Sublinear-Space Streaming Algorithms for Estimating Graph Parameters on Sparse Graphs

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Abstract

Fundamental graph problems have been studied extensively in the RAM model of computation. Over the last decade, there have been great advances in streaming algorithms for such graph problems. In this paper, we design new streaming algorithms for three problems: maximum independent set, minimum dominating set, and maximum matching. Our methods approximate these three graph parameters in sparse graph classes, i.e., where the number of edges is linear in the number of vertices. Since our aim is sublinear working-space algorithms, and each of the three parameters can have linear size even in sparse graphs, we can only estimate these parameters instead of actually producing solutions in our working space. We obtain the following two families of results:

- **Estimating Max Independent Set via the Caro-Wei bound**: Caro and Wei each showed that $\lambda = \sum_v 1/(d(v)+1)$ lower bounds the size of a max independent set, where $d(v)$ is the degree of vertex v. For graphs with constant average degree, $\bar{d}$, and with maximum degree $\Delta = \mathcal{O}(\varepsilon^2 \bar{d}^{-3} n)$, we develop algorithms that, with at least $1-\delta$ success probability:
  - In the *online streaming* model, return an independent set of size $1\pm\varepsilon$ times λ. This improves on results by Halldórsson et al. [Algorithmica '16], with less space, i.e., $\mathcal{O}(\log \varepsilon^{-1} \cdot \log n \cdot \log \delta^{-1})$, faster update time, i.e., $\mathcal{O}(\log \varepsilon^{-1})$, and bounded success probability.
  - In insertion-only streams, approximate the Caro-Wei bound within a factor of $1 \pm \varepsilon$, in one pass and in $\mathcal{O}(\bar{d}\varepsilon^{-2} \log n \cdot \log \delta^{-1})$ space. This aligns with the state-of-art result of Cormode et al. [ISCO '18], though our method also works in the *online streaming* model. If the stream is vertex-arrival and random-order, the space can be further reduced to $\mathcal{O}(\log(\bar{d}\varepsilon^{-1}))$. Trading off extra space and post processing, we can remove the max-degree constraint.
- **Sublinear-Space Algorithms on Forests**: When the input is a forest, Esfandiari et al. [SODA '15, TALG '18] showed space lower bounds[1] of $\Omega(\sqrt{n})$ for 1-pass randomized algorithms which can approximately estimate our three graph parameters. We narrow the gap between upper and lower bounds by designing new streaming algorithms that approximate:
  - The size of a max independent set within $3/2 \cdot (1 \pm \varepsilon)$ in one pass and in $\log^{\mathcal{O}(1)} n$ space, and within $4/3 \cdot (1 \pm \varepsilon)$ in two passes and in $\tilde{\mathcal{O}}(\sqrt{n})$ space; the lower bound of Esfandiari et al. holds for approximation better than $4/3$.
  - The size of a min dominating set within $3 \cdot (1 \pm \varepsilon)$ in one pass and in $\log^{\mathcal{O}(1)} n$ space, and within $2 \cdot (1 \pm \varepsilon)$ in two passes and in $\tilde{\mathcal{O}}(\sqrt{n})$ space; the lower bound of Esfandiari et al. holds for approximation better than $3/2$.
  - The size of a max matching within $2 \cdot (1 \pm \varepsilon)$ in one pass and in $\log^{\mathcal{O}(1)} n$ space, and within $3/2 \cdot (1 \pm \varepsilon)$ in two passes and in $\tilde{\mathcal{O}}(\sqrt{n})$ space; the lower bound of Esfandiari et al. holds for approximation better than $3/2$.

[1] In fact, the lower bounds hold even when each component is a path! For deterministic algorithms, the lower bound can be improved to $\Omega(n)$.

Leibniz International Proceedings in Informatics
Schloss Dagstuhl – Leibniz-Zentrum für Informatik, Dagstuhl Publishing, Germany

2012 ACM Subject Classification Theory of computation → Streaming models; Theory of computation → Streaming, sublinear and near linear time algorithms

Keywords and phrases Graph Algorithms, Data Streams Model, Caro-Wei Bound, Independent Set, Dominating Set, Matching

Digital Object Identifier 10.4230/LIPIcs...

1 Introduction

Numerous real-world problems can be nicely modeled and solved as graphs. Consequently, over the last century or so, a wide range of graph problems have been studied, such as independent set, dominating set, and matching. For example, independent set models optimization and scheduling problems where conflicts should be avoided [1, 21, 24, 31], dominating set models guardian selection problems [36, 38, 42, 43, 47], while a matching models similarity [14, 20, 45] and inclusion dependencies [3, 16].

Parameters of interest: Given an undirected graph G, comprising vertices V and edges E, let n and m be the sizes of V and E, respectively. A subset S of V is an independent set if and only if the subgraph induced by S contains no edges. Alternatively, S is a dominating set if and only if every vertex in V is either in S itself or adjacent to some vertex in S. A subset M of edges E is a matching if and only if no pair of edges in M share a vertex. The three parameters of interest to us are the size of a maximum independent set, which we denote by the independence number, β; the size of a minimum dominating set, which we denote by the domination number, γ; and the size of a maximum matching, which we denote by the matching number, ϕ. It is well known that $n - \phi \geq \beta \geq n - 2\phi$.

Deciding whether each of these parameters is larger (for max) or smaller (for min) than a specified quantity is NP-complete [30]. There are many *approximation algorithms* that run in polynomial time and return a solution within a guaranteed factor of optimum. For instance, a greedy algorithm for maximum matching outputs a 2-approximation. Similarly, there exist greedy algorithms that $\mathcal{O}(\Delta)$-approximate the maximum independent set [23] and approximate within $\mathcal{O}(\ln \Delta)$ the minimum dominating set [27, 33], where Δ is the maximum degree. More promising results are obtained on sparse graphs, i.e., graphs with a linear number of edges, which we study in this paper. For example, the Caro-Wei bound [7, 46], given by $\sum_{v \in V(G)} \frac{1}{1+\deg(v)}$, which we denote by λ, is a lower bound on β.

Data streams: We focus on estimating these graph parameters in the data stream model. In the *data stream* model, instead of being pre-stored in memory, data elements arrive sequentially, and the available memory is typically much less than the complete data set size. The *semi-streaming* model [19], where $\mathcal{O}(n \log n)$ bits are allowed, is commonplace for general graphs. In this paper we consider *sparse* graphs, so we restrict our space allowance to $o(n)$ bits, as $\mathcal{O}(n)$ bits would be sufficient to store the entire sparse graph. A graph stream is *edge-arrival* if edges arrive one by one, sequentially, in arbitrary order. It is *vertex-arrival* if it comprises a sequence of (vertex, vertex-list) pairs, (u_i, A_i), where the list A_i comprises the subset of the vertices $\{u_j\}_{j<i}$ that have occurred previously in the stream that are adjacent to u_i. We call a graph stream *insertion-only* if there are no deletions. While in a *turnstile* stream, an element (vertex or edge) can be deleted if it has been inserted and not been deleted. Moreover, the order of arrival can be either *arbitrary* or (uniformly) *random*.

Halldórsson et al. [22] introduced the *online streaming* model, combining the data-stream and online models. In the *online streaming* model, elements arrive as a stream. After each arrival, on demand, the algorithm and data structures must be able to report efficiently a

valid solution. Usually, the algorithm will start with a initial solution, and modify it based on each newly-arrived element. Similar to the *online* model, each decision on the solution is irrevocable. To clarify the efficiencies, we distinguish between *working space*, involved in computing the solution, and an additional *solution space* for storing the solution itself. This solution space may be larger than the working space, but is write only.

1.1 Previous Results

Estimating the size of a maximum matching in a sparse graph has been extensively studied in the streaming setting. When the input is a tree, several $(2+\varepsilon)$-approximation algorithms are known. Esfandiari et al. [18] designed an $\tilde{\mathcal{O}}(\sqrt{n})$-space algorithm for insertion-only streams, where $\tilde{\mathcal{O}}$ notation suppresses a poly-logarithmic factor. The space was further reduced to $\log^{\mathcal{O}(1)} n$ by Cormode et al. [12], while Bury et al. [5] generalized it to turnstile streams.

However, there is relatively little research on independent set and dominating set in this scenario. Halldórsson et al. [22] studied general hypergraphs and gave a one-pass insertion-only streaming algorithm in $\mathcal{O}(n)$ space, outputting an independent set with size at least the Caro-Wei bound, λ, in expectation. Their algorithm suits the *online streaming* model. Cormode et al. [10] designed an algorithm that $(1\pm\varepsilon)$-approximates λ with constant success probability in $\mathcal{O}(\varepsilon^{-2}\bar{d}\log n)$ space. They also showed a nearly tight lower bound: every randomized one-pass algorithm with constant error probability requires $\Omega(\varepsilon^{-2}\bar{d})$ space to c-approximate λ. Meanwhile, Esfandiari et al. [18] established space lower bounds for estimating ϕ in a forest: for every 1-pass randomized streaming approximation algorithm with factor better than $3/2$, we need $\Omega(\sqrt{n})$ space. Their techniques can be adapted to give lower bounds for other parameters: $\Omega(\sqrt{n})$ bits are required to approximate β better than $4/3$, and γ better than $3/2$.

1.2 Our Results

We design new sub-linear streaming algorithms for estimating fundamental graph quantities such as independent set, dominating set, and matching in sparse graphs. These sparse graph classes have linear number of edges and hence achieving sublinear-space is non-trivial. Moreover, if we allow only sublinear-space then we are restricted to only estimating these quantities (as opposed to say finding an independent set) since each of these quantities can have linear size on sparse graph classes.

Our results fall into two categories:

- **Approximating independent set via the Caro-Wei bound in bounded average degree graphs**: Bounded average degree graphs contain several interesting graph classes such as planar graphs, bounded treewidth, bounded genus, H-minor-free, etc.
- **Approximating graph quantities on forests**: Trees, and more generally forests, are one of the simplest examples of sparse graphs. However, Esfandiari et al. [18] showed that one needs $\Omega(n)$ space[2], i.e., essentially store the whole graph, to estimate (in 1-pass streaming) fundamental graph quantities such as independence number, domination number and matching number on forests. In fact, the lower bound example is quite simple: we cannot distinguish between (multiple copies of) two P_3 versus one P_2 and one P_4 where P_n is a path on n vertices.

[2] This lower bound is for deterministic algorithms. The lower bound for randomized algorithms is $\Omega(\sqrt{n})$

1.2.1 Approximating the Caro-Wei bound in bounded avg. degree graphs

We approximate the independence number in graphs with bounded average degree via the Caro-Wei bound, λ. As Boppana et al. [4] show, λ is at least the Turán Bound [44], i.e., $\beta \geq n/(\bar{d}+1)$. Hence, every $\mathcal{O}(1)$-approximation of λ gives a $\mathcal{O}(\bar{d})$-approximation of β. Our results are summarized in Table 1, and our main algorithm is called CARAWAY.

There is a known offline algorithm for estimating λ, based on a random permutation of the vertices. CARAWAY simulates a random permutation via hash functions drawn uniformly at random from an ε-min-wise hash family[3]. In the analysis, bounding what we call the approximation error *rate*[4], ε, is non-trivial, due to positive correlation between some pairs of vertices. We provide the first analysis that bounds ε, when the max degree is not too large. Our analysis also provides concentration bounds for the permutation-based offline algorithm.

For graphs with average degree $\bar{d}$ and max degree $\Delta \in \mathcal{O}(\varepsilon^2 \bar{d}^{-3} n)$, we present an arbitrary-order edge-arrival algorithm, CARAWAY, that $(1 \pm \varepsilon)$-approximates λ. Our algorithm fails with probability at most δ, and takes working space $\mathcal{O}(\bar{d}\varepsilon^{-2} \log n \cdot \log \delta^{-1})$. By allowing additional, write-only *solution space* to store the output, the modification CARAWAY$_1$ reports an actual solution set in the *online streaming* model with $\mathcal{O}(\log \varepsilon^{-1})$ update time and working space only $\mathcal{O}(\log \varepsilon^{-1} \cdot \log n \cdot \log \delta^{-1})$. The max-degree constraint is only used to provide a bound on the failure probability, δ; if only an *in-expectation* bound suffices, then this constraint can be disregarded.

Our other modified version, CARAWAY$_2$, removes the max-degree assumption but has a slightly larger space requirement. Failing with low probability, i.e., at most $n^{-c'}$ for some constant c', CARAWAY$_2$ requires $\mathcal{O}(\bar{d}^4 \varepsilon^{-2} \log^2 n)$ space, and returns a $(1 \pm \varepsilon)$-approximation. Since post-processing is a component, CARAWAY$_2$ cannot be used in the *online streaming* model. Moreover, if the stream is vertex arrival and random-order, then the space need of our algorithm can be further reduced to only $\mathcal{O}(\log(\bar{d}\varepsilon^{-1}) \log \delta^{-1})$.

Comparison with Cormode et al. [10] : Algorithms by Cormode et al. [10] report β. However, since their approach relied on estimating the degrees of uniformly-at-random sampled vertices, we believe it cannot be modified to report an actual independent set. When the stream is insertion-only and the maximum degree is not too large, CARAWAY achieves asymptotically the same approximation ratio in the same space as Cormode et al. [10]. Importantly, the *online streaming* adaptation, CARAWAY$_1$, can output an independent set solution. Moreover, in the vertex-arrival and random-order models, CARAWAYrequires much less space than the method of Cormode et al. [10].

Comparison with Halldórsson et al. [22] : The online streaming algorithm by Halldórsson et al. [22] has update time $\mathcal{O}(\log n)$ in working space $\mathcal{O}(n)$. CARAWAY$_1$ has lower update time, $\mathcal{O}(\log \varepsilon^{-1})$, and less working space, i.e., $\mathcal{O}(\log \varepsilon^{-1} \cdot \log n \cdot \log \delta^{-1})$. Besides, we offer a guaranteed probability that our estimate is within error rate, ε; there is no such guarantee from Halldórsson et al. [22].

1.2.2 Estimating parameters β, γ and ϕ on Streamed Forests

Our results on estimating the independence number, β, domination number, γ and matching number, ϕ on forests in the streaming model are summarized in Table 3. We trade off approximation ratio for space and number of passes, and obtain two classes of results.

[3] See Section 2.3 for the definition of ε-min-wise hash families.
[4] The relative error between the estimate and the actual value.

Table 1 Streaming Caro-Wei Bound Approximation. *Arrival* options are *Edge* for edge-arrival or *Vertex* for vertex-arrival. Stream *type* options are *Insert.* for insertion-only or *Turns.* for turnstile. Stream *order* options are *Arb.* for arbitrary or *Ran.* for random. Online is "Yes" if it is an *online streaming* algorithm. Column *Prob.* indicates success probability, where $c \in (0,1)$ and $c' > 1$ are arbitrary constants. Theorems marked with '*' require the max vertex degree to be at most $\frac{\varepsilon^2 n}{3(\bar{d}+1)^3}$: if an *in-expectation guarantee* suffices, we remove this condition.

