 1$  to  $O(\log(nN))$  do
2   for  $j = 1$  to  $C = O(1)$  do
3     Generate a random hash function  $H : [n] \rightarrow \{0, 1\}$  with pairwise independent
     entries.
4     Let  $H' = 1 - H$  (i.e.,  $h'_k = 1 - h_k$ , where  $h_k$  is the  $k^{\text{th}}$  entry of  $H$ ).
5     Let  $f_H$  be a  $(1 \pm .1)$ -approximation of  $F_2$  on  $D_H$  (with negligible error
     probability), via the smooth histogram method for sliding windows [8].
6     Let  $f_{H'}$  be a  $(1 \pm .1)$ -approximation of  $F_2$  on  $D_{H'}$  (with negligible error
     probability), via the smooth histogram method for sliding windows [8].
7     Let  $X_{ij} = 10 \min\{f_H, f_{H'}\}$ .
8     Let  $Y_i = \frac{X_{i1} + \dots + X_{iC}}{C}$  (i.e.,  $Y_i$  is the average of  $C$  independent  $X_{ij}$ 's).
9 Output  $r = \sqrt{\text{median}_i\{Y_i\}}$  for the current window  $W$ .
```

**Algorithm 1:** Residual-Approximation( $D$ )

► **Lemma 4.** Let  $D(n, N)$  be any input stream. Algorithm Residual-Approximation makes a single pass over  $D$  and uses polylogarithmic space in  $n$  and  $N$ . Moreover, if the current window  $W$  contains an  $(F_2, 2)$ -heavy element  $m_i$  with respect to  $M(W)$ , then the algorithm maintains and outputs a value  $r$  such that  $2\sqrt{F_2^{\text{res}}(M(W))} < r < 3\sqrt{F_2^{\text{res}}(M(W))}$  (except with negligible probability).

**Proof.** Assume the current window  $W$  has an  $(F_2, 2)$ -heavy element  $m_k$  with respect to the vector  $M(W)$ . Due to the properties of smooth histograms from [8], we know that  $.9F_2(M(W_H)) \leq f_H \leq 1.1F_2(M(W_H))$ , where  $M(W_H)$  is the multiplicity vector of the current window in substream  $D_H$  (similarly for  $f_{H'}$ ). Hence, the random variable  $X_{ij} = 10 \min\{f_H, f_{H'}\}$  is a  $(1 \pm .1)$ -approximation of the random variable  $Z = 10 \sum_{\ell} \mathbf{1}_{H(\ell) \neq H(k)} m_{\ell}^2$  (here,  $\mathbf{1}_{H(\ell) \neq H(k)}$  is the indicator random variable which is 1 if  $H(\ell) \neq H(k)$  and 0 otherwise). To see why, suppose that element  $k$  is mapped to 1 by  $H$ , so that  $k$  belongs to the sampled substream  $D_H$ . Then observe that

$$f_H \geq .9F_2(M(W_H)) \geq .9m_k^2 > 1.8 \sum_{\ell \neq k} m_{\ell}^2 \geq 1.1 \sum_{\ell} \mathbf{1}_{H(\ell) \neq H(k)} m_{\ell}^2 \geq f_{H'}.$$

Thus, the minimum of  $f_H$  and  $f_{H'}$  is indeed a  $(1 \pm .1)$ -approximation to  $\sum_{\ell} \mathbf{1}_{H(\ell) \neq H(k)} m_{\ell}^2$ , since this is the second moment of the vector  $M(W_{H'})$  (the case is symmetric if element  $k$  is mapped to 0 by  $H$ ).

Now, since  $H$  is pairwise independent, we have that  $E(Z) = 5F_2^{\text{res}}(M(W))$ . In particular, since we always have  $0 \leq Z \leq 10F_2^{\text{res}}(M(W))$ , we can bound the variance by  $\text{Var}(Z) \leq E(Z^2) \leq 100(F_2^{\text{res}}(M(W)))^2$ . If we denote by  $A$  the random variable which is the average of

$C$  independent  $Z$ 's, then we have  $\text{Var}(A) = \frac{1}{C} \text{Var}(Z) \leq \frac{100}{C} (F_2^{\text{res}}(M(W)))^2$ . Hence, if we choose  $C$  to be sufficiently large, then by Chebyshev's inequality we have:

$$P(|A - 5F_2^{\text{res}}(M(W))| \geq 0.1F_2^{\text{res}}(M(W))) \leq \frac{100\text{Var}(A)}{(F_2^{\text{res}}(M(W)))^2} \leq \frac{10^4}{C} \leq 0.1$$

(for instance,  $C = 10^5$  is sufficient).