Arrival	Type	Order	Factor	(Work) Space	Online	Prob.	Reference
Edge	Insert.	Arb.	1	$\mathcal{O}(n)$	Yes	*Expected*	[22]
Edge	Turns.	Arb.	$(1 \pm \varepsilon)$	$\mathcal{O}(\bar{d}\varepsilon^{-2} \log n)$	No	c	[10]
Edge	Insert.	Arb.	$(1 \pm \varepsilon)$	$\mathcal{O}(\bar{d}\varepsilon^{-2} \log n)$	No	c	Theorem 5*
Edge	Insert.	Arb.	$(1 \pm \varepsilon)$	$\mathcal{O}(\log \varepsilon^{-1} \log n)$	Yes	c	Theorem 9*
Edge	Insert.	Arb.	$(1 \pm \varepsilon)$	$\mathcal{O}(\bar{d}^2\varepsilon^{-2} \log^2 n)$	No	$n^{-c'}$	Theorem 12
Vertex	Insert.	Ran.	$(1 \pm \varepsilon)$	$\mathcal{O}(\log(\bar{d}\varepsilon^{-1}))$	No	c	Theorem 13*
Either	Either	Arb.	c	$\Omega(\bar{d}/c^2)$	-	c	[10]

- **One-pass $\log^{\mathcal{O}(1)} n$ space algorithms**:
  - an $3/2 \cdot (1 \pm \epsilon)$-approximation for the independence number, β
  - an $3 \cdot (1 \pm \epsilon)$-approximation for the domination number, γ
  - an $2 \cdot (1 \pm \epsilon)$-approximation for the matching number, ϕ.

  These approaches rely on approximating both the numbers of leaves[5] and non-leaf vertices in a stream.

- **Two-pass $\tilde{\mathcal{O}}(\sqrt{n})$ space**[6]:
  - an $4/3 \cdot (1 \pm \epsilon)$-approximation algorithm for β
  - an $2 \cdot (1 \pm \epsilon)$-approximation algorithm for γ
  - an $3/2 \cdot (1 \pm \epsilon)$-approximation algorithm for ϕ.

  Our results narrow the gap between the upper and lower bounds of Esfandiari et al. [18]. Our innovation here is including the notion of support vertices which has previously been used in structural graph theory [15]. A *support vertex* is one that is adjacent to one or more leaves. We anticipate that this useful quantity could be used to design (streaming) algorithms for many other graph problems on trees and forests.

1.3 Further Related Streaming Work

To further set the scene for our contributions, we describe some of the streaming-algorithm context. On graphs with bounded arboricity, α, there are known approximation algorithms for estimating ϕ. For insertion-only streams, Esfandiari et al. [18] developed a $(5\alpha + 9)$-approximation algorithm that requires $\tilde{\mathcal{O}}(\alpha n^{2/3})$ space. Cormode et al. [12] traded off approximation ratio for space, and showed a $(22.5\alpha + 6)$-approximation algorithm in only $\mathcal{O}(\alpha \log^{\mathcal{O}(1)} n)$ space. McGregor and Vorotnikova [34] showed that the algorithm of Cormode et al. [12] can achieve a $(\alpha + 2)$-approximation via a tighter analysis. For turnstile streams, Chitnis et al. [9] designed a $(22.5\alpha + 6)$-approximation algorithm using $\tilde{\mathcal{O}}(\alpha n^{4/5})$ space. Bury et al. [5] improved the approximation ratio to $(\alpha + 2)$, their

[5] A leaf is a vertex with degree one.
[6] The symbol $\tilde{\mathcal{O}}$ suppresses the poly-logarithmic factor in the bound.

Table 2 Upper & lower bounds for approximately estimating graph parameters γ, β and ϕ in streaming model when the input graph is a forest. Each algorithm succeeds with high probability.

Problem	Arrival	Type	Order	Pass	Factor	Space	Reference
γ	Edge	Turnstile	Arb.	1	$3(1 \pm \varepsilon)$	$\log^{\mathcal{O}(1)} n$	Theorem 32
γ	Edge	Turnstile	Arb.	2	$2(1 \pm \varepsilon)$	$\tilde{\mathcal{O}}(\sqrt{n})$	Theorem 38
γ	Any	Any	Arb.	1	$3/2 - \varepsilon$	$\Omega(\sqrt{n})$	[18]
β	Edge	Turnstile	Arb.	1	$3/2(1 \pm \varepsilon)$	$\log^{\mathcal{O}(1)} n$	Theorem 22
β	Edge	Turnstile	Arb.	2	$4/3(1 \pm \varepsilon)$	$\tilde{\mathcal{O}}(\sqrt{n})$	Theorem 25
β	Any	Any	Arb.	1	$4/3 - \varepsilon$	$\Omega(\sqrt{n})$	[18]
ϕ	Edge	Insertion	Arb.	1	$2(1 \pm \varepsilon)$	$\tilde{\mathcal{O}}(\sqrt{n})$	[18]
ϕ	Edge	Insertion	Arb.	1	$2(1 \pm \varepsilon)$	$\log^{\mathcal{O}(1)} n$	[12]
ϕ	Edge	Turnstile	Arb.	1	$2(1 \pm \varepsilon)$	$\log^{\mathcal{O}(1)} n$	[6], Theorem 39
ϕ	Edge	Turnstile	Arb.	2	$3/2(1 \pm \varepsilon)$	$\tilde{\mathcal{O}}(\sqrt{n})$	Theorem 45
ϕ	Any	Any	Arb.	1	$3/2 - \varepsilon$	$\Omega(\sqrt{n})$	[18]

algorithm requires $\tilde{\mathcal{O}}(\alpha n^{4/5})$ space in edge-arrival streams, but only $\mathcal{O}(\log n)$ in vertex-arrival streams.

Switching the arrival order from arbitrary to random, Monemizadeh et al. [37] converted known constant-time RAM-model approximation algorithms into constant-space streaming algorithms. Their algorithm approximates γ in bounded-degree graphs, with an additive error term of εn. Peng and Sohler [41] further extended this result to graphs of bounded average degree. They also showed that in planar graphs β can be $(1+\varepsilon)$-approximated using constant, but still massive, space, i.e., $\mathcal{O}(2^{(1/\varepsilon)^{(1/\varepsilon)^{\log^{\mathcal{O}(1)}(1/\varepsilon)}}})$. Cormode et al. [10] also showed that in a vertex-arrival stream with very large average degree, there exists an algorithm that $(\log n)$-approximates the Caro-Wei Bound using $\mathcal{O}(\log^3 n)$ space.

Meanwhile, Chitnis and Cormode [8] adapted the lower-bound reduction technique from Assadi et al. [2] and showed that, for graphs with arboricity $\alpha + 2$, for $\alpha \geq 1$, every randomized $\alpha/32$-approximation algorithm for minimum dominating set requires $\Omega(n)$ space. This lower bound holds even under the vertex-arrival model.

2 Preliminaries

2.1 Graphs and Problem Definitions

For an undirected graph, G, given a subset of its vertices, $S \subseteq V(G)$, let $N(S)$ be the neighbors of S (excluding S itself), and $N[S]$ be $S \cup N(S)$. Let $d(v)$ be the degree of vertex v, Δ be the max degree in G, and $\bar{d}$ be the average degree in G. Let $Deg_i(G)$ be the subset of V with $d(v) = i$, and $Deg_{\geq i}(G)$ be the subset of V with $d(v) \geq i$. Many of the estimates in this paper use the notion of a support vertex [15] which is defined as a vertex that is adjacent to at least one leaf. We denote the set of support vertices by $Supp(G)$.

2.2 Fundamental Streaming Algorithms

To support effective algorithm development, we make several assumptions about the stream. For convenience, consistent with prior art, we assume the vertex identifiers are $[n] = \{1, 2, \ldots, n\}$, and that the value of n is known in advance. Second, our results are developed for certain graph classes, hence we assume that at the *end* of the stream, the graph is guaranteed to be in the target class. But there is no such requirement *during* the stream.

We employ several streaming primitives. Consider a vector $x \in \mathbb{R}^n$, whose coordinates are updated by a turnstile stream of length m, where increments can be negative. Consistent with prior art, assume each update is in the range $[-M, M]$. Let x_i be the final value of coordinate i. For $p > 0$, the L_p norm of x is $\|x\|_p = (\sum_{i=1}^{n} |x_i|^p)^{1/p}$. The L_0 *norm* of x is $\|x\|_0 = (\sum_{i=1}^{n} |x_i|^0)$. Both L_p norm and L_0 *norm* can be $(1 \pm \varepsilon)$-approximated in streams in small space.

▶ **Theorem 1.** *[28] For all $p \in (0, 2)$, given $\varepsilon, \delta \in (0, 1)$, there exists an algorithm that uses $\mathcal{O}(\varepsilon^{-2} \log(mM) \log \delta^{-1})$ space and with probability at least $1 - \delta$ reports a value in the range $(1 \pm \varepsilon) \|x\|_p$. Both update time and reporting time are in $\tilde{\mathcal{O}}(\varepsilon^{-2} \log \delta^{-1})$.*

▶ **Theorem 2.** *[29] Given $\varepsilon, \delta \in (0, 1)$, there is an algorithm that with probability at least $1 - \delta$ that $(1 \pm \varepsilon)$-approximates $\|x\|_0$. It uses space $\mathcal{O}(\varepsilon^{-2} \log(\delta^{-1}) \log(n)(\log(1/\varepsilon) + \log\log(mM)))$, and its update and reporting time are both in $\mathcal{O}(\log \delta^{-1})$.*

Definition: Given $k \leq n$,vector x is k-sparse if it has at most k non-zero coordinates.

▶ **Theorem 3.** *[11] Given k and error probability $\delta \in (0, 1)$, there exists an algorithm that recovers a k-sparse vector exactly with probability $1 - \delta$ with space usage $\mathcal{O}(k \log n \cdot \log(k/\delta))$.*

The heavy hitter problem seeks all coordinates from a frequency vector with value greater than a certain threshold. In this work, we focus on a variant of this problem where we have to output all coordinates i such that $|x_i| \geq \varepsilon \|x\|_1$. Cormode and Muthukrishnan [13] obtained the following theorem, based on Count-Min sketch.

▶ **Theorem 4.** *[13] Given $\varepsilon, \tau \in (0, 1)$, there is an algorithm that outputs all items with frequency at least $(\varepsilon + \tau)\|x\|_1$, and with probability $1 - \delta$ outputs no items with frequency less than $\tau \|x\|_1$. It has update time $\mathcal{O}(\log n \cdot \log(2 \log n/(\delta\tau)))$ and query time $\mathcal{O}(\varepsilon^{-1})$, and uses $\mathcal{O}(\varepsilon^{-1} \log n \cdot \log(2 \log n/(\delta\tau)))$ space.*

2.3 ε-min-wise Hash Functions

To simulate uniform-at-random selection, we invoke carefully chosen hash families. A family of hash functions, $\mathcal{H} = \{h : [n] \to [m]\}$, is "$\varepsilon$-min-wise" if for every $A \subset [n]$ and $x \in [n] \setminus A$,

$$\Pr_{h \in \mathcal{H}}[\forall a \in A, h(x) < h(a)] = \frac{1 \pm \varepsilon}{|A| + 1}.$$

Indyk [25] showed that the ε-min-wise property can be achieved via a $\mathcal{O}(\log \varepsilon^{-1})$-wise hash family, under the constraints that $|A| \in \mathcal{O}(\varepsilon m)$. Its space usage is $\mathcal{O}(\log \varepsilon^{-1} \cdot \log m)$ and its computation time is $\mathcal{O}(\log \varepsilon^{-1})$.

3 Estimating Caro-Wei Bound in Graphs of Bounded Avg. Degree

In this section, we introduce efficient streaming algorithms to estimate the independence number, β in graphs with bounded average degree. As mentioned in the introduction, a *folklore* offline greedy algorithm in expectation outputs an independent set of size the Caro-Wei Bound, λ. Given a permutation of the vertices chosen uniformly at random, the greedy algorithm adds a vertex to the solution whenever it is *earlier* in the permutation than all its neighbors. Importantly, storing such a permutation explicitly requires $\Omega(n \log n)$ bits, which breaks our space requirement.

We overcome this by simulating a random permutation with a hash function drawn randomly from an ε-min-wise hash function family, $\mathcal{H}$. To generate a pseudo-permutation,

we order vertices by their hash values. Our solution will consist only of vertices whose hash value is lower than all of their neighbours; in this way each edge received will disqualify a vertex. By the property of the ε-min-wise hash, the probability that each vertex has smallest hash is approximately proportional to the reciprocal of its neighborhood size. In general, the neighborhood might be large, i.e., $\Theta(n)$, so we rely on standard sampling and recovery techniques in our algorithm, CARAWAY, Algorithm 1.

Algorithm 1 CARAWAY, estimating Caro-Wei bound

Input: Average degree, $\bar{d}$, error rate, ε.
Initialization: Uniformly-at-random select $h \in \mathcal{H}$, an ε-min-wise hash from $[n]$ to $[n^3]$
$p \leftarrow 4(\bar{d}+1)/(\varepsilon^2 n)$
$S \leftarrow pn$ vertices sampled uniformly at random from $[n]$
for all $e = (u, v)$ in the stream **do**
    if $(u \in S) \wedge (h(u) \geq h(v))$ **then**
        Remove u from S
    Perform the same operation on v
return $\widehat{\lambda} \leftarrow |S|/p$

We choose the output universe of $\mathcal{H}$ in CARAWAY to be $[n^3]$ for two reasons. First, with high probability, there are no colliding pairs. Second, assuming $\varepsilon \in \omega(n^{-2})$ ensures that the ε-min-wise property can be applied to every subset of size at least n.

▶ **Theorem 5.** *Given $\varepsilon \in (0,1)$, for every graph with average degree $\bar{d}$ and max degree $\Delta \leq \varepsilon^2\, n/(3\,(\bar{d}+1)^3)$, CARAWAY is an insertion-only algorithm that $(1\pm\varepsilon)$-approximates λ. It needs work space $\mathcal{O}(\bar{d}\,\varepsilon^{-2}\log n)$ and succeeds with probability at least $2/3$.*

To prove Theorem 5, we start with a result on the expectation of $\widehat{\lambda}$.

▶ **Lemma 6.** [⋆][7] $E[\widehat{\lambda}] = (1 \pm \varepsilon)\,\lambda$.

We next show that, with constant probability, estimate $\widehat{\lambda}$ is no more than a $1 \pm \varepsilon$ factor away from its mean, $E[\widehat{\lambda}]$. Let X_v be a indicator random variable for the event that v is both sampled and not removed from S. Let $X = \sum_{v \in V} X_v$. Since some pairs of X_v and X_u are positively correlated, we cannot directly apply a Chernoff(-like) bound. There are three cases to consider.

- Vertices u and v are adjacent: $u \in N[v]$ and $v \in N[u]$. If X_u is 1, wlog, we know that X_v is 0. Hence X_v and X_u are negatively correlated.
- Vertices u and v are not adjacent, but share at least one neighbor. Consider one such common neighbor, w, i.e., $w \in N[u] \cap N[v]$. Knowing u has the smallest hash value in $N[u]$, i.e., $h(w) > h(u)$, implies $h(v) < h(w)$ is likely, and hence v is more likely to have the smallest hash value in $N[v]$. Random variables X_v and X_u are positively correlated. Lemma 7 consolidates this positive correlation.
- Vertices u and v do not fall into the previous two cases: X_v and X_u are independent (as confirmed by Lemma 7).

▶ **Lemma 7.** [⋆] *Let x and y be non-adjacent vertices with $|N(x) \cap N(y)| = k$, $|N(x) \setminus N(y)| = l$, and $|N(y) \setminus N(x)| = r$. If the vertices are mapped to a random permutation, we have* $\Pr[y < N(y) \mid x < N(x)] = \frac{(l+r+2k+2)}{(r+k+1)(l+k+r+2)}$.