Now, if we take the median  $T$  of  $O(\log(nN))$  independent  $A$ 's, then by Chernoff bound this would make the probability negligible. That is, we have  $4.9F_2^{\text{res}}(M(W)) \leq T \leq 5.1F_2^{\text{res}}(M(W))$  except with negligible probability. We can repeat these arguments and consider the median of  $O(\log(nN))$  averages (i.e., the  $Y_i$ 's) of  $O(1)$  independent  $X_{ij}$ 's. Since there are only  $O(\log(nN))$   $X_{ij}$ 's total (with each one being a  $(1 \pm .1)$ -approximation to its corresponding random variable  $Z$ , except with negligible probability), then by the union bound all the  $X_{ij}$ 's are  $(1 \pm .1)$ -approximations except with negligible probability (since the sum of polylogarithmically many negligible probabilities is still negligible). Therefore, the median of averages would give a  $(1 \pm .1)$ -approximation to  $T$ . Taking the square root guarantees that  $2\sqrt{F_2^{\text{res}}(M(W))} < r < 3\sqrt{F_2^{\text{res}}(M(W))}$  (except with negligible probability).

Note that the subroutine for computing an approximation to  $F_2$  on sliding windows using smooth histograms can be done in one pass and in polylogarithmic space (even if we demand a  $(1 \pm .1)$ -approximation and a negligible probability of failure). ◀

Now, we claim that Algorithm Compute-Hybrid-Major solves the following:

► **Problem 3 (Hybrid-Major( $D, \epsilon$ )).** Given a stream  $D$  and  $\epsilon > 0$ , maintain a value  $r \geq 0$  for each window  $W$  such that: 1) If  $r \neq 0$ , then  $r$  is a  $(1 \pm 4\epsilon)$ -approximation of  $G(m_j)$  for some  $m_j$ , and 2) If the current window  $W$  has an element  $m_i$  such that  $\pi_\epsilon(m_i) \geq 20^5 \sqrt{F_2^{\text{res}}(M(W))}$ , then  $r$  is a  $(1 \pm 4\epsilon)$ -approximation of  $G(m_i)$ .

- 1 Let  $a$  be a  $(1 \pm \epsilon')$ -approximation of  $L_2$  for window  $W$  using the smooth histogram method [8] (with negligible probability of error), for  $\epsilon' = \frac{1}{\log^{\Omega(1)}(N)}$ .
- 2 Repeat  $O(\log(nN))$  times, independently and in parallel:
  - 3 Generate a uniform pairwise independent vector  $H \in \{0, 1\}^n$ .
  - 4 Maintain and denote by  $X'$  a  $(1 \pm .2)$ -approximation of the second moment for the window  $W_H$  using a smooth histogram [8] (with negligible probability of error).
  - 5 Similarly define  $Y'$  for the window  $W_{1-H}$ .
  - 6 If  $X' < (20)^4 Y'$  and  $Y' < (20)^4 X'$ , output 0 and terminate the algorithm.
- 7 In parallel, apply Residual-Approximation( $D$ ) to maintain the residual second moment approximation, let  $b$  denote the output of the algorithm.
- 8 If  $(1 - 4\epsilon)G(a + b + 2\epsilon'a) > G(a)$  or  $G(a) > (1 + 4\epsilon)G(a - b - 2\epsilon'a)$ , output 0.
- 9 Otherwise, output  $G(a)$ .

**Algorithm 2:** Compute-Hybrid-Major( $D, \epsilon$ )

Before delving into the proof, we show the following lemma.

► **Lemma 5.** Suppose the current window  $W$  has an  $(F_2, 1)$ -heavy element  $m_i$ . Moreover, let  $a$  be a  $(1 \pm \epsilon')$ -approximation of the  $L_2$  norm of the current window  $W$ , where  $\epsilon' < 1$ . Then  $-\epsilon' m_i \leq a - m_i \leq (1 + \epsilon') \sqrt{F_2^{\text{res}}(M(W))} + \epsilon' m_i \leq 2\sqrt{F_2^{\text{res}}(M(W))} + \epsilon' m_i$ .

**Proof.** Since  $a$  is a  $(1 \pm \epsilon')$ -approximation of the  $L_2$  norm of the vector  $M(W)$ , we know

$(1 - \epsilon')\sqrt{\sum_{k=1}^n m_k^2} \leq a \leq (1 + \epsilon')\sqrt{\sum_{k=1}^n m_k^2}$ . Hence, we have that

$$\begin{aligned} a - m_i &\leq (1 + \epsilon')\sqrt{\sum_{k=1}^n m_k^2} - m_i \leq (1 + \epsilon')m_i + (1 + \epsilon')\sqrt{\sum_{j \neq i} m_j^2} - m_i \\ &\leq \epsilon' m_i + (1 + \epsilon')\sqrt{F_2^{res}(M(W))}. \end{aligned}$$

We conclude by noting that  $m_i - a \leq m_i - (1 - \epsilon')\sqrt{\sum_{k=1}^n m_k^2} \leq m_i - (1 - \epsilon')m_i$ , which gives the lemma.  $\blacktriangleleft$

$\blacktriangleright$  **Lemma 6.** For any function  $G \in \text{Tractable}$ , Algorithm Compute-Hybrid-Major solves the Hybrid-Major problem with negligible probability of error.