[7] Due to space constraints, proofs of all results labeled [⋆] are deferred to the Appendix.

Within a factor $1 \pm \varepsilon$, as shown in Appendix Section B.2, in a permutation generated from a hash function from a ε-min-wise family, Lemma 7 also holds: it relies only the probability of an element being the smallest.

By Lemma 7, $\Pr[y < N(y) \mid x < N(x)] = 1/(r+1) = \Pr[y < N(y)]$ when $k = 0$: x and y have no common neighbor. However, if $k > 0$, then $\frac{(l+r+2k+2)}{(r+k+1)(l+k+r+2)} > \frac{1}{r+k+1}$. So $x < N(x)$ and $y < N(y)$ are positively correlated if and only if x and y share at least one neighbor.

With a Chernoff-type bound unavailable, we bound the variance of $\widehat{\lambda}$, and apply Chebyshev's inequality. It suffices to bound the variance of X as $\mathrm{Var}(\widehat{\lambda}) = \mathrm{Var}(|S|/p) = \mathrm{Var}(X)$.

▶ **Lemma 8.** [⋆] *If $\Delta \leq \varepsilon^2 n/(3(\bar{d}+1)^3)$ and $p \geq 4(\bar{d}+1)/(\varepsilon^2 n)$, then $Var(X) \leq \varepsilon^2 E^2[X]/3$.*

We now have all the tools to prove Theorem 5, Appendix Section B.4 contains the details. By running several instances and returning the median, the success probability of Algorithm 1 can be boosted to $1-\delta$ in $\mathcal{O}(\bar{d}\varepsilon^{-2} \log \delta^{-1} \cdot \log n)$ total space.

3.1 Extension 1: An Independent Set in the Online Streaming model

Algorithm 1 only returns an estimate of the Caro-Wei bound, λ. It can be easily modified to give CARAWAY$_1$, which outputs an actual independent set in the *online streaming* model. Select h uniformly at random from the ε-min-wise hash family $\mathcal{H}$. We initialise an n-bit array to true in the solution space. Upon each arriving edge (u, v), ordered such that $h(u) \leq h(v)$, we irrevocably set the v-th index to false. At any time the set of indices set to true represents an independent set with respect to the edges observed so far. By the analysis above this set is indeed independent, has expected size in $(1 \pm \varepsilon)\lambda$ and is concentrated around its expectation. The working space of our algorithm is merely the space to store the hash function, i.e., $\mathcal{O}(\log \varepsilon^{-1} \cdot \log n)$. Moreover, the update time of our algorithm is the time to compute a hash value, $\mathcal{O}(\log \varepsilon^{-1})$.

▶ **Theorem 9.** [⋆] *For a graph with average degree $\bar{d}$ and max degree $\Delta \leq \varepsilon^2 \, n/(3 \, (\bar{d} \, + \, 1)^3)$, there is an* online streaming *algorithm that, with probability at least $2/3$, reports an independent set of size $1 \pm \varepsilon$ times the Caro-Wei bound. Working space is $\mathcal{O}(\log \varepsilon^{-1} \cdot \log n)$ and update time is $\mathcal{O}(\log \varepsilon^{-1})$.*

3.2 Extension 2: Removing the Constraint on Max Degree

The bound $\frac{\varepsilon^2 \, n}{3 \, (\bar{d}+1)^3}$ on the maximum degree, Δ, is a requirement for Theorem 5. CARAWAY$_2$ avoids this constraint by first removing the heavy hitters (i.e., high-degree vertices) in the graph. In the Caro-Wei Bound $\sum_{v \in V(G)} \frac{1}{1+\deg(v)} \geq \lambda$ high-degree vertices have very little impact on the lower bound of λ. We formalize this observation in Lemma 10. Let H_μ be the set of vertices with degree at least μ and G_{-H_μ} be the induced subgraph on $V \setminus H_\mu$.

▶ **Lemma 10.** [⋆] *For every $\mu \geq (\bar{d}+1)/\sqrt{\varepsilon}$, we have $\lambda(G) \geq \lambda(G_{-H_\mu}) \geq (1-\varepsilon)\lambda(G)$.*

Let R be the vertices to be ignored, which should[8] be those with degree at least $(\bar{d}+1)/\sqrt{\varepsilon}$. Ignoring R, similar to Algorithm 1, Subroutine 2 approximates λ in the remaining graph by sampling vertices and removing some, based on hash values. By Lemma 10, this estimate is at most factor $1-\varepsilon$ away from the Caro-Wei bound of the original graph, G.

We identify the high-degree vertices, R, via the heavy-hitter sketch, with specific parameter settings, as described in Section 3.2.1). Since R is only reported at the end of stream,

[8] Lemma 11 clarifies exactly which vertices are in R.

Subroutine 2 retains all neighbors of vertices in S, deferring the heavy-hitter ignoring and hash-value removal to post-processing. If two passes are allowed, or if R is provided in advance, then these two steps can be done on-the-fly, fitting into the online streaming model. In particular, the condition in line 6 of Algorithm 1 becomes $(u \in S) \wedge (v \notin R) \wedge (h(u) \geq h(v))$.

Subroutine 2 Estimating Caro-Wei Bound without Degree Constraint

Input 1: the average degree, $\bar{d}$, the error rate, ε
Initialization: Uniformly-at-random select $h \in \mathcal{H}$, an ε-min-wise hash family from $[n]$ to $[n^3]$
$p \leftarrow 4(\bar{d}+1)/(\varepsilon^2 n)$
$S \leftarrow pn$ vertices sampled uniformly at random from $[n]$
for $u \in S$ **do**
    $L_u \leftarrow \{\}$
for all $e = (u, v)$ in the stream **do**
    if $u \in S$ **then**
        Add v to L_u
    Perform the same operation on v
    if $\sum_{u \in S}(1 + |L_u|) > 10\bar{d}pn$ **then**
        Abort

Post-processing:
Input 2: a set of ignored vertices, R
for all $v \in R$ **do**
    for all $u \in S$ **do**
        $L_u \leftarrow L_u \setminus \{v\}$
for all $v \in S$ **do**
    for all $u \in L_v$ **do**
        if $h(u) < h(v)$ **then**
            $S \leftarrow S \setminus \{v\}$
return $\widehat{\lambda} = |S|/p$

▶ **Lemma 11.** [⋆] *Suppose R contains all vertices with degree at least $\frac{\varepsilon^2}{3}\frac{n}{(\bar{d}+1)^3}$, but no vertices with degree less than $(\bar{d}+1)/\sqrt{\varepsilon}$. Given $\varepsilon \in (0,1)$, for every graph with average degree $\bar{d}$, Subroutine 2 is an insertion-only algorithm that $(1 \pm \varepsilon)$-approximates the Caro-Wei bound. Its work space is $\mathcal{O}(\bar{d}^2\varepsilon^{-2}\log n)$; it succeeds with probability at least $19/30$.*

3.2.1 Identifying heavy hitters

The success of Subroutine 2 depends on identifying the set of high-degree vertices, R, well. By Theorem 4, running the algorithm of Cormode and Muthukrishnan [13] with $\varepsilon = \tau = \frac{\varepsilon^2}{6(\bar{d}+1)^4}$ and $\delta = n^{-c'}$, for some constant $c' > 1$, we obtain the desired removal set with high probability. This heavy-hitter technique requires $\mathcal{O}(\varepsilon^{-1}\log^2 n)$ space. By running $c' \log n$ (for some constant c') instances of Subroutine 2 concurrently, and returning the maximum result, the failure probability drops to $n^{-c'}$. Importantly, there is only one instance of the heavy-hitter algorithm. The heavy-hitter algorithm is independent of Subroutine 2, but feeds its output R to every instance of Subroutine 2. Via a union bound, the whole process's failure probability is at most $2n^{-c'}$, hence:

▶ **Theorem 12.** *Given $\varepsilon \in (0,1)$, for every graph with average degree $\bar{d}$, there is an algorithm that whp $(1 \pm \varepsilon)$-approximates the Caro-Wei bound in $\mathcal{O}(\bar{d}^2\varepsilon^{-2}\log^2 n)$ space.*

3.3 Vertex-Arrival Random-Order Algorithm

So far, we have relied on a hash function drawn from an ε-min-wise hash family to approximate a random permutation. However, if the stream is vertex-arrival and random-order, the random permutation comes for free. Moreover, we can immediately tell whether the vertex is earliest

in its neighborhood because the *convention* is only to list prior-occurring neighbors. Hence, it suffices to count the number of such vertices, with space further reduced to $\mathcal{O}(\log(\bar{d}\varepsilon^{-1}))$ by incrementing the counter with probability $p = \frac{4(\bar{d}+1)}{\varepsilon^2 n}$. We abort the algorithm if $c > 10pn$, which happens with probability less than $1/10$. We conclude the following theorem:

▶ **Theorem 13.** [⋆] *Given $\varepsilon \in (0,1)$, for every graph with average degree $\bar{d}$ and max degree $\Delta \leq \frac{\varepsilon^2 n}{3(\bar{d}+1)^3}$, there is an insertion-only random-order vertex-arrival algorithm that $(1 \pm \varepsilon)$-approximates λ. It succeeds with probability at least $19/30$ in space $\mathcal{O}(\log(\bar{d}\varepsilon^{-1}))$.*

4 Sublinear-space Streaming Algorithms for Approximating Graph Parameters in Forests

In this section, we describe our sublinear-space streaming approximation algorithms for approximately estimating graph parameters such as the independence number β, domination number γ and the matching number ϕ when the input graph is a forest[9]. Our algorithms work for turnstile streams in the edge-arrival model. For each of these three parameters, we design two algorithms: the first one is a 1-pass algorithm which uses $\log^{\mathcal{O}(1)} n$ space, and the second one is a 2-pass algorithm which uses $\tilde{\mathcal{O}}(\sqrt{n})$ space but has a better approximation ratio than the 1-pass algorithm. The details are in Table 3.

Table 3 Our algorithms for approximately estimating the three graph parameters γ, β, and ϕ in a streaming model when the input graph is a forest. Each algorithm succeeds with high probability and works for the turnstile streams in the edge-arrival model.

Problem	Number of Passes	Approximation Factor	Space	Reference
β	1	$3/2 \cdot (1 \pm \varepsilon)$	$\log^{\mathcal{O}(1)} n$	Theorem 22
β	2	$4/3 \cdot (1 \pm \varepsilon)$	$\tilde{\mathcal{O}}(\sqrt{n})$	Theorem 25
γ	1	$3 \cdot (1 \pm \varepsilon)$	$\log^{\mathcal{O}(1)} n$	Theorem 32
γ	2	$2 \cdot (1 \pm \varepsilon)$	$\tilde{\mathcal{O}}(\sqrt{n})$	Theorem 38
ϕ	1	$2 \cdot (1 \pm \varepsilon)$	$\log^{\mathcal{O}(1)} n$	[6], Theorem 39
ϕ	2	$3/2 \cdot (1 \pm \varepsilon)$	$\tilde{\mathcal{O}}(\sqrt{n})$	Theorem 45

For brevity, we describe in this section our 1-pass and 2-pass algorithms for approximately estimating the independence number, β. The algorithms for the domination number γ and matching number ϕ are deferred to Appendix C.8 and C.9, respectively. Our sublinear-space streaming algorithms for approximating β comprise two main steps:

- (Section 4.1) Obtain bounds on β in terms of other structural graph quantities such as number of leaves and number of support vertices (i.e., vertices adjacent to leaves),
- (Section 4.2) Estimate the aforementioned two structural graph quantities in the streaming model in sublinear-space.

4.1 Structural Bounds on Independence Number

Leaves and Non-leaf Vertices: The number of non-leaf vertices has been used in 2-approximating the matching number [6, 18] and 3-approximating the domination number in forests [17, 32, 35]. Non-leaf vertices are helpful because one can always reason that some

[9] Unless otherwise stated, we assume that the input forest has no isolated vertices.

max dominating set must be a subset of the non-leaf vertices. Similarly, there exists a max independent set that includes all leaves. For a forest, F, let Deg_1 be the set of leaves, c be the number of tree components of F, and $Deg_{\geq 2}$ be the set of non-leaf vertices. First we show how to upper and lower bound β using only the quantities Deg_1 and c.

▶ **Theorem 14.** [⋆] *For every forest,* $\max\left\{\frac{1}{2}n\,,\,|Deg_1|-c\right\} \leq \beta \leq \frac{1}{2}\big(n+|Deg_1|\big)\,.$

Support Vertices: We now demonstrate the power of support vertices, i.e., vertices adjacent to leaves. For all three graph problems, the optimal solution can be found in linear space via standard dynamic programming once we determine whether each leaf is in the solution. Due to our (sublinear) space constraints, we cannot apply dynamic programming, but it turns out going just one-level above the leaves, i.e., counting the number of support vertices, is sufficient to improve the approximations. Let *Supp* be the set of support vertices in forest F.

▶ **Theorem 15.** [⋆] *For every forest,* $\frac{1}{2}\big(n+|Deg_1|-|Supp|\big) \leq \beta \leq \frac{2}{3}\big(n+|Deg_1|-|Supp|\big)\,.$

Furthermore, combining Theorem 15 with Theorem 14, we have:

▶ **Corollary 16.** [⋆] *If* $|Supp| \leq \frac{1}{2}\,|Deg_{\geq 2}|$*, then* $\frac{3}{8}\big(n+|Deg_1|\big) \leq \beta \leq \frac{1}{2}\big(n+|Deg_1|\big)\,.$

In the next section, we will obtain streaming algorithms for estimating β by designing streaming algorithms for estimating the structural graph quantities introduced in this section.

4.2 One-pass and two-pass Streaming Algorithms for estimating β

In this setcion, we show how to estimate the number of leaves and support vertices within factor $(1\pm\varepsilon)$ in graph streams, with small space usage.

Our algorithms rely on the concept of the degree vector $\mathcal{D}(G)$. The i^{th} coordinate holds the degree of vertex i. Storing this vector would require $\Theta(n\log n)$ bits, exceeding our working-space bound. Here, we rely on sparse-recovery techniques to estimate quantities that are functions of $\mathcal{D}(G)$, supporting fast stream updates. We start with one-pass algorithms.

4.2.1 Estimating β in one pass

Non-leaf Vertices: The quantity $|Deg_{\geq 2}(G)|$ is exactly the number of coordinates in $\mathcal{D}(G)$ with value at least 2. Since we assume no isolated vertices in G, by subtracting the all-ones vector $\{1\}^n$ from $\mathcal{D}(G)$, i.e., initializing $\mathcal{D}'(G) \leftarrow \mathcal{D}(G)+\{-1\}^n$, the number of non-zero entries, $\|\mathcal{D}'(G)\|_0$, is exactly $|Deg_{\geq 2}(G)|$. Hence, an algorithm that $(1\pm\varepsilon)$-approximates $\|\mathcal{D}'(G)\|_0$ gives an $(1\pm\varepsilon)$ approximation to $|Deg_{\geq 2}(G)|$. Applying Theorem 2, we may $(1\pm\varepsilon)$-approximate $|Deg_{\geq 2}(G)|$ in $\log^{\mathcal{O}(1)} n$ space with constant success probability. Since no existing L_0 sketch supports initialization with $\{-1\}^n$, degree decrements are post-processed. Running $\mathcal{O}(\log\delta^{-1})$ independent instances and returning the median, the success probability exceeds $(1-\delta)$, hence:

▶ **Lemma 17.** *For every* $\varepsilon,\delta\in(0,1)$ *and graph* G*,* $|Deg_{\geq 2}|$ *can be* $(1\pm\varepsilon)$*-approximated in a turnstile stream with probability* $(1-\delta)$ *and in* $\mathcal{O}(\log^{\mathcal{O}(1)} n\cdot\log\delta^{-1})$ *space.*

Lemma 17 allows us to $(2\pm\varepsilon)$-approximate the matching number (see Theorem 18) and $(3\pm\varepsilon)$-approximate the domination number (see Theorem 19) in streaming forests using only polylog space. Sections C.8 and C.9 in the Appendix have full details.