**Proof.** First, we show that if there is no  $(F_2, 2)$ -heavy entry in the current window  $W$ , then the output is 0 except with negligible probability. Consider a single iteration of the main loop of the algorithm. Let  $M'$  be the vector with entries  $m_i^2$  and denote  $X = \langle M', H \rangle$ ,  $Y = |M'| - \langle M', H \rangle$ . Since we have an  $F_2$  approximation over sliding windows, except with negligible probability,  $X'$  and  $Y'$  are  $(1 \pm .2)$ -approximations of  $X$  and  $Y$ , respectively. Hence,  $\frac{4}{5}X \leq X' \leq \frac{5}{4}X$  and  $\frac{4}{5}Y \leq Y' \leq \frac{5}{4}Y$ . By Lemma 1, except with probability at most  $0.5 + o(1)$ :  $X' \leq \frac{5}{4}X \leq \frac{5}{2}(10)^4 Y' < (20)^4 Y'$  and  $Y' < (20)^4 X'$ . Thus, the algorithm outputs 0 except with negligible probability.

Assume that there is an  $(F_2, 2)$ -heavy entry  $m_i$ . Then, applying Lemma 5 with some  $0 < \epsilon' < 1$  to be set later, we know  $|m_i - a| \leq 2\sqrt{F_2(M(W))} + \epsilon' m_i$  and  $a \geq (1 - \epsilon')m_i$  (except with negligible probability). By Lemma 4, it follows that  $2\sqrt{F_2^{res}(M(W))} < b < 3\sqrt{F_2^{res}(M(W))}$  except with negligible probability. Hence, we have  $|m_i - a| \leq b + \epsilon' m_i \leq b + 2\epsilon' a$ , since  $2\epsilon' a \geq 2\epsilon'(1 - \epsilon')m_i \geq \epsilon' m_i$  (assuming  $\epsilon' \leq \frac{1}{2}$ ). Now, observe that if the algorithm outputs  $G(a)$ , then it must be that  $(1 - 4\epsilon)G(a + b + 2\epsilon' a) \leq G(a) \leq (1 + 4\epsilon)G(a - b - 2\epsilon' a)$ . Thus, by applying Lemma 3 with parameters  $x = m_i, u = a, v = b$ , and  $y = 2\epsilon' a$ , it follows that if the algorithm outputs  $G(a)$ , then  $G(a)$  is a  $(1 \pm 4\epsilon)$ -approximation of  $G(m_i)$ . Thus, the first condition of Hybrid-Major follows.

Finally, assume  $\pi_\epsilon(m_i) \geq (20)^5 \sqrt{F_2^{res}(M(W))}$ . By definition,  $m_i \geq \pi_\epsilon(m_i)$  and so  $m_i$  is  $(F_2, 20^{10})$ -heavy with respect to  $M(W)$ . By Lemma 1, we have (except with negligible probability):  $X' > 20^4 Y'$  or  $Y' > 20^4 X'$ . Hence, except with negligible probability, the algorithm does not terminate before the last line. Let  $N_1$  be the constant given by the definition of tractability in Definition 5 ( $N_1$  may depend on the parameter  $k$  from Definition 5, but we apply the definition for  $k = O(1)$  determined by  $\epsilon$ ). We assume  $m_i \geq N_1$  (otherwise the number of elements in the window is constant). Also, let  $r$  be given by Definition 5. By applying Lemma 5 with  $\epsilon' = \frac{1}{\log^{r+1}(N)}$ , we have  $|m_i - a| \leq 2\sqrt{F_2^{res}(M(W))} + \epsilon' m_i \leq 0.01\pi_\epsilon(m_i) + \frac{1}{\log N} \frac{m_i}{\log^r(m_i)} \leq .01\pi_\epsilon(m_i) + \frac{\pi_\epsilon(m_i)}{\log N} \leq .02\pi_\epsilon(m_i)$  for sufficiently large  $N$  (since  $G$  is tractable) and  $b \leq 3\sqrt{F_2^{res}(M(W))} < 0.01\pi_\epsilon(m_i)$ . Since  $b \leq .1\pi_\epsilon(m_i)$  and  $2\epsilon' a \leq 2\epsilon'(m_i + b + \epsilon' m_i) \leq 2 \cdot (.03\pi_\epsilon(m_i)) \leq .1\pi_\epsilon(m_i)$ , then by Lemmas 2 and 3 (which we apply with the same parameters,  $x = m_i, u = a, v = b$ , and  $y = 2\epsilon' a$ ), the algorithm outputs  $G(a)$  which is a  $(1 \pm 4\epsilon)$ -approximation of  $G(m_i)$ . Thus, the second condition of Hybrid-Major follows, which gives the lemma.  $\blacktriangleleft$

Consider the following lemma, which is from [9] (the proof of Lemma 7 uses Predicate (1) from Definition 5).