▶ **Theorem 18.** [⋆] *[6, 18] For every forest* F*,* $\max\left\{\,c\,,\,\frac{1}{2}(|Deg_{\geq 2}|+c)\,\right\} \leq \phi \leq |Deg_{\geq 2}|+c\,.$

▶ **Theorem 19.** [⋆] *For every forest F, $\frac{1}{3}|Deg_{\geq 2}| \leq \gamma \leq |Deg_{\geq 2}| + c$.*

Leaves: To estimate β, we could estimate $|Deg_1(F)|$ in a forest F with no isolated vertices by the expression $n - |Deg_{\geq 2}(F)|$. But if $|Deg_1(F)|$ is $o(|Deg_{\geq 2}(F)|)$, the error relative to $|Deg_1(F)|$ could be very large. We turn elsewhere.

▶ **Lemma 20.** [⋆] *For every forest F, we have $|Deg_1(F)| = 2c(F) + \sum_{i=3}^{\Delta}(i-2)\cdot|Deg_i(T)|$.*

Now, subtracting 2 from each coordinate of degree vector $\mathcal{D}(F)$ (i.e., post-processing $\mathcal{D}''$ with $\{-2\}^n$), we have $\|\mathcal{D}''(F)\|_1 = |Deg_1(F)| + \sum_{i=3}^{\Delta}(i-2)\cdot|Deg_i(F)|$. Folding in Lemma 20, we have $|Deg_1(F)| = \|\mathcal{D}''(F)\|_1/2 + c(F)$. By Theorem 1, $|Deg_1(F)|$ is $(1\pm\varepsilon)$-approximated with constant success probability in $\log^{\mathcal{O}(1)} n$ space, and hence:

▶ **Lemma 21.** *For every $\varepsilon, \delta \in (0,1)$ and forest F with n vertices, $|Deg_1(F)|$ can be $(1\pm\varepsilon)$-approximated in a turnstile stream with probability $(1-\delta)$ in $\mathcal{O}(\log^{\mathcal{O}(1)} n \cdot \log \delta^{-1})$ space.*

Combining Lemma 21 with Theorem 14, we conclude:

▶ **Theorem 22.** *On forests, independence number β can be $3/2\cdot(1\pm\varepsilon)$-approximated with probability $(1-\delta)$ on turnstile streams using $\mathcal{O}(\log^{\mathcal{O}(1)} n \cdot \log \delta^{-1})$ space.*

The above theorem shows that we can approximate the independence number β in one-pass to within factor 3/2 in $\log^{\mathcal{O}(1)} n$ space. In the next subsection, we are able to improve the approximation factor by using an extra pass and invoking the notion of support vertices.

4.2.2 Estimating β in two passes

Support Vertices: Identifying a support vertex requires knowing the degree of all its neighbours. The degree-counting approach above does not suffice to estimate the number of support vertices. We first show in Subroutine 3, given a forest F in a turnstile stream, how to output an $(1\pm\varepsilon)$-estimate of $|Supp(F)|$ when $|Supp(F)|$ is at least threshold K_1. Then we demonstrate that, when $|Supp(F)|$ is small, we can still approximate the graph parameter β with the same approximation ratio, using either one of the following methods:

- When the set of non-leaf vertices is also small, we can compute both $|Deg_{\geq 2}|$ and $|Supp(F)|$ exactly (Subroutine 4).
- When the set of non-leaf vertices is large, as suggested by Corollary 16, we can exclude $|Supp(F)|$ without sacrificing the approximation ratio.

Similar to Cormode et al. [12] and to Jayaram and Woodruff [26], we assume the number of deletions is $\mathcal{O}(n)$. If an arbitrary number of deletions were allowed, the number of insertions might be arbitrarily large, breaking our analyses of Subroutine 3 and Subroutine 4.

▶ **Lemma 23.** [⋆] *Given size threshold K_1, constant c_1, and error rate $\varepsilon_1 \in (0,1)$, Subroutine 3 is a turnstile-stream algorithm that uses $\widetilde{\mathcal{O}}(\frac{n}{\varepsilon_1^2 K_1})$ space and has the following approximation properties. When $|Supp(F)| \geq K_1$, it $(1\pm\varepsilon_1)$-approximates $|Supp(F)|$ with probability at least $1 - 3e^{-\frac{c_1}{3}}$. When $|Supp(F)| < K_1$, it returns an estimate at most $2K_1$ with probability at least $1 - 2e^{-\frac{c_1}{3}}$.*

Again, letting $\mathcal{D}'(F) \leftarrow \mathcal{D}(F) + \{-1\}^n$, when $|Deg_{\geq 2}(F)|$ is small, sparse recovery on $\mathcal{D}'(F)$ recovers all vertices in $Deg_{\geq 2}(F)$, with probability $1-\delta$, but no leaves. In a second pass, we verify whether each recovered vertex is a support vertex thus. For a length-two path (a P_2, i.e., an isolated edge), neither endpoint is recovered. Except for a P_2, if some neighbor of vertex u is not recovered, then u is a support vertex. To verify whether $Deg_{\geq 2}(F)$ is recovered, we sum degrees, knowing all leaves are unrecovered: details are in Subroutine 4.

Subroutine 3 Estimating $|Supp(F)|$ for a forest F

Input: A size threshold K_1, a constant c_1, and error rate $\varepsilon_1 \in (0,1)$
Initialization: $I \leftarrow \frac{c_1 n}{\varepsilon_1^2 K_1}$ vertices sampled uniformly at random from $[n]$
for $v \in I$ **do**
    $l(v) \leftarrow \{\}$
$m, t \leftarrow 0$
First Pass:
for all $(e = (u, v), i)$ in the stream, with $i \in \pm 1$ **do**
    $m \leftarrow m + i$
    if $u \in I$ **then**
        Toggle v's presence in $l(u)$
        $t \leftarrow t + i$
    Perform the same operation on v
    Abort if $t \geq \frac{2m}{n} \frac{c_1 n}{\varepsilon_1^2 K_1} e^{\frac{c_1}{3}}$
Second Pass:
Count the degree of every vertex in $I \cup \left(\bigcup_{w \in I} l(w)\right)$
$C \leftarrow \{u \mid u \in I \text{ and there is some leaf } v \in l(u)\}$
return $\widehat{|Supp(F)|} \leftarrow |C| \times \frac{\varepsilon_1^2 K_1}{c_1}$

Subroutine 4 Estimating $|Supp(F)|$ when $|Deg_{\geq 2}(F)| \leq K_2$

Input: A size threshold K_2, a large constant c_2
Initialization: Sparse recovery sketch $\mathcal{L}$ in [11] with $k = K_2$ and $\delta = 1/c_2$
First Pass:
for all $(e = (u, v), i)$ in the stream, $i \in \pm 1$ **do**
    Update $\mathcal{L}$ with (e, i)
Decrement all coordinates in $\mathcal{L}$ by 1
Denote the result of sparse recovery from $\mathcal{L}$ by $\mathcal{R}$
$p, m \leftarrow 0$
for $v \in \mathcal{R}$ **do**
    $d_v, l_v \leftarrow 0$ ▷ Degree and leaf counters
Second Pass:
for all $(e = (u, v), i)$ in the stream, $i \in \pm 1$ **do**
    $m \leftarrow m + i$; $d_u \leftarrow d_u + i$; $d_v \leftarrow d_v + i$
    if $u \notin \mathcal{R}$ and $v \notin \mathcal{R}$ **then**
        $p \leftarrow p + i$
    if $u \in \mathcal{R}$ and $v \notin \mathcal{R}$ **then**
        $l_u \leftarrow l_u + i$
    if $v \in \mathcal{R}$ and $u \notin \mathcal{R}$ **then**
        $l_v \leftarrow l_v + i$
if $2m \neq n - |\mathcal{R}| + \sum_{v \in \mathcal{R}} d_v$ **then**
    Return *FAIL*
else
    $\widehat{|Supp(F)|} \leftarrow |\{u \mid u \in \mathcal{R} \text{ and } l(u) = 1\}| + 2p$
    $\widehat{|Deg_{\geq 2}(F)|} \leftarrow |\mathcal{R}|$
    Return $\widehat{|Supp(F)|}$ and $\widehat{|Deg_{\geq 2}|}$

▶ **Lemma 24.** [⋆] *Given size threshold K_2, constant c_2, and a forest F as a turnstile stream, Subroutine 4 in two passes outputs exact estimates for both $|Supp(F)|$ and $|Deg_{\geq 2}(F)|$ in $\widetilde{\mathcal{O}}(K_2)$ space. When $|Deg_{\geq 2}(F)| \leq K_2$, the failure probability is at most $1/c_2$.*

Finalizing the Algorithms: In our two-pass Algorithm 5 for estimating the independence number in a forest, we run Subroutine 3 and Subroutine 4 concurrently, returning the minimum of $(3(n + |Deg_1(F)|))/8$ and $(n + |Deg_1(F)| - |Supp(F)|)/2$.

▶ **Theorem 25.** [⋆] *For every $\varepsilon, \delta \in (0, 1)$ and forest F (with no isolated vertices), Algorithm 5 estimates the independence number $\beta(F)$ in a turnstile stream within factor $4/3 \cdot (1 \pm \varepsilon)$ with probability $(1 - \delta)$ in $\widetilde{\mathcal{O}}(\sqrt{n})$ space and two passes.*

Algorithm 5 Estimating $\beta(F)$ in forest F

Input: error rate ε, fail rate δ
Initialization: $K_1 \leftarrow \sqrt{n}$, $K_2 \leftarrow 8\sqrt{n}$, $c_1 \leftarrow 3\ln(6\delta^{-1})$, $c_2 \leftarrow \delta^{-1}$, $\varepsilon_1 \leftarrow \varepsilon/2$
Run the following three subroutines concurrently:
$\widehat{|Deg_1|} \leftarrow$ Lemma 21 algorithm with $\varepsilon/2$ and $\delta/2$
$\widehat{|Supp(F)|} \leftarrow$ Subroutine 3 with K_1, c_1, and ε_1
$(\widehat{|Supp(F)|}', \widehat{|Deg_{\geq 2}|}') \leftarrow$ Subroutine 4 with K_2 and c_2
if Subroutine 4 returns *FAIL* **then**
    Return $\widehat{\beta} \leftarrow \min\{\frac{3(n+\widehat{|Deg_1|})}{8}, \frac{n+\widehat{|Deg_1|}-\widehat{|Supp(F)|}}{2}\}$
else
    $\widehat{|Deg_1|}' \leftarrow n - \widehat{|Deg_{\geq 2}|}'$
    Return $\widehat{\beta} \leftarrow \min\{\frac{3(n+\widehat{|Deg_1|}')}{8}, \frac{n+\widehat{|Deg_1|}'-\widehat{|Supp(F)|}'}{2}\}$

5 Conclusion & Open Questions

In this paper, we introduced Caraway, an algorithm for estimating the Caro-Wei bound. It improves the result of Halldórsson et al. [22] to match the state-of-the-art of Cormode et al. [10] while being able to report a solution in the online streaming model. Having a dedicated, write-only *solution space* supports computing $\mathcal{O}(n)$-size solutions over data streams. This is a fruitful direction for future research.

We have also used the notion of *support vertices* from structural graph theory [15] in the streaming setting. This allowed us to show the independence number, dominating number and matching number can be approximated using sketches to estimate the number of leaves, non-leaves and support vertices. Our algorithms make progress towards meeting the lower bound of Esfandiari et al. [18]. Including the number of support vertices gives a substantial improvement in the bounds that can be achieved; we believe this concept of support vertices can be applied to other problems of computing graph properties of forests under memory restrictions.

It is an interesting direction to see whether some of our results & techniques could also be adapted to other computational models such as distributed algorithms.

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A Supplementary Preliminaries

A.1 Extension of Chernoff-Hoeffding inequality

Panconesi and Srinivasan [40] extended the Chernoff-Hoeffding inequality to negatively correlated Boolean random variables as follows:

▶ **Theorem 26.** *[40] For r negatively correlated Boolean random variables $X_1, \ldots, X_r$, let $X = \sum_{i=1}^{r} X_i$, $\mu = E[X]$ and $0 < \delta < 1$, then*

$$\Pr\left[|X - \mu| \geq \delta\mu\right] \leq 2e^{-\frac{\mu\delta^2}{2}}.$$

B Omitted Proofs from Section 3

B.1 Proof of Lemma 6

Proof. Let binary random variables X_v denote whether vertex v is sampled in S and not removed during the stream. That is, $X_v = 1$ if $v \in S$ at the end of the stream. Let $X = \sum_{v \in V(G)} X_v$, clearly, $X = |S|$ at the end. It is not hard to see that $X_v = 1$ if and only if v is sampled and v has the smallest hash value in $N[v]$. According to the property of ε-min-wise hash families, v has the smallest hash value in $N[v]$ with probability $\frac{1\pm\varepsilon}{|N[v]|}$. By the construction of algorithm, v is sampled with probability p. Since these two events are independent, the probability that $X_v = 1$ is $(1 \pm \varepsilon)\frac{p}{d(v)+1}$. Hence, by the linearity of expectation, we have

$$E[|S|] = \sum_{v \in V(G)} E[X_v] = p \sum_{v \in V(G)} \frac{1 \pm \varepsilon}{d(v)+1} = (1 \pm \varepsilon)\, p\, \lambda$$

Hence, $E[\widehat{\lambda}] = E[\frac{|S|}{p}] = (1 \pm \varepsilon)\, \lambda$.

◀

B.2 Proof of Lemma 7

Proof. Consider $n = l + k + r + 2$ vertices x, y, $w_1, \ldots, w_l$, $u_1, \ldots, u_k$, and $v_1, \ldots, v_r$. For simplicity, we use W, U, and V to denote $\{w_1, \ldots, w_l\}$, $\{u_1, \ldots, u_k\}$, and $\{v_1, \ldots, v_r\}$. Note that x has degree $l + k$, y has degree $k + r$, x and y are not adjacent and they share k neighbors, i.e., $N(x) = \{W \cup U\}$ and $N(y) = \{U \cup V\}$.

We use $x < N(x)$ if x has the smallest order among $N[x]$ in the permutation. Note that since we assume uniform random permutation, $\Pr[x < N(x)] = 1/(|N(x)|+1) = 1/(l+k+1)$. Consider the event $X < N(X)$ and $Y < N(Y)$, we have two cases:

- x, followed by zero or more W, y, followed by a mixture of the rest of W, V, and U.
- y, followed by zero or more V, x, followed by a mixture of W, the rest of V, and U.