$\blacktriangleright$  **Lemma 7.** Let  $G$  be a non-decreasing tractable function. Then for any  $k = O(1)$ , there exists  $t = O(1)$  such that for any  $n, N$  and for any  $\epsilon > \log^{-k}(nN)$  the following holds. Let

$D(n, N)$  be a stream and  $W$  be the current window. If there is a  $(G, 1)$ -heavy element  $m_i$  with respect to  $M(W)$ , then there is a set  $S \subseteq [n]$  such that  $|S| = O(\log(N))$  and:  $\pi_\epsilon^2(m_i) = \Omega\left(\log^{-t}(nN) \sum_{j \notin S \cup \{i\}} m_j^2\right)$ .

We now give the algorithm Compute- $G$ -Core, which solves  $G$ -Core (i.e., Problem 2), and prove its correctness. A similar algorithm appears in [9], we repeat it here for completeness, and to help design and understand our main result on universality.

- 1 Generate a pairwise independent hash function  $H : [n] \mapsto \tau$ , where  $\tau = O^*\left(\frac{1}{p}\right)$ .
- 2  $\forall k \in [\tau]$ , compute in parallel  $c_k = \text{Compute-Hybrid-Major}(D_{H_k}, \frac{\epsilon}{4})$ , where  $H_k(i) = \mathbf{1}_{H(i)=k}$ .
- 3 Output  $S = \{c_k : c_k > 0\}$ .

**Algorithm 3:** Compute- $G$ -Core( $D, \epsilon, p$ )

► **Theorem 5.** Algorithm Compute- $G$ -Core solves the  $G$ -Core problem, except with probability asymptotically equal to  $p$ . The algorithm uses  $O^*(1)$  memory bits if  $p = \Omega(1/\log^u(nN))$  and  $\epsilon = \Omega(1/\log^k(nN))$  for some  $u, k \geq 0$ .

**Proof.** Let  $W$  denote the current window. First, except with negligible probability, every positive  $c_i$  is a  $(1 \pm 4 \cdot \frac{\epsilon}{4})$ -approximation of some distinct entry  $G(m_j)$ , which implies that  $c_i$  is a  $(1 \pm \epsilon)$ -approximation of  $G(m_j)$ . Second, assume that there exists a  $(G, 1)$ -heavy entry  $m_i$  with respect to  $M(W)$ . Denote  $X = \sum_{j \neq i} m_j^2 \mathbf{1}_{H(j)=H(i)}$ . By pairwise independence of  $H$ , we have  $E(X) = \frac{1}{\tau}(F_2(M) - m_i^2)$ . By Lemma 7, there exists a set  $S$  and  $t \geq 0$  such that  $|S| = O(\log N)$  and:

$$\pi_\epsilon^2(m_i) = \Omega\left(\frac{\sum_{j \notin S \cup \{i\}} m_j^2}{\log^t(nN)}\right). \quad (5)$$

Let  $\mathcal{L}$  be the event that  $\pi_\epsilon^2(m_i) > 20^{10}X$ , and let  $\mathcal{B}$  be the event that  $\forall j \in S : H(j) \neq H(i)$ . By Markov's inequality, by pairwise independence of  $H$ , and by Equation (5), there exists  $\tau = O^*\left(\frac{1}{p}\right)$  such that:

$$\begin{aligned} P(\neg\mathcal{L}) &= P(\neg\mathcal{L} \mid \mathcal{B}) \cdot P(\mathcal{B}) + P(\neg\mathcal{L} \mid \neg\mathcal{B}) \cdot P(\neg\mathcal{B}) \\ &\leq \frac{E(X \mid \mathcal{B})20^{10}}{\pi_\epsilon^2(m_i)} \cdot 1 + 1 \cdot \frac{O(\log N)}{\tau} \leq O^*\left(\frac{1}{\tau}\right) = p. \end{aligned}$$

If  $\mathcal{L}$  occurs, which happens with probability at least  $1-p$ , then  $c_{H(i)}$  is a  $(1 \pm \epsilon)$ -approximation of  $G(m_i)$  except with negligible probability (by Lemma 6). Thus, the final probability of error is approximately equal to  $p$ .

It is not too hard to see that Algorithm Compute- $G$ -Core uses polylogarithmic memory. The subroutine depth is constant, and there are only polylogarithmically many subroutine calls at each level. At the lowest level, we only do direct computations on the stream that require polylogarithmic space or a smooth histogram computation for  $F_2$  or  $L_2$ , which also requires polylogarithmic space. We get that for any constant  $k$ , there exists a constant  $t$  such that we can solve  $G$ -Core (except with probability  $p$ ) using  $O(\log^t(nN))$  space, where  $\epsilon \geq \log^{-k}(nN)$ . ◀

In Appendix A, we show how to reduce the  $G$ -Sum problem to the  $G$ -Core problem. In particular, we prove the following theorem. The algorithm and proof of correctness follow from [10]. We restate the algorithm and results using our notation for completeness.

► **Theorem 6.** If there is an algorithm that solves  $G$ -Core using memory  $O^*(1)$  and makes one pass over  $D$  except with probability  $O(\log^{-u}(nN))$  for some  $u > 0$ , then there is an algorithm that solves  $G$ -Sum using memory  $O^*(1)$  and makes one pass over  $D$  except with probability at most 0.3.

We can reduce the failure probability to inverse polynomial using standard methods. Combining this with Theorem 5 and Theorem 4, we have Theorem 2.

## 4 Universality

In this section, we show the main result of this paper, Theorem 1, by designing a universal sum algorithm. We first construct a universal core algorithm, which we call  $UCA$ . That is, given a data stream, the algorithm produces a universal core structure with respect to the frequency vector  $(m_1, \dots, m_n)$  defined by the current window  $W$  without knowing the function  $G$  to be approximated in advance. Let  $C$  be a constant and let  $\mathcal{U}(C)$  be the set according to Definition 6. The structure guarantees that, when queried with any function  $G$  from  $\mathcal{U}(C)$  (after processing the stream), it outputs the set  $T$  according to Definition 7.

*Universal Core Algorithm (UCA):* The algorithm constructs a universal core structure  $S$  and our techniques build on the results from Section 3. Algorithm Residual-Approximation from Section 3 does not depend on the function  $G$ , and hence it clearly carries over to our universal setting.

Algorithm Compute-Hybrid-Major depends on  $G$ , so we modify it accordingly. We do not rewrite the whole algorithm, as there are only a few modifications. In Step 1, we set  $\epsilon' = \frac{1}{\log^{10C+1}(N)}$ . We get rid of Steps 8 and 9, and instead create a new Step 8 where we find the index  $j$  of an  $(F_2, 2)$ -heavy element  $m_j$ , if it exists (finding such an index can be done using standard methods, the details of which we omit for brevity). We also create a new Step 9 where we output the triple  $(a, b, j)$  (assuming none of the parallel copies from Step 2 outputs 0).

We also modify Algorithm Compute- $G$ -Core. In particular, the value of  $\tau$  in Step 1 should depend on  $C$ , and we set it to be  $\frac{\log^{10C+2}(nN)}{p}$ . Moreover, we remove Step 3 from the algorithm and store  $c_k$  for each  $k \in [\tau]$  as part of our data structure  $S$  (recall that  $c_k$  is either 0 or a triple  $(a_k, b_k, j_k)$ , where  $a_k, b_k$  are the values computed in the  $k^{\text{th}}$  parallel instance of the subroutine Compute-Hybrid-Major and  $j_k$  is the index of the corresponding  $(F_2, 2)$ -heavy element).

*Querying the Structure:* Given a function  $G \in \mathcal{U}(C)$  as a query to our universal core structure, we explain how to produce the set  $T$  according to Definition 7. For each stored  $c_k$  in the data structure  $S$  ( $k \in [\tau]$ ), if  $c_k = 0$ , then we do not include it in our output set  $T$ . Otherwise, if  $c_k$  is a triple  $(a_k, b_k, j_k)$ , then we include the pair  $(G(a_k), j_k)$  in our set  $T$  as long as  $(1 - 4\epsilon)G(a_k + b_k + 2\epsilon'a_k) \leq G(a_k) \leq (1 + 4\epsilon)G(a_k - b_k - 2\epsilon'a_k)$  (recall  $\epsilon' = \frac{1}{\log^{10C+1}(N)}$ ).

► **Theorem 7.** Fix a constant  $C$  and let  $\mathcal{U}(C)$  be the set of tractable functions corresponding to the definition of universal tractability. Then  $UCA$  is a universal core algorithm with parameters  $\epsilon = \Omega(1/\log^k(nN))$  (for  $0 \leq k \leq C$ ),  $\delta = \Omega(1/\log^u(nN))$  (for  $u \geq 0$ ),  $\alpha = 1$ , and  $\mathcal{G} = \mathcal{U}(C)$ .

**Proof.** The correctness of  $UCA$  essentially follows from the proofs of the results in Section 3. In particular, Lemma 4 still holds since Algorithm Residual-Approximation is unchanged.