By the uniform random permutation, the *first* event occurs with probability

$$\frac{1}{l+k+r+2} \cdot \frac{1}{k+r+1},$$

where the first factor is the probability of $x < \{y\} \cup W \cup V \cup U$, and the second factor is the probability of $y < V \cup U$, regardless of how it relates with vertices in W. Note these two events are independent of each other, since knowing x being the smallest across y, W, U, V does not reveal anything about the ordering inside of y, W, U, V. By similar argument, the *second* event occurs with probability

$$\frac{1}{l+k+r+2} \cdot \frac{1}{l+k+1}$$

Summing over the two cases, we have

$$\begin{aligned}
&\Pr[y < N(y) \wedge x < N(x)] \\
&= \frac{1}{l+k+r+2} \cdot \frac{1}{k+r+1} + \frac{1}{l+k+r+2} \cdot \frac{1}{l+k+1} \\
&= \frac{l+2k+r+2}{(l+k+r+2)(k+r+1)(l+k+1)}
\end{aligned}$$

Finally, the conditional probability can be calculated as:

$$\begin{aligned}
\Pr[y < N(y) \mid x < N(x)] &= \frac{\Pr[y < N(y) \wedge x < N(x)]}{\Pr[x < N(x)]} \\
&= \frac{(l+r+2k+2)}{(l+k+r+2)(k+r+1)}
\end{aligned} \tag{1}$$

◀

B.2.1 The ε-min-wise property:

In Lemma 7, we assume uniform random permutation, but the same claim still holds with ε-min-wise property, since we only consider the probability of an element being the smallest. The proof of ε-min-wise property has similar structure with the proof of Lemma 7, except we need to take $1 \pm \varepsilon$ into account every time we calculate the probability of an element being smallest. We also observe that the statement "knowing x being the smallest across y, W, U, V does not reveal anything about the ordering inside of y, W, U, V" requires a small amount of extra independence in the $\mathcal{O}(\log \varepsilon^{-1})$-wise hash function. This has no effect on the asymptotic space bound for the hash function.

Therefore, we have $\Pr[x < N(x)] = (1 \pm \epsilon)/(l+k+1)$ and

$$\begin{aligned}
&\Pr[y < N(y) \wedge x < N(x)] \\
&= \frac{1 \pm \varepsilon}{l+k+r+2} \cdot \frac{1 \pm \varepsilon}{k+r+1} + \frac{1 \pm \varepsilon}{l+k+r+2} \cdot \frac{1 \pm \varepsilon}{l+k+1} \\
&= (1 \pm \varepsilon)^2 \frac{l+2k+r+2}{(l+k+r+2)(k+r+1)(l+k+1)}
\end{aligned}$$

And finally

$$\begin{aligned}
\Pr[y < N(y) \mid x < N(x)] &= \frac{\Pr[y < N(y) \wedge x < N(x)]}{\Pr[x < N(x)]} \\
&= (1 \pm \varepsilon) \frac{(l+r+2k+2)}{(l+k+r+2)(k+r+1)}
\end{aligned} \tag{2}$$

B.3 Proof of Lemma 8

Proof. We persist with the binary random variables from the proof of Lemma 6. In the proof of Lemma 6, we show that $\Pr[X_v = 1] = \frac{(1\pm\varepsilon)p}{d(v)+1}$. For simplicity, we ignore the $(1 \pm \varepsilon)$ factor in the following analysis, that is, $\Pr[X_v = 1] = \frac{p}{d(v)+1}$; the same analysis holds without removing $(1 \pm \varepsilon)$. For each X_v, we have

$$\mathrm{Var}(X_v) = E[X_v^2] - E^2[X_v] < \frac{p}{d(v)+1}(1 - \frac{p}{d(v)+1}) < \frac{p}{d(v)+1}$$

For non-adjacent vertices i and j share $k \geq 0$ common neighbors, the covariance between X_i and X_j is

$$\begin{aligned}
&\mathrm{Cov}(X_i, X_j \mid (i,j) \notin E)\\
&= E[X_iX_j] - E[X_i]E[X_j]\\
&= \Pr[X_i = 1 \wedge X_j = 1] - \Pr[X_i = 1]\Pr[X_j = 1]\\
&= \Pr[X_i = 1 \mid X_j = 1]\Pr[X_j = 1] - \Pr[X_i = 1]\Pr[X_j = 1]\\
&= (\frac{p(d(i)+d(j)+2)}{(d(i)+1)(d(i)+d(j)-l+2)} - \frac{p}{d(i)+1})\frac{p}{d(j)+1}\\
&= \frac{p^2 l}{(d(i)+d(j)-l+2)(d(i)+1)(d(j)+1)},
\end{aligned} \tag{3}$$

where the second-last equality arises from Equation (1), where we set $d(i) = l + k$ and $d(j) = k + r$. It is not hard to see that Equation (3) is a constraint function subject to $d(i) \geq 1$, $d(j) \geq 1$, and $k \leq \min\{d(i), d(j)\}$. This function has global maximum at $\frac{p^2 k}{(k+1)^2(k+2)}$ when $k = d(i) = d(j)$ for different values of $d(i)$ and $d(j)$, which is maximized at $\frac{p^2}{12}$ when $k = 1$. As indicated in Equation (2), when considering ε-min-wise property instead of uniform random permutation, the global maximum is $\frac{(1\pm\epsilon)p^2}{12} \leq \frac{p^2}{8}$ if $\varepsilon < 1/2$.

And for adjacent vertices i and j, X_i and X_j are strongly negatively correlated, as $X_i = 1$ implies $X_j = 0$. Hence,

$$\mathrm{Cov}(X_i, X_j \mid (i,j) \in E) = E[X_iX_j] - E[X_i]E[X_j] < 0$$

Let P be the number of vertex pairs that share at least one common neighbor. Assume the vertices are labeled from 1 to n, by variance equation for the sum of correlated variables, we have (for simplicity, we sometimes abbreviate $1 \leq i < j \leq n$ as $i < j$),

$$\begin{aligned}
\mathrm{Var}(X) &= \sum_{i=1}^{n} \mathrm{Var}(X_i) + 2\sum_{i<j} \mathrm{Cov}(X_i, X_j)\\
&< \sum_{i=1}^{n} \mathrm{Var}(X_i) + 2\sum_{i<j\wedge(i,j)\notin E} \mathrm{Cov}(X_i, X_j)\\
&< \sum_{i=1}^{n} \frac{p}{d(i)+1} + 2\sum_{i<j\wedge(i,j)\notin E} \frac{p^2 k}{(k+1)^2(k+2)}\\
&\leq p\,\lambda + 2p^2\frac{P}{8}
\end{aligned} \tag{4}$$

Note that each vertex v introduces at most $\frac{d_v(d_v-1)}{2} \leq \frac{d_v^2}{2}$ new pairs of vertices that share a common neighbor, i.e., v. Thus, $P \leq \frac{1}{2}\sum_{v\in V(G)} d_v^2$. Since $\Delta \leq \frac{\varepsilon^2 n}{3(\bar{d}+1)^3}$, we have

$$\begin{aligned}\sum_{v\in V(G)} d_v^2 \leq \Delta^2\frac{\bar{d}n}{\Delta} + (n - \frac{\bar{d}n}{\Delta}) &< (\Delta\bar{d}+1)n \\ &\leq (\frac{\varepsilon^2 n\bar{d}}{3(\bar{d}+1)^3}+1)n \\ &< \frac{\varepsilon^2 n^2}{3(\bar{d}+1)^2} + n \\ &\leq \frac{\varepsilon^2 n^2}{2(\bar{d}+1)^2}\,,\end{aligned} \tag{5}$$

where the last inequality holds when $n \geq \frac{\sqrt{6}(\bar{d}+1)}{\varepsilon}$. And the first inequality is because given the maximum degree Δ and a sufficiently large n, the term $\sum_{v\in V(G)} d_v^2$ is maximized when the degree distribution is highly biased. That is, when there are around $\frac{\bar{d}n}{\Delta}$ vertices hitting the maximum degree Δ, while the rest of vertices are having degree 1 (as the graph is assumed to be connected).

Finally, combining Equation (4) and (5), we have

$$\begin{aligned}\mathrm{Var}(X) \leq p\,\lambda + 2p^2\frac{P}{8} &\leq p\,\lambda + p^2\frac{\varepsilon^2 n^2}{16(\bar{d}+1)^2} \\ &\leq p\,\lambda + p^2\frac{\varepsilon^2\lambda^2}{16} \\ &= E[X] + \frac{\varepsilon^2 E^2[X]}{16} \\ &\leq \frac{\varepsilon^2 E^2[X]}{3}\end{aligned}$$

where the third inequality holds because the Turán Bound is at most the Caro-Wei Bound, i.e., $\frac{n}{\bar{d}+1} \leq \sum_{v\in V(G)} \frac{1}{d(v)+1} = \lambda$ (see [4] for a proof). The inequality in the second-last line arises from $E[X] = p\lambda$. And the last inequality holds if $E[X] \geq \frac{16}{7\varepsilon^2}$, since we have assumed that $p \geq \frac{4(\bar{d}+1)}{\varepsilon^2 n}$ and $\lambda \geq \frac{n}{\bar{d}+1}$, $E[X] = p\lambda \geq \frac{4}{\varepsilon^2}$, thus the inequality follows. ◀

B.4 Proof of Theorem 5

Proof. From Lemma 6 and Lemma 8, the expectation of the returned result, $\widehat{\lambda}$, is $(1\pm\varepsilon)\lambda$, and $\mathrm{Var}(X) \leq \frac{\varepsilon^2 E^2[X]}{3}$. Hence, by Chebyshev's inequality,

$$\Pr\left[|\widehat{\lambda}-\lambda| \geq 3\varepsilon\lambda\right] \leq \Pr\left[|\widehat{\lambda}-E[\widehat{\lambda}]| \geq \varepsilon E[\widehat{\lambda}]\right] = \Pr\left[|X-E[X]| \geq \varepsilon E[X]\right] \leq \frac{1}{3}\,,$$

where the first inequality holds because $\varepsilon < 1$ so that $(1+\varepsilon)^2 < (1+3\varepsilon)$ and $(1-\varepsilon)^2 > (1-3\varepsilon)$. Therefore, the algorithm succeeds with probability at least $2/3$.

The algorithm stores an ε-min-wise hash function and $\mathcal{O}(\bar{d}\varepsilon^{-2})$ vertices. The hash function takes $\mathcal{O}(\log\varepsilon^{-1}\cdot\log n)$ bits to store, and each vertex v takes $\mathcal{O}(\log n)$ bits to store. Therefore, the total space usage is $\mathcal{O}((\bar{d}\varepsilon^{-2}+\log\varepsilon^{-1})\log n) = \mathcal{O}(\bar{d}\varepsilon^{-2}\log n)$. ◀

B.5 Proof of Theorem 9

Proof. Let a binary random variable X_v denote whether v is not removed by our algorithm. Since v is removed if and only if it does not have the smallest hash value across $N[v]$, $\Pr[X_v = 1] = \frac{1\pm\varepsilon}{d(v)+1}$. Let $X = \sum_{v\in V(G)} X_v$. By the construction of the algorithm, X is exactly the size of the final solution. And

$$E[X] = \sum_{v\in V(G)} \Pr[X_v = 1] = (1\pm\varepsilon) \sum_{v\in V(G)} \frac{1}{d(v)+1} = (1\pm\varepsilon)\,\lambda$$

Following the same proof as Lemma 8 (note: since we are not sampling, the sampling probability p in Lemma 8 can be regarded as 1), we can bound the variance of X as $\mathrm{Var}(X) \le \frac{\varepsilon^2 E^2[X]}{3}$. Applying Chebyshev's inequality, we have

$$\Pr\left[|\widehat{\lambda} - \lambda| \ge 3\varepsilon\lambda\right] \le \frac{1}{3},$$

Hence, the algorithm fails with probability at most $1/3$.

The algorithm only uses a randomly drawn ε-min-wise hash function, which can be stored using $\mathcal{O}(\log\varepsilon^{-1}\log n)$ bits and has calculation time $\mathcal{O}(\log\varepsilon^{-1})$. Hence, the space usage and the update time our algorithm is $\mathcal{O}(\log\varepsilon^{-1}\log n)$ and $\mathcal{O}(\log\varepsilon^{-1})$.

◀

The success probability can be boosted to $1-\delta$ by running $\log\delta^{-1}$ independent instances simultaneously and return the maximum result. By the union bound, the probability that all instances fail is at most $(\frac{1}{3})^{\log\delta^{-1}} = \delta$. After boosting, the new algorithm uses space $\mathcal{O}(\log\varepsilon^{-1}\log\delta^{-1}\log n)$ and has update time $\mathcal{O}(\log\varepsilon^{-1}\log\delta^{-1})$.

B.6 Proof of Lemma 10

Proof. The upper bound, $\lambda(G) \ge \lambda(G_{-Deg_{\ge\mu}})$, clearly holds as removing vertices does not increase λ.

Given the average degree $\bar{d}$, by the Pigeonhole Principle, there are at most $\frac{\bar{d}n}{\mu}$ vertices with degree at least μ. Hence, we have

$$\begin{aligned}\lambda(G_{-Deg_{\ge\mu}}) > \lambda(G) - \sum_{v\in Deg_{\ge\mu}} \frac{1}{d(v)+1} &\ge \lambda(G) - \frac{1}{\mu+1}\frac{\bar{d}n}{\mu} \\ &> \lambda(G) - \varepsilon\frac{\bar{d}n}{(\bar{d}+1)^2} \\ &\ge \lambda(G) - \varepsilon\lambda(G),\end{aligned}$$

where the last inequality holds as the Turán Bound is at most the Caro-Wei Bound, i.e., $\frac{n}{d+1} \le \sum \frac{1}{d(v)+1} = \lambda$ (see [4] for a proof).

◀

B.7 Proof of Lemma 11

Proof. By the construction of Subroutine 2, vertices in R are removed from the neighborhood list of sampled vertices (line 15 – 17). Hence, by the condition on line 20, each sampled vertex

is retained in S if and only if it has the smallest hash value across its neighbors in the subgraph after removing R. Let X_v be the binary indicator of v being sampled and retained in S. Since the sampling procedure and the hash function are independent, $\Pr[X_v = 1] = (1 \pm \varepsilon)\frac{p}{d'(v)+1}$, where $d'(v)$ is the degree of v after removing R. Let $X = \sum_{v \in V(G)} X_v$. Clearly, $X = |S|$ at the end of stream. Let λ' be the Caro-Wei bound of the subgraph after removing R, by the linearity expectation, the expectation of $\widehat{\lambda'}$ is

$$E[\widehat{\lambda'}] = E[|S|]/p = \frac{1}{p} \sum_{v \in V(G)} E[X_v] = \sum_{v \in V(G)} \frac{1 \pm \varepsilon}{d'(v)+1} = (1 \pm \varepsilon)\lambda' ,$$

The variance of $\widehat{\lambda'}$ is the same as the variance of X, as p is fixed. By Lemma 8, condition on R contains all vertices with degree at least $\frac{\varepsilon^2 \ n}{3 \ (\bar{d}+1)^3}$, we have $\mathrm{Var}(X) \leq \frac{\varepsilon^2 E^2[X]}{3}$. Hence,

$$\Pr\left[|\widehat{\lambda'} - \lambda'| \geq 3\varepsilon\lambda'\right] \leq \Pr\left[|X - E[X]| \geq \varepsilon E[X]\right] \leq \frac{\mathrm{Var}(X)}{\varepsilon^2 E^2[X]} \leq \frac{1}{3}$$

By Lemma 10, condition on R contains no vertices with degree smaller than $\frac{\bar{d}(\bar{d}+1)}{\varepsilon}$, $\lambda' \geq (1-\varepsilon)\lambda$. Also, the algorithm aborts if the size of sampled vertices and their neighborhood lists are too large (line 11 – 12). Since the average degree is $\bar{d}$, in expectation, $\bar{d}pn$ vertices are stored by our algorithm. By Markov's inequality, the probability that the actual number of stored vertices is 10 times greater than its expectation, $\bar{d}pn$, is at most $1/10$. Applying the union bound, the probability that this algorithm fails or aborts is at most $11/30$.