Lemma 6 still mostly holds without much modification. Using the same notation as in the original proof, if there is no  $(F_2, 2)$ -heavy element, then the proof of Lemma 6 can still be applied and the modified version of Compute-Hybrid-Major outputs 0 (except with negligible probability). In such a case, the universal core structure stores the value 0. If there is an  $(F_2, 2)$ -heavy element  $m_{i_k}$  and the structure stores  $(a_k, b_k, i_k)$ , then again the proof applies. The reason is that, when querying the universal core structure with a function  $G$ , we check if  $(1 - 4\epsilon)G(a_k + b_k + 2\epsilon'a_k) \leq G(a_k) \leq (1 + 4\epsilon)G(a_k - b_k - 2\epsilon'a_k)$ , in which case the proof argues that  $G(a_k)$  is a  $(1 \pm 4\epsilon)$ -approximation of  $G(m_{i_k})$ . In the case that  $\pi_\epsilon(m_{i_k}) \geq (20)^5 \sqrt{F_2^{res}(M(W))}$ , the proof still goes through since we apply Lemma 5 with  $\epsilon' = \frac{1}{\log^{10C+1}(N)}$ , and we have  $|m_{i_k} - a_k| \leq 2\sqrt{F_2^{res}(M(W))} + \epsilon'm_{i_k} \leq 0.01\pi_\epsilon(m_{i_k}) + \frac{1}{\log N} \frac{m_{i_k}}{\log^{10C}(m_{i_k})} \leq 0.01\pi_\epsilon(m_{i_k}) + \frac{1}{\log N} \frac{m_{i_k}}{\log^r(m_{i_k})} \leq .01\pi_\epsilon(m_{i_k}) + \frac{\pi_\epsilon(m_{i_k})}{\log N} \leq .02\pi_\epsilon(m_{i_k})$  (here, similarly to Lemma 6,  $r$  is the constant given by the definition of universal tractability for  $\mathcal{U}(C)$ , and hence  $r \leq 10C$ ).

Finally, we must argue the correctness of Theorem 5. Using some notation taken from the proof, consider an output  $c_k = (a_k, b_k, i_k)$  (if  $c_k = 0$ , the data structure does not output it to the set  $T$ ) and observe that  $G(a_k)$  for any  $a_k$  satisfying  $(1 - 4\epsilon)G(a_k + b_k + 2\epsilon'a_k) \leq G(a_k) \leq (1 + 4\epsilon)G(a_k - b_k - 2\epsilon'a_k)$  is a  $(1 \pm 4 \cdot \frac{\epsilon}{4})$ -approximation of a distinct entry  $G(m_{i_k})$ . Moreover, if there is a  $(G, 1)$ -heavy element  $m_{i_k}$ , then we again have  $\pi_\epsilon^2(m_{i_k}) = \Omega\left(\log^{-(t+1)}(nN) \sum_{j \notin S \cup \{i_k\}} m_j^2\right)$ . In fact, delving into the proof of Lemma 7 (found in [9]), we see that the specific value of  $t$  depends on  $G$ , and is given by the definition of universal tractability for  $\mathcal{U}(C)$ . Since  $t \leq 10C$  and we choose  $\tau = \frac{\log^{10C+2}(nN)}{p}$ , we get the probability of the bad event  $\neg\mathcal{L}$  (using the same notation from Theorem 5) is bounded by:

$$\frac{E(X | \mathcal{B})20^{10}}{\pi_\epsilon^2(m_{i_k})} + \frac{O(\log N)}{\tau} = \frac{20^{10} \log^{t+1}(nN) \sum_{j \notin S \cup \{i_k\}} m_j^2}{\tau \sum_{j \notin S \cup \{i_k\}} m_j^2} + \frac{O(\log N)}{\tau} \leq p.$$

The rest of the proof goes through in the same way, and hence this gives the theorem.  $\blacktriangleleft$

We now argue how to use our universal core algorithm  $UCA$  as a subroutine to give the main result of the paper. The proof of the theorem below can be found in Appendix B, the argument of which follows a similar one found in [10].

**► Theorem 8.** Fix a constant  $C$  and let  $\mathcal{U}(C)$  be the set of tractable functions from the definition of universal tractability. Suppose there is a universal core algorithm that has parameters  $\epsilon = \Omega(1/\log^k(nN))$  (for  $0 \leq k \leq C$ ),  $\delta = \Omega(1/\log^u(nN))$  (for  $u \geq 0$ ),  $\alpha = 1$ , and  $\mathcal{G} = \mathcal{U}(C)$ , uses polylogarithmic memory in  $n$  and  $N$ , and makes one pass over  $D$ . Then there is a universal sum algorithm that has parameters  $\epsilon = \Omega(1/\log^k(nN))$  (for  $0 \leq k \leq C$ ),  $\delta = 0.3$ , and  $\mathcal{G} = \mathcal{U}(C)$ , uses polylogarithmic space in  $n$  and  $N$ , and makes one pass over  $D$ .

We can reduce the failure probability to inverse polynomial using standard techniques. Our main result, Theorem 1, follows from Theorem 7 and Theorem 8.

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## A $G$ -Sum from $G$ -Core

We now show how to reduce the  $G$ -Sum problem to the  $G$ -Core problem. In particular, we prove the following theorem. The algorithm and proof of correctness follow from [10]. We restate the algorithm and results using our notation for completeness.

► **Theorem 9.** If there is an algorithm that solves  $G$ -Core using memory  $O^*(1)$  and makes one pass over  $D$  except with probability  $O(\log^{-u}(nN))$  for some  $u > 0$ , then there is an algorithm that solves  $G$ -Sum using memory  $O^*(1)$  and makes one pass over  $D$  except with probability at most 0.3.

Note that we can reduce the failure probability from constant to inverse polynomial using standard techniques. Combining this with Theorem 5 and Theorem 4, we have Theorem 2.

Let  $G$  be a tractable function according to Definition 5, and let  $D(n, N)$  be a stream given as input. We show how to construct an algorithm that solves the  $G$ -Sum problem by using our algorithm for  $G$ -Core as a subroutine. In particular, consider the Compute- $G$ -Core algorithm that solves the  $G$ -Core problem. Note that for the output set  $S = \{g'_1, \dots, g'_\ell\}$  maintained by Compute- $G$ -Core, using standard techniques one can easily obtain the explicit set of indices  $\{j_1, \dots, j_\ell\}$  such that  $(1 - \epsilon)G(m_{j_i}) \leq g'_i \leq (1 + \epsilon)G(m_{j_i})$  for each  $1 \leq i \leq \ell$ . Hence, we assume that Compute- $G$ -Core outputs a set of pairs of the form  $\{(g'_1, j_1), \dots, (g'_\ell, j_\ell)\}$ .

In the language of [10], Compute- $G$ -Core produces a  $(1, \epsilon)$ -cover with respect to the vector  $G(M(W)) = (G(m_1), \dots, G(m_n))$  with probability at least  $1 - \delta$ , where  $\epsilon = \Omega(1/\log^k(nN))$  (for any  $k \geq 0$ ) and  $\delta = \Omega(1/\log^u(nN))$  (for any  $u \geq 0$ ). Given the tractable function  $G$ , our algorithm for  $G$ -Sum is as follows:

- 1 Generate  $\phi = O(\log(n))$  pairwise independent, uniform zero-one vectors  $H_1, \dots, H_\phi : [n] \rightarrow \{0, 1\}$ , and let  $h_i^k = H_k(i)$ . Let  $D_k$  be the substream defined by  $D_{H_1 H_2 \dots H_k}$ , and let  $G(M(W_k))$  denote  $(G(m_1), \dots, G(m_n))$  for the substream  $D_k$  and window  $W$  (where  $k \in [\phi]$ ).
- 2 Maintain, in parallel, the cores  $Q_k = \text{Compute-}G\text{-Core}(D_k, \frac{\epsilon^2}{\phi^3}, \epsilon, \frac{1}{\phi})$  for each  $k \in [\phi]$ .
- 3 If  $F_0(G(M(W_\phi))) > 10^{10}$ , then output 0.
- 4 Otherwise, precisely compute  $Y_\phi = |G(M(W_\phi))|$ .
- 5 For each  $k = \phi - 1, \dots, 0$ , compute  $Y_k = 2Y_{k+1} - \sum_{(g'_i, j_i) \in Q_k} (1 - 2h_{j_i}^k) g'_i$ .
- 6 Output  $Y_0$ .

### Algorithm 4: Compute- $G$ -Sum( $D, \epsilon$ )

Note that, in our paper, Compute- $G$ -Core( $D, \epsilon, \delta$ ) only takes three parameters (the stream  $D$ , error bound  $\epsilon$ , and failure probability  $\delta$ ), while the algorithm from [10] assumes four parameters of the form Compute- $G$ -Core( $D, \alpha, \epsilon, \delta$ ). Here,  $D$ ,  $\epsilon$ , and  $\delta$  have the same meaning as in our paper. The parameter  $\alpha$  controls how heavy an element needs to be (according to the function  $G$ ) in order to necessarily be in the output set of Compute- $G$ -Core. That is, in the set  $T = \{(g'_1, j_1), \dots, (g'_\ell, j_\ell)\}$  output by Compute- $G$ -Core, if there is an  $i$  such that  $m_i$  is  $(G, \alpha)$ -heavy with respect to  $M(W)$ , then  $i \in \{j_1, \dots, j_\ell\}$ . We solve the  $G$ -Core problem for  $\alpha = 1$ , but Algorithm Compute- $G$ -Sum needs the problem solved for  $\alpha = \frac{\epsilon^2}{\phi^3}$ . However, using standard techniques, we can reduce the problem of solving  $G$ -Core for  $\alpha = \frac{\epsilon^2}{\phi^3}$  to the same problem for  $\alpha = 1$ .