The space usage of this algorithm is space to store the hash function, plus the space to store sampled vertices and their neighbors. Because of our constraints on line 11, there are at most $\mathcal{O}(\bar{d}pn) = \mathcal{O}(\bar{d}^2\varepsilon^{-2})$ vertices being stored, where each vertex takes $\mathcal{O}(\log n)$ bits to store. The hash function takes $\mathcal{O}(\log \varepsilon^{-1} \log n)$ bits to store. Hence, the total space usage is $\mathcal{O}(\bar{d}^2\varepsilon^{-2} \log n)$.

◀

B.8 Proof of Theorem 13

Proof. Firstly, note that $\widehat{\lambda}$ is an unbiased estimator of λ, as each vertex arrives uniformly at random. Let binary random variable X_v denote whether c is incremented when v arrives. That is, $X_v = 1$ if the neighborhood list of v is empty, and the counter is increased. These two events are independent, the former event happens with probability $\frac{1}{d(v)+1}$, while the latter event happens with probability p. Let $X = \sum_{v \in V(G)} X_v$. By the linearity of expectation, we have

$$E[\widehat{\lambda}] = E[c]/p = E[X]/p = \frac{1}{p} \sum_{v \in V(G)} E[X_v] = \sum_{v \in V(G)} \frac{1}{d(v)+1} = \lambda$$

By Lemma 8, the variance of X, $\mathrm{Var}(X) \leq \frac{\varepsilon^2 E^2[X]}{3}$. Applying Chebyshev's inequality, we have

$$\Pr\left[|\widehat{\lambda} - \lambda| \geq \varepsilon\lambda\right] = \Pr\left[|X - E[X]| \geq \varepsilon E[X]\right] \leq \frac{\mathrm{Var}(X)}{\varepsilon^2 E^2[X]} \leq \frac{1}{3} ,$$

The algorithm might fail if the counter, c, becomes too large, i.e., $c > 10pn$. By Markov's inequality,

$$\Pr[X \geq 10pn] \leq \frac{E[X]}{10pn} = \frac{\lambda}{10n} \leq \frac{1}{10}$$

Therefore, applying the union bound, the algorithm succeeds with probability at least $1 - 1/3 - 1/10 = 19/30$.

The space usage of the algorithm is the size of the counter c. By the construction of the algorithm, c is incremented to at most $10pn$. Hence, it can be stored in $\mathcal{O}(\log(pn)) = \mathcal{O}(\log(\bar{d}\varepsilon^{-2}))$ bits.

◀

Similarly, by running $\log \delta^{-1}$ independent instances concurrently and return the maximum result, the success probability can be boosted to $1 - \delta$. This is because the probability that all instances fail is at most $(\frac{11}{30})^{\log \delta^{-1}} = \delta$. The space usage of the new algorithm after boosting is $\mathcal{O}(\log(\bar{d}\varepsilon^{-2}) \log \delta^{-1})$.

C Omitted Proofs from Section 4

C.1 Proof of Theorem 14

To start with, note that

▶ **Lemma 27.** *In every connected graph with $n \geq 3$, there is a maximum independent set containing all leaves and no support vertices.*

Proof. This can be proved via adjusting vertices from a given maximum independent set. Given a maximum independent set, if it contains all leaves, then we are done. If some leaves are not included, then their support vertices must be in the independent set. This can be seen via contradiction, assume that both the support vertex and its leaves are not in the independent set, then clearly we could add all the leaves in the set, which violates the assumption that the set is maximum. Note that each support vertex must be adjacent to at least one leaf, and if it is in the independent set, none of its leaves is in the set. Hence, we could remove all support vertices from the given independent set, and add all their leaves into the set, the size of the resulting set does not decrease. Given that the original set is already maximum, the new set is also maximum. ◀

Hence, in the proof of Theorem 14, we assume that the max independent set of every tree contains every leaf unless $n = 2$.

Proof of Theorem 14. To prove Theorem 14, it suffices to show that, for every tree T with $|T| \geq 2$,

$$\max\left\{\frac{|T|}{2}, |Deg_1(T)| - 1\right\} \leq \beta(T) \leq \frac{|T| + |Deg_1(T)|}{2}$$

First we show the lower bound for $\beta(T)$. Since a tree is bipartite and each side of the bipartition forms an independent set, it follows that $\beta(T) \geq |T|/2$. Also, when $|T| = 2$, $|Deg_1(T)| = 2$ and $\beta(T) = 1$. When $|T| \geq 3$, no two leaves can be adjacent to each other and hence the set of all leaves forms an independent set, i.e., $\beta(T) \geq |Deg_1(T)|$. Summarizing, we have $\beta \geq \max\left\{\frac{|T|}{2}, |Deg_1(T)| - 1\right\}$.

Next, we show that $\beta(T) \leq \frac{|T|+|Deg_1(T)|}{2}$ by induction on $|T|$. We say that a pair of leaves of a tree are twins if they share a common neighbor. We have two cases to consider depending on whether T has twins or not:

Case 1: T has twins Let the twins be v, w and their shared neighbor be u. Delete w, and let the tree obtained be $T' = T - w$. Then we have $|T'| = |T| - 1$, $|Deg_1(T')| = |Deg_1(T)| - 1$, and $\beta(T') = \beta(T) - 1$. By induction hypothesis, we know that

$$\begin{aligned} &|T'| + |Deg_1(T')| \geq 2\beta(T') \\ &\Rightarrow |T| - 1) + \big(|Deg_1(T)| - 1\big) \geq 2\big(\beta(T) - 1\big) \\ &\Rightarrow |T| + |Deg_1(T)| \geq 2\beta(T) \end{aligned}$$

Case 2: T does not have twins: Let a leaf w have a parent u whose parent is x. We can assume that u has no other leaves adjacent to it since T has no twins. Delete both u and w, and let the obtained tree be $T' = T - \{u, w\}$. It is easy to see that $|T'| = |T| - 2$ and $|Deg_1(T)| \geq |Deg_1(T')| \geq |Deg_1(T)| - 1$, because x may or may not be a leaf in T' but w was definitely a leaf in T. Given an independent set I in T, we can obtain an independent set $I' = I - w$ in T'. Similarly, given an independent set H' in T', we can obtain an independent set $H = H' \cup \{w\}$ in T. Hence, it follows that $\beta(T') = \beta(T) - 1$. By induction hypothesis, we have

$$\begin{aligned} &|T'| + |Deg_1(T')| \geq 2\beta(T') \\ &\Rightarrow \big(|T| - 2\big) + (|Deg_1(T')|) \geq 2(\beta(T) - 1) \\ &\Rightarrow |T| + |Deg_1(T)| \geq 2\beta(T). \end{aligned}$$

◀

C.2 Proof of Theorem 15

By Lemma 27, we will assume that (unless $n = 2$) we assume that the maximum independent set contains every leaf and none of the support vertices. We first show a lower bound (Claim 28) and upper bound (Claim 29) on the independence number of trees. Theorem 15 then follows immediately by applying Claim 28 and Claim 29 to each tree component of the forest F.

▷ Claim 28. For every tree T with $|T| \geq 2$,

$$\beta \geq \frac{1}{2}(|T| + |Deg_1(T)| - |Supp(T)|)$$

Proof. We prove the claim by showing that there exists an independent set of size at least $\frac{1}{2}\big(|T| + |Deg_1(T)| - |Supp(T)|\big)$. When $|T| = 2$, then T is a path on two vertices which implies $\beta(T) = 1$ and $|Deg_1| = 2 = |S| = 2$ and the desired inequality holds. Hence, we assume $n > 2$.

It suffices to show that there exists an independent set H containing all leaves and half of the vertices from $Deg_{\geq 2}(T)$ vertices that are not support vertices. We first remove all leaves and support vertices from T to obtain a forest F' with $|Deg_{\geq 2}(T)| - |Supp(T)|$ vertices. Since each forest is bipartite and each side of a bipartition forms an independent set, it follows that F' has an independent set H' of size $\geq |F'|/2$. We take H to be the union of $Deg_1(T)$ and H': it follows that H is an independent set in T since F' does not contain any support

vertices of T. The size of H is $|Deg_1(T)| + \frac{|Deg_{\geq 2}(T) - Supp(T)|}{2} = \frac{|T| + |Deg_1(T)| - Supp(T)}{2}$ since $|T| = |Deg_1(T)| + |Deg_{\geq 2}(T)|$. This concludes the proof of Claim 28. ◀

The lower bound of Claim 28 is tight: consider a path P on $2n$ vertices. Then $\beta(P) = n$ and $|Deg_1(P)| = 2 = |Supp(P)|$.

▷ Claim 29. For every tree T with $|T| \geq 2$,

$$\beta(T) \leq \frac{2}{3}(|T| + |Deg_1(T)| - |Supp(T)|)$$

Proof. We prove this lemma via induction on the graph order, n. Consider all trees with $2 \leq n \leq 4$ as the base cases: these can all be easily verified by hand. Assume the inequality holds for every tree with order $n - 1$. We say that two leaves are twins if they share a common parent. There are two cases to consider depending on whether T has twins or not:

Case 1: There exist twins in T. Let T' be the tree after removing either one of the twins. Then we have $|T| = |T'| + 1$, $|Deg_1(T)| = |Deg_1(T')| + 1$, and $|Supp(T)| = |Supp(T')|$. According to Lemma 27, we have that $\beta(T) = \beta(T') + 1$. Hence, by the induction hypothesis, the inequality holds.

Case 2: There are no twins in T. Consider the longest path in T, denote one of its leaves as x, the parent of x as y, the parent of y as w, and the parent of w as z. If T has more than four nodes and no twins, then z is guaranteed to be non-leaf. This can be seen by contradiction. Assuming z is a leaf and there is more than four nodes. By our choice of the longest path, the neighbors of w and y can only be leaves. Since there is no twin, w and y both have degree two, there are exactly four nodes, a contradiction. By similar argument, it is not hard to see that y must have degree two.

Case 2.1: $\mathbf{d(w) > 2}$. Remove both x and y, and let the resulting tree be T'. Note that this removes a leaf and a support vertex. Because w has degree more than 2, the removal of x and y does not make w a leaf, and w is a support vertex in T' if and only if it is a support vertex in T. Therefore, we have $|T| = |T'| + 2$, $|Deg_1(T)| = |Deg_1(T')| + 1$, and $|Supp(T)| = |Supp(T')| + 1$. Moreover, we claim that $\beta(T) = \beta(T') + 1$. This is because by Lemma 27, there is a max independent set H such that $x \in H$ and $y \notin H$: thus the existence of x and y has no impact on whether w and the rest of vertices are in $\beta(T)$ or not, i.e., they are in $\beta(T)$ if and only if they are also in $\beta(T')$. By the induction hypothesis, the inequality holds.

Case 2.2: $\mathbf{d(w) = 2}$. We have two cases to consider.

Case 2.2.1: $\mathbf{z \notin} Supp(\mathbf{T})$. We remove x and y. Similar to Case 2.1, we have $|T| = |T'| + 2$ and $\beta(T) = \beta(T') + 1$. However, since $d(w) = 2$ and z is not a support vertex in T, the removal of x and y makes w a leaf and z a support vertex. Hence $|Deg_1(T)| = |Deg_1(T')|$, and $|Supp(T)| = |Supp(T')|$. By the induction hypothesis, the inequality holds.

Case 2.2.2: $\mathbf{z} \in Supp(\mathbf{T})$. Remove x, y, and w from T and denote the remaining graph as T'. Since z is a support vertex, according to Lemma 27, there is a max indpendent set H of T such that $z \notin H$ which implies that w must in the maximum independent set. Therefore, we have $|T| = |T'| + 3$ and $\beta(T) = \beta(T') + 2$. Moreover, removing these vertices make z neither a leaf nor a support vertex in T', unless $d(z) = 2$. In the former case, we have $|Deg_1(T)| = |Deg_1(T')| + 1$, and $|Supp(T)| = |Supp(T')| + 1$. By the induction hypothesis, the inequality holds. In the latter case, T is a path of length 5: the claim can easily verified by hand for this specific graph. This concludes the proof of Claim 29.

◀

This upper bound of Claim 29 is asymptotically tight: consider the graph constructed from a star on $(r+1)$ vertices by replace each of the r leaves with a path consisting of 3 vertices. Clearly, this graph is a tree, say T, with $3r+1$ vertices, and $|Supp(T)| = |Deg_1(T)| = r$. Also, $\beta(T) = 2r$ as all the r leaves and their r grandparents form a maximum independent set.

C.3 Proof of Corollary 16

Proof. By definition, $n = |F| = |Deg_{\geq 2}(F)| + |Deg_1(F)|$ and $|Supp(F)| \leq |Deg_1(F)|$. Hence, if $|Supp(F)| \leq \frac{|Deg_{\geq 2}(F)|}{2}$, we have the following inequality,

$$|Supp(F)| \leq \frac{|Deg_{\geq 2}(F)|}{4} + \frac{|Deg_1(F)|}{2} = \frac{n + |Deg_1|}{4} \tag{6}$$

Utilizing Equation (6) and combining Theorem 14 and Theorem 15, if $|Supp(F)| \leq \frac{|Deg_{\geq 2}(F)|}{2}$, then

$$\frac{3(n + |Deg_1(F)|)}{8} \leq \frac{n + |Deg_1(F)| - |Supp(F)|}{2} \leq \beta(F)$$

By Theorem 14, we have $\beta(F) \leq \frac{n+|Deg_1(F)|}{2}$, Corollary 16 follows. ◀

C.4 Proof of Lemma 20

Proof. We prove this lemma by showing the following claim

▷ Claim 30. For each tree T (with at least two vertices) we have $|Deg_1(T)| = 2 + \sum_{i=3}^{\Delta}(i-2) \cdot |Deg_i(T)|$.

Proof.

$$\begin{aligned}
\sum_{i=3}^{\Delta}(i-2) \cdot |Deg_i(T)| &= \sum_{i=2}^{\Delta}(i-2) \cdot |Deg_i(T)| \\
&= \sum_{i=2}^{\Delta} i \cdot |Deg_i(T)| - 2 \cdot \sum_{i=2}^{\Delta}|Deg_i(T)| \\
&= \Big((2|T|-2) - Deg_1(T)\Big) - 2 \cdot \Big(|T| - Deg_1(T)\Big) \\
&= Deg_1(T) - 2
\end{aligned}$$

In the penultimate line, we have used the fact that $\sum_{v \in T} d(v) = 2|T| - 2$. ◀

The proof of Lemma 20 then follows from Claim 30 by summing up over all tree components of F.

◀

C.5 Proof of Lemma 23

Proof. We claim that, conditional on the algorithm does not abort at line 13, the candidate set, C, obtained at line 16 contains only vertices from $Supp(F)$, and if a vertex in $Supp(F)$ is sampled, it is kept in C. This is not hard to see, by performing a second pass, we obtain the exact degree of all sampled vertices and their neighbors. So that a vertex is included in C if there is at least one leaf adjacent to it, which is exactly the definition of $Supp(F)$.

It remains to bound the output of this algorithm, conditional on the algorithm does not abort. For all $i \in V(G)$, let the binary random variable X_i indicate whether a vertex v_i is a support vertex and is sampled in I. Let $X = \sum_{i \in G} X_i$. Clearly, since vertices are sampled uniformly with probability $p = \frac{c_1}{\varepsilon_1^2 K_1}$, we have

$$\begin{aligned} E[X] &= \sum_{i \in G} E[X_i] = \sum_{i \in G} \Pr[v_i \in (Supp(F) \cap I)] \\ &= \sum_{i \in Supp(F)} \Pr[v_i \in I] = \frac{c_1 \, |Supp(F)|}{\varepsilon_1^2 K_1} \end{aligned}$$

Thus by the algorithm design, $E[|\widehat{S}|] = \frac{\varepsilon_1^2 K_1}{c_1} \times E[X] = |Supp(F)|$. Moreover, when $|Supp(F)| \geq K_1$, we have

$$\begin{aligned} \Pr\left[\left||\widehat{S}| - |Supp(F)|\right| \geq \varepsilon_1 |Supp(F)|\right] &= \Pr\left[|X - E[X]| \geq \varepsilon_1 E[X]\right] \\ &\leq 2e^{-\frac{\varepsilon_1^2 E[X]}{3}} \\ &\leq 2e^{-c_1/3}, \end{aligned}$$

where we apply the negatively correlated Chernoff bound (Theorem 26) in the second-last inequality, since the X_i's are negatively correlated. And the last inequality holds as $|Supp(F)| \geq K_1$.

Next, we show that when $|Supp(F)| < K_1$, conditional on the algorithm does not abort, the probability that $|\widehat{S}|$ is larger than $2K_1$ is small. Again, by the negatively correlated Chernoff bound, we have

$$\begin{aligned} \Pr\left[|\widehat{S}| > 2K_1\right] &= \Pr\left[X > \frac{2K_1 E[X]}{|Supp(F)|}\right] \\ &< \Pr\left[X > (1 + \frac{K_1}{|Supp(F)|}) E[X]\right] \\ &\leq e^{-\frac{K_1^2 E[X]}{3|Supp(F)|^2}} \\ &\leq e^{-\frac{c_1 K_1}{3|Supp(F)|\varepsilon^2}} \\ &\leq e^{-\frac{c_1}{3\varepsilon^2}}, \end{aligned}$$

where the first and the last inequalities hold because $|Supp(F)| < K_1$.

Lastly, we bound the space usage of the algorithm and the probability that the algorithm aborts at line 13. By the design of the algorithm, at most $\frac{2m}{n} \frac{c_1 n}{\varepsilon_1^2 K_1} e^{\frac{c_1}{3}}$ vertices can be stored from the first pass, otherwise it aborts at line 13. By our assumption, the number of total edge insertions is bounded by $\mathcal{O}(n)$. Thus, we have $\frac{2m}{n} \in \mathcal{O}(1)$. Therefore, since each vertex requires $\mathcal{O}(\log n)$ bits to store its identifier, the space usage of the first pass is $\widetilde{\mathcal{O}}(\frac{n}{\varepsilon_1^2 K_1})$.

Moreover, in the second pass, each degree counter takes an additional $\mathcal{O}(\log n)$ space to store. Hence, the total space usage is still $\widetilde{\mathcal{O}}(\frac{n}{\varepsilon_1^2 K_1})$.

In this algorithm, we sampled $\frac{c_1 n}{\varepsilon_1^2 K_1}$ vertices in advance. And the average degree of the graph at any point of the stream can be calculated as $\frac{2m}{n}$. Hence, if we pick a vertex uniformly at random, the expected number of its neighbors is bounded by $\frac{2m}{n}$. Let Y_i be the number of neighbors of vertex i, $E[Y_i] = \frac{2m}{n}$. Let $Y = \sum_{v \in V(G)} Y_v$ We have

$$E[Y] = \sum_{v \in V(G)} E[Y_v] = \frac{2m}{n} \frac{c_1 n}{\varepsilon_1^2 K_1}$$

The algorithm aborts if more than $\frac{2m}{n} \frac{c_1 n}{\varepsilon_1^2 K_1} e^{\frac{c_1}{3}}$ vertices are retained. Since we are sampling each vertex uniformly at random, Y_v's are independent of each other. By Markov inequality, the probability that the actual size is $e^{\frac{c_1}{3}}$ times larger than the expected size is at most $e^{-\frac{c_1}{3}}$.

Applying a union bound with the abort probability and the concentration probabilities above, we can bound the fail probability of our algorithm as follows. When $|Supp(F)| \geq K_1$, Algorithm 3 outputs an $(1 \pm \varepsilon)$-estimate of $|Supp(F)|$ and fails with probability at most $2e^{-\frac{c_1}{3}} + e^{-\frac{c_1}{3}} = 3e^{-\frac{c_1}{3}}$. Similarly, when $|Supp(F)| < K_1$, the estimate returned by Algorithm 3 is larger than $2K_1$ with probability at most $e^{-\frac{c_1}{3\varepsilon_1^2}} + e^{-\frac{c_1}{3}} < 2e^{-\frac{c_1}{3}}$ since $\varepsilon_1 < 1$.

◀

C.6 Proof of Lemma 24

Proof. Note that in the first pass, only vertices in $Deg_{\geq 2}(F)$ are sampled. This is because line 6 decreases the frequency of all vertices by one, so leaves has 0 frequency and by the definition of k-sparse recovery, they are not recovered by the sparse recovery data structure. Moreover, by Theorem 3, when $|Deg_{\geq 2}(F)| \leq K_2$ (i.e., when it is small), the sparse recovery fails with probability no more than $1/c_2$.

Since the sparse recovery data structure might output a false positive result. That is, a k-sparse vector when the original degree vector is not k-sparse. We need to perform a sanity check at line 20 to ensure that the returned result contains all $Deg_{\geq 2}(F)$ vertices, rather than part of them. For this sanity check to work, we prove that the equality holds if and only if $H = Deg_{\geq 2}(F)$.

To begin with, we prove that the equality holds if $H = Deg_{\geq 2}(F)$. This is trivial as for every graph G, we have the following relationship between its edge count, m, and its total degree count.

$$2m = \sum_{v \in V(G)} d(v) = n - |Deg_{\geq 2}(G)| + \sum_{v \in Deg_{\geq 2}(G)} d(v)$$

Next, we prove that the equality does not hold if $H \neq Deg_{\geq 2}(F)$. Suppose $H \neq Deg_{\geq 2}(F)$, note that by our argument above, it is impossible to have x such that $x \in H$ but $x \notin Deg_{\geq 2}(F)$. Hence, we only consider the case where $\exists x \in Deg_{\geq 2}(F)$ s.t. $x \notin H$. We have

$$
\begin{aligned}
&2m - (n - |H| + \sum_{v \in H} d(v)) \\
&= n - |Deg_{\geq 2}(F)| + \sum_{v \in Deg_{\geq 2}(F)} d(v) - (n - |H| + \sum_{v \in H} d(v)) \\
&= \sum_{v \in Deg_{\geq 2}(F)} d(v) - |Deg_{\geq 2}(F)| - \sum_{v \in H} d(v) + |H| \\
&= \sum_{v \in Deg_{\geq 2}(F) \setminus H} d(v) - |Deg_{\geq 2}(F)| - \sum_{v \in H \setminus Deg_{\geq 2}(F)} d(v) + |H| \\
&= \sum_{v \in Deg_{\geq 2}(F) \setminus H} d(v) - |Deg_{\geq 2}(F)| - |H \setminus Deg_{\geq 2}(F)| + |H| \\
&= \sum_{v \in Deg_{\geq 2}(F) \setminus H} d(v) - |Deg_{\geq 2}(F) \setminus H| \\
&> 0
\end{aligned}
$$

where the second-last equality holds because by our definition of $Deg_{\geq 2}(F)$, all non-leaf vertices are in $Deg_{\geq 2}(F)$, hence vertices in $H \setminus Deg_{\geq 2}(F)$ are leaves and have degree of 1. And the last inequality holds because vertices in $Deg_{\geq 2}(F)$ have degree at least 2, and $Deg_{\geq 2}(F) \setminus H \neq \emptyset$. Hence, if some $Deg_{\geq 2}(F)$ vertices are not sampled, i.e., $Deg_{\geq 2}(F) \setminus H \neq \emptyset$, then $2m \neq n - |H| + \sum_{v \in H} d(v)$.

It remains to prove that conditional on all $Deg_{\geq 2}(F)$ vertices being sampled, both $|\widehat{S}|$ and $|\widehat{Deg_{\geq 2}}|$ are exact estimates of $|Supp(F)|$ and $|Deg_{\geq 2}(F)|$ respectively. If all $Deg_{\geq 2}(F)$ vertices are sampled successfully, we can infer whether a vertex is a leaf or not by testing whether it is in H or not. As indicated by line 16 – 19, by counting the number of leaves, l_v, for each $v \in H$, we can identify all the support vertices in H. But not all support vertices have degree greater than 1, by definition we can have support vertices of degree 1 in paths of length 2 (P_2), such that each P_2 has two support vertices. Thus, we need to also count the number of P_2, p. This is trivial as if an edge arrives with none of its endpoints in $Deg_{\geq 2}(F)$, this edge must be an instance of P_2.

Lastly, the space used by Algorithm 4 is $\widetilde{\mathcal{O}}(K_3)$. In the first pass, the sparse recovery structure takes $\widetilde{\mathcal{O}}(K_3)$ space to store. And in the second pass, for each vertex in $Supp(F)$, we introduce two counters that can be both stored in $\mathcal{O}(\log n)$ bits. Therefore the total space usage of Algorithm 4 is $\widetilde{\mathcal{O}}(K_3)$.

◀

C.7 Proof of Theorem 25

Proof. By Theorem 15 and Corollary 16, the min between $\frac{3(n+|Deg_1(F)|)}{8}$ and $\frac{n+|Deg_1(F)|-|Supp(F)|}{2}$ is guaranteed to be a 4/3 approximation of the independence number $\beta(F)$. Hence, it suffices to show that we could either approximate both terms well, or approximate the smaller term well.

When $|Deg_{\geq 2}(F)| \leq K_2 = 8\sqrt{n}$, the algorithm returns at line 11. By Lemma 24, Subroutine 4 gives exact estimates of $|Deg_{\geq 2}(F)|$ and $|Supp(F)|$ with probability $1 - 1/c_2 = 1 - \delta$. Also, since $n = |Deg_{\geq 2}(F)| + |Deg_1(F)|$ and n is known, we can obtain an exact estimate of $|Deg_1(F)|$. Hence, no matter which estimate is returned, it is a 4/3 approximation of $\beta(F)$.

When $|Deg_{\geq 2}(F)| > 8\sqrt{n}$, there are two cases to consider. If $|Supp(F)| \geq \sqrt{n}$, we may return the estimate of $\frac{3(n+|Deg_1(F)|)}{8}$ or the estimate of $\frac{n+|Deg_1(F)|-|Supp(F)|}{2}$ (line 8). On the one hand, if the estimate of $\frac{3(n+|Deg_1(F)|)}{8}$ is returned, by Lemma 21 we know that the algorithm on line 4 outputs a $(1 \pm \frac{\varepsilon}{2})$ estimate of $|Deg_1(F)|$ with probability $1 - \delta/2$. Thus our result is a $4/3(1 \pm \varepsilon/2)$ approximation of β.

On the other hand, if the estimate of $\frac{n+|Deg_1(F)|-|Supp(F)|}{2}$ is returned, by Lemma 23, Subroutine 3 gives a $(1 \pm \frac{\varepsilon}{2})$ estimate of $|Supp(F)|$ with probability

$$1 - 3e^{-c_1/3} = 1 - 3e^{-\ln(6\delta^{-1})} = 1 - \frac{\delta}{2}.$$

Applying a union bound over successfully returning an estimate of $|Supp(F)|$ and an estimate of $|Deg_1(F)|$, the probability of failure is at most δ. And on the upper-bound side, the approximation ratio is at most

$$\begin{aligned}
&\frac{1}{2}\left(n + (1 \pm \frac{\varepsilon}{2})|Deg_1(F)| - (1 \pm \frac{\varepsilon}{2})|Supp(F)|\right) \\
&\leq \frac{1}{2}\left(n + |Deg_1(F)| - |Supp(F)| + \frac{\varepsilon}{2}\left(|Deg_1| + |S|\right)\right) \\
&\leq (1+\varepsilon)\frac{1}{2}\left(n + |Deg_1(F)| - |Supp(F)|\right)
\end{aligned}$$

where the last inequality holds because $|Supp(F)| \leq |Deg_1(F)| \leq n$, thus $|Deg_1(F)| + |Supp(F)|$ is no more than $2(n+|Deg_1(F)|-|Supp(F)|)$. Similarly, we prove the approximation ratio on the lower bound as

$$\begin{aligned}
&\frac{1}{2}\left(n + (1 \pm \frac{\varepsilon}{2})|Deg_1(F)| - (1 \pm \frac{\varepsilon}{2})|Supp(F)|\right) \\
&\geq (1-\varepsilon)\frac{1}{2}\left(n + |Deg_1(F)| - |Supp(F)|\right)
\end{aligned}$$

Hence, our algorithm returns a $4/3 \cdot (1 \pm \varepsilon)$ approximation of $\beta(F)$.

Lastly, if $|Supp(F)| < \sqrt{n}$, then by Lemma 23, $|\widehat{S}| \leq 2\sqrt{n}$ with probability $1 - 2e^{-\frac{c_1}{3\varepsilon^2}}$. By Theorem 17, we know that

$$|\widehat{Deg_{\geq 2}}| \geq (1-\varepsilon) \cdot |Deg_{\geq 2}(F)| > \frac{|Deg_{\geq 2}(F)|}{2}$$

holds with probability $1 - \delta/2$ (if $\varepsilon < 1/2$). Since $|Deg_{\geq 2}(F)| \geq 8\sqrt{n}$, by applying a union bound, we can claim that the probability of $2|\widehat{S}| > |\widehat{Deg_{\geq 2}}|$ is at most $2e^{-\frac{c_1}{3\varepsilon^2}} + \delta/2 \leq \delta$. Therefore, with probability $1 - \delta$, the minimum between the two estimates is $\frac{3(n+|\widehat{Deg_1(F)}|)}{8}$. As shown before, $|\widehat{Deg_1(F)}|$ is a $1 \pm \varepsilon$ estimate of $|Deg_1(F)|$, hence the return estimate is a $4/3 \cdot (1 \pm \varepsilon)$ approximation of $\beta(F)$ with probability $1 - \delta$.

The space usage of Algorithm 5 is the maximum space usage of its three sub-algorithms, which is $\widetilde{\mathcal{O}}(\sqrt{n})$.

◀

C.8 Estimating Domination Number in Forests

C.8.1 3 Approximation

There are known results bound the domination number of forests, γ, in terms of the aforementioned graph parameters.

▶ **Theorem 31.** *[17, 35, 32] For every tree T with $n \geq 3$, $\frac{1}{3}|Deg_{\geq 2}| \leq \gamma \leq |Deg_{\geq 2}|$.*

In Theorem 31, constraint $n \geq 3$ accounts for the case of a tree with $n = 2$ vertices, a single edge: each vertex is a leaf, but $\gamma = 1$. If the upper bound is relaxed by 1, all trees are included, leading to the forest generalization of Theorem 31 referred to earlier as Theorem 19: For every forest F, $\frac{1}{3}|Deg_{\geq 2}| \leq \gamma \leq |Deg_{\geq 2}| + c$.