► **Theorem 10.** For any tractable function  $G$ , Algorithm Compute- $G$ -Sum outputs a  $(1 \pm \epsilon)$ -approximation of  $|G(M(W))|$  except with probability at most 0.3, where  $\epsilon = \Omega(1/\log^k(nN))$  for any  $k \geq 0$ . The algorithm uses memory that is polylogarithmic in  $n$  and  $N$ .

**Proof.** The proof of this theorem follows directly from Theorem 1 in [10]. ◀

## B Universal Sum from Universal Core

We now prove Theorem 8. The algorithm and proof are similar to that of the reduction from the  $G$ -Sum problem to the  $G$ -Core problem found in Appendix A, except that we need to carry out the argument within our universal framework. As mentioned, the algorithm and correctness follow from [10]. We do not rewrite the whole algorithm, but instead describe the necessary modifications that need to be made from Appendix A.

Let  $D(n, N)$  be a stream given as input to our universal sum algorithm. Let  $UCA$  be our universal core algorithm from Theorem 7, Section 4, the parameters of which are specified in our universal sum algorithm description.

*Universal Sum Algorithm:* We describe the modifications that need to be made to Algorithm Compute- $G$ -Sum from Appendix A.

In Step 2, instead we need to maintain and store the output  $Q_k = UCA$  with parameters  $\alpha = \frac{\epsilon^2}{\phi^3}$ ,  $\epsilon$  (i.e., the one given as input to our universal sum algorithm),  $\delta = \frac{1}{\phi}$ , and  $\mathcal{G} = \mathcal{U}(C)$  for each  $k \in [\phi]$  (in the  $k^{\text{th}}$  parallel iteration,  $UCA$  is given the stream  $D_k$  as input). As in Appendix A, we construct a universal core structure for  $\alpha = 1$ , but we can reduce the problem of  $\alpha = \frac{\epsilon^2}{\phi^3}$  to  $\alpha = 1$ . Note that  $Q_k$  is of the form  $\{(a_1, b_1, j_1), \dots, (a_\ell, b_\ell, j_\ell)\}$  ( $Q_k$  may have 0's as well, which we simply ignore). For each such triple  $(a_i, b_i, j_i)$ , we also store the value of  $h_{j_i}^k = H_k(j_i)$ .

In Step 3, instead we check if  $F_0(M(W_\phi)) \leq 10^{10}$ , and if so we store  $M(W_\phi)$  (recall  $M(W_\phi)$  denotes the frequency vector  $(m_1, \dots, m_n)$  for the substream  $D_\phi$  induced by  $W$ ). We remove Steps 4, 5, and 6.

*Querying the Structure:* Now, given a function  $G \in \mathcal{U}(C)$ , we explain how to query the universal sum structure output by our universal sum algorithm to approximate  $|G(M(W))|$ . In particular, for each  $k$  we first query the universal core structure output by  $UCA$  to get a set  $Q'_k = \{(x_1, j_1), \dots, (x_\ell, j_\ell)\}$ . Then, we compute  $Y_\phi = |G(M(W_\phi))|$  and, for each  $k = \phi - 1, \dots, 0$ , we recursively compute  $Y_k$  according to:

$$Y_k = 2Y_{k+1} - \sum_{(x_i, j_i) \in Q_k} (1 - 2h_{j_i}^k)x_i.$$

Once each  $Y_k$  has been computed for  $0 \leq k \leq \phi$ , we output  $Y_0$ .

► **Theorem 11.** Fix a constant  $C$  and let  $\mathcal{U}(C)$  be the set of tractable functions corresponding to the definition of universal tractability. There is a universal sum algorithm with parameters  $\epsilon = \Omega(1/\log^k(nN))$  (for  $0 \leq k \leq C$ ),  $\delta = 0.3$ , and  $\mathcal{G} = \mathcal{U}(C)$ . The algorithm uses polylogarithmic space in  $n$  and  $N$  and makes a single pass over  $D$ . When querying the universal sum structure (output by the universal sum algorithm) with a function  $G \in \mathcal{U}(C)$ , it outputs a  $(1 \pm \epsilon)$ -approximation of  $|G(M(W))|$  except with probability at most 0.3.

**Proof.** The proof of this theorem follows directly from Theorem 1 in [10]. ◀