Recall that in Lemma 17, we show that the number of non-leaf vertices, $|Deg_{\geq 2}|$, can be $(1 \pm \varepsilon)$-approximated in a turnstile stream with probability $(1 - \delta)$ and in $\mathcal{O}(\log^{\mathcal{O}(1)} n \cdot \log \delta^{-1})$ space. Combining Lemma 17 with Theorem 19, we have:

▶ **Theorem 32.** *For every $\varepsilon, \delta \in (0, 1)$ and forest F, the domination number γ can be $(3 \pm \varepsilon)$-approximated in turnstile streams with probability $(1 - \delta)$ and in $\mathcal{O}(\log^{\mathcal{O}(1)} n \cdot \log \delta^{-1})$ space.*

C.8.2 2 Approximation

Similar to independence number, we can improve the approximation ratio of domination number to 2 with the help of support vertices, *Supp*.

▶ **Theorem 33.** *For every forest F, $\frac{1}{4}\big(|Deg_{\geq 2}| + |Supp|\big) \leq \gamma \leq \frac{1}{2}\big(|Deg_{\geq 2}| + |Supp|\big)$.*

Before proving Theorem 33, we first prove the following lemma.

▶ **Lemma 34.** *In every connected graph with $n \geq 3$, there is a minimum dominating set that contains all support vertices and no leaves.*

Proof. We prove this by showing that every minimum dominating set can be adjusted to contain no leaves and all support vertices without increasing its size. Clearly, in every connected graph with $n \geq 3$, all support vertices have degree greater than 1 (i.e., are not themselves leaves). If a support vertex has degree 1, then its neighbor must be a leaf, hence $n = 2$, violating our assumption about $n \geq 3$. Given a minimum dominating set, If it contains no leaves and all support vertices, then we are done. If not, we can remove the leaves from the dominating set and add their neighbors into the set. By the definition of support vertices, the neighbors of leaves are themselves support vertices. The resulting set is still a dominating set as the newly added support vertices dominate the removed leaves and themselves. Moreover, before the adjustment, for each pair of support vertex and leaf, at least one of them is in the dominating set, otherwise the domination property is broke. The adjustment ensures that the new dominating set contains all support vertices and no leaves. The size of the new dominating set is at most the size of original dominating set, because each removed leaf introduces at most one neighbor (i.e., support vertex) into the new set. ◀

Hence, in the following section, we assume the minimum dominating set contains all support vertices and no leaves. Now we give our upper and lower bounds on the domination number of forest. For the upper bound, we have

▶ **Lemma 35.** *For every forest F, we have*

$$\frac{|Deg_{\geq 2}(F)| + |Supp(F)|}{2} \geq \gamma(F)$$

Proof. If $|F| = 2$, then F has to be an edge. Every vertex of F is both a support vertex and a leaf, and we can pick either of them as the dominating set. Hence, $\frac{|Deg_{\geq 2}(F)|+|Supp(F)|}{2} = \frac{1+1}{2} = 1 = \gamma(F)$, the inequality holds.

Next, we prove the inequality for graphs with $n > 2$. It is known that if a graph G is split into k disjoint subgraphs, $G_1, G_2, \ldots, G_k$, we have $\gamma(G) \leq \gamma(G_1) + \gamma(G_2) + \cdots + \gamma(G_k)$. This is because connecting two or more disjoint graphs only introduces new edges, thus vertices do not become undominated. Rewriting the inequality as $|Supp(F)| + \frac{|Deg_{\geq 2}(F)|-|Supp(F)|}{2}$, the first term could be viewed as adding all vertices in $Supp(F)$ into the dominating set. By doing so, all of the vertices in $N\big[Supp(F)\big]$ are dominated (note that this includes all the leaves). The second term could be viewed as an upper bound on the domination number after removing all the leaves and support vertices. Clearly, the remaining graph is a forest with $(|Deg_{\geq 2}(F)| - |Supp(F)|)$ vertices. Thus it remains to prove that in such a forest, there exists a dominating set of size at most $\frac{|Deg_{\geq 2}(F)|-|Supp(F)|}{2}$.

Note that all isolated vertices in the induced graph are already dominated by some vertices in S, because a vertex becomes isolated only if all of its neighbors are in S. Thus, we only need to show that for each tree components (of size $k > 1$) in the remaining forest, there exists a dominating set of size at most $\frac{k}{2}$. Ore [39] proved that for every graph G with order n and no isolated vertices, $\gamma(G) \leq \frac{n}{2}$. Hence, this lemma follows.

◀

The upper bound of Lemma 35 is tight: construct a tree T by taking a path of on r vertices for any $r \geq 2$, add a leaf to every vertex on the path. It is easy to see that $|Supp(T)| = |Deg_{\geq 2}(T)| = \gamma(T) = r$.

Next, we present our lower bound result.

▶ **Lemma 36.** *For every tree T with $T| \geq 2$,*

$$\gamma(T) \geq \frac{|Deg_{\geq 2}(T)| + |Supp(T)|}{4}$$

Proof. We prove the lemma by induction on $n = |T|$. Consider the base cases as trees with $n \leq 4$, which can be easily verified by hand.

It remains to show the inductive step. Two or more leaves are twins if they share a common parent node (i.e., support vertex). If the tree contains twins, we remove any one of them from T to obtain a tree T'. By Lemma 34, the minimum dominating set of T does not contain either of the twins. Hence, we have that $\gamma(T) = \gamma(T')$, $|Deg_{\geq 2}(T)| = |Deg_{\geq 2}(T')|$, and $|Supp(T) = Supp(T')|$ stay the same. By induction hypothesis, Lemma 36 holds.

Next, assume there are no twins and consider the longest path, say P, in T. Denote the two endpoints as x and x' respectively, let y be the parent of x, w be the parent of y, and z be the parent of w. Note that the length of P must be greater than 4, thus z and x' must be different. This can be seen via contradiction. Assuming P has length 4, by our choice of longest path P, x and z (or equivalently, x') must be leaves, and the neighbors of y and w must be also leaves. Otherwise, our choice of the longest path is violated. If y or w is adjacent to leaves other than x and z, then there exist twins in the graph, violating

our assumption of no twins. Hence, the graph contains exactly 4 vertices, x, y, w, and z. It should be considered as the base case. Similar arguments can be obtained for P with length less than 4. Hence, P's length must be greater than 4.

If w is a support vertex, then we are done by induction. By Lemma 34, the minimum dominating set contains both w and y as they are support vertices. Note that y must have degree exactly 2, otherwise the graph has either a twin, or a path longer than P. Delete x and let the tree obtained be T''. Then y becomes a leaf which is already dominated by w. Therefore, after removing x, we have $\gamma(T'') = \gamma(T) - 1$. Since y becomes a leaf, we have $|Deg_{\geq 2}(T'')| = |Deg_{\geq 2}(T)| - 1$ and $|Supp(T'')| = |Supp(T)| - 1$, and hence the inequality holds by our induction hypothesis.

It remains to show that $\gamma(T) \geq \frac{|Deg_{\geq 2}(T)|+|Supp(T)|}{4}$ when there are no twins and w is not a support vertex. Let T' and T'' be the two sub-trees formed by deleting the edge (w, z), such that T' contains w and T'' contains z. Note that by our previous analysis, only paths of length 2 can be incident on w, otherwise the tree falls into previous cases. Denote the number of such paths as r; clearly, $r \geq 1$ because there is path from x to w. In tree T', clearly, by picking all r child nodes of w, we obtain a minimum dominating set. Since w is not included in the set, $\gamma(T) = \gamma(T') + \gamma(T'') = r + \gamma(T'')$. Let $\psi(T) = |Deg_{\geq 2}(T)| + |Supp(T)|$. We have the following three inequalities:

1. if $d(z) > 1$ in T'': $Deg_{\geq 2}(T) \leq Deg_{\geq 2}(T'') + (r + 1)$
2. if $d(z) = 1$ in T'': $Deg_{\geq 2}(T) \leq Deg_{\geq 2}(T'') + (r + 1) + 1$
3. $Supp(T) \leq Supp(T'') + r$

The first two inequalities hold because T' contains r vertices from $Deg_{\geq 2}(T')$ vertices, and by adding the edge (w, z) we have $d(w) \geq 2$. Also, if z has degree 1 in T'', z is not in $Deg_{\geq 2}(T'')$, but adding the edge (w, z) puts it in $Deg_{\geq 2}(T)$. Otherwise, z is already in $Deg_{\geq 2}(T'')$. Similarly, the *third* inequality holds because there are r vertices in $Supp(T')$, adding the edge (w, z) does not introduce new support vertices in $Supp(T)$. Hence, $Supp(T) \leq Supp(T'') + r$. This is an inequality because adding the edge (w, z) makes z non-leaf, thus a support vertex might be removed. Therefore, we have $\psi(T) \leq \psi(T'') + 2r + 2$. And:

$$\begin{aligned}
\gamma(T) &= \gamma(T'') + r \\
&\geq \frac{|Deg_{\geq 2}(T'')| + |Supp(T'')|}{4} + r \qquad \text{[by induction hypothesis]} \\
&= \frac{|Deg_{\geq 2}(T'')| + |Supp(T'')| + 4r}{4} \\
&\geq \frac{|Deg_{\geq 2}(T)| + |Supp(T)|}{4}
\end{aligned}$$

where the last inequality holds because $4r \geq 2r + 2$ when $r \geq 1$, which is always true since there is a path from w to x.

◀

The lower bound in Lemma 36 is asymptotically tight. Consider the tree T containing a vertex v with one leaf and r paths of length 4 (P_4) incident on it, which can be also viewed as a star graph with $r + 1$ leaves except r of them are replaced with P_4. Clearly, we have $\gamma(T) = r + 1$. Meanwhile, we have $|Deg_{\geq 2}(T)| = 3r + 1$, and $|Supp(T)| = r + 1$. Hence,

$$\frac{|Deg_{\geq 2}(T)| + |Supp(T)|}{4} = \frac{4r + 2}{4} \approx r + \frac{1}{2}.$$

When $r \to \infty$, $\frac{r+1}{r+\frac{1}{2}} \to 1$.

Combining Lemma 35 and 36, we have Theorem 33. Moreover, combining Theorem 33 and Theorem **??**, if $|Supp(F)| \leq \frac{|Deg_{\geq 2}(F)|}{3}$, we have

$$\frac{|Deg_{\geq 2}(F)|}{3} \leq \gamma(F) \leq \frac{|Deg_{\geq 2}(F)| + |Supp(F)|}{2} \leq \frac{2|Deg_{\geq 2}(F)|}{3}$$

Hence, Corollary 37 follows.

▶ **Corollary 37.** *If* $|Supp| \leq \frac{1}{3}|Deg_{\geq 2}|$*, then* $\frac{1}{3}|Deg_{\geq 2}| \leq \gamma \leq \frac{2}{3}|Deg_{\geq 2}|$*.*

As shown in Section 4.2, $|Supp|$ can be $(1 \pm \varepsilon)$-estimated in small spaces when it is large (Subroutine 3), and can be exactly estimated in small spaces when both it and $|Deg_{\geq 2}|$ are small (Subroutine 4). When it does not fall into either case, Corollary 37 suggests we can drop it without losing approximation ratio. Hence, using a similar approach to Algorithm 5, we can approximate the domination number with the desired ratio. Our main algorithm (Algorithm 6) and theorem (Theorem 38) for approximating the domination number are shown below.

Algorithm 6 Main Algorithm for Estimating γ

1: **Input**: error rate ε and fail rate δ
2: **Initialization**: Set $K_1 = \sqrt{n}$, $K_2 = 12\sqrt{n}$, $c_1 = 3\ln(6\delta^{-1})$, $c_2 = \delta^{-1}$, $\varepsilon_1 = \varepsilon$
3: Run the following algorithms concurrently:
4: 1. Algorithm for estimating $|Deg_{\geq 2}|$ with ε and $\delta/2$, denote the returned result as $|\widehat{Deg_{\geq 2}}|$
5: 2. Subroutine 3 with K_1, c_1, and ε_1, denote the returned result as $|\widehat{S}|$
6: 3. Subroutine 4 with K_2 and c_2, denote the returned result as $|\widehat{S}|'$ and $|\widehat{Deg_{\geq 2}}|'$
7: **if** Subroutine 4 returns *FAIL* **then**
8: **Return** $\widehat{\gamma} = \max\{\frac{2|\widehat{Deg_{\geq 2}}|}{3}, \frac{|\widehat{Deg_{\geq 2}}|+|\widehat{S}|}{2}\}$
9: **else**
10: **Return** $\widehat{\gamma} = \max\{\frac{2|\widehat{Deg_{\geq 2}}|'}{3}, \frac{|\widehat{Deg_{\geq 2}}|'+|\widehat{S}|'}{2}\}$

▶ **Theorem 38.** [⋆] *For every* $\varepsilon, \delta \in (0,1)$ *and forest* F *(with no isolated vertices), Algorithm 6 estimates the domination number* $\gamma(F)$ *in a turnstile stream within factor* $2 \cdot (1 \pm \varepsilon)$ *with probability* $(1-\delta)$ *in* $\widetilde{\mathcal{O}}(\sqrt{n})$ *space and two passes.*

Proof. By Theorem 33 and Corollary 37, the max between $\frac{2|Deg_{\geq 2}(F)|}{3}$ and $\frac{|Deg_{\geq 2}(F)|+|Supp(F)|}{2}$ is guaranteed to be a 2 approximation of the domination number $\gamma(F)$. Thus it suffices to show that we could approximate both terms well, or we could approximate the larger term well.

When $|Deg_{\geq 2}(F)| \leq K_2 = 12\sqrt{n}$, by Lemma 24, Subroutine 4 succeeds with probability $1 - 1/c_2 = 1 - \delta$. Thus, by the design of the algorithm, we return our estimate at line 10. By Lemma 24, conditional on successful returning, Subroutine 4 gives exact estimates of $|Deg_{\geq 2}(F)|$ and $|Supp(F)|$. Hence we have a 2-approximation of $\gamma(F)$.

When $|Deg_{\geq 2}(F)| > 12\sqrt{n}$, there are two cases to consider. If $|Supp(F)| \geq \sqrt{n}$, then the algorithm might return the estimate of $\frac{2|Deg_{\geq 2}(F)|}{3}$ or the estimate of $\frac{|Deg_{\geq 2}(F)|+|Supp(F)|}{2}$. On the one hand, if the estimate of $\frac{2|Deg_{\geq 2}(F)|}{3}$ is returned, by Theorem 17, we know that the algorithm on line 4 outputs a $(1 \pm \varepsilon)$ estimate of $|Deg_{\geq 2}(F)|$ with probability $1 - \delta/2$. Hence we obtain a $2(1 \pm \varepsilon)$ approximation of γ.

On the other hand, if the estimate of $\frac{|Deg_{\geq 2}(F)|+|Supp(S)|}{2}$ is returned by our algorithm, we know that by Lemma 23, Subroutine 3 gives a $(1 \pm \varepsilon)$ estimate of $|Supp(S)|$ with probability

$$1 - 3e^{-c_1/3} = 1 - 3e^{-\ln(6\delta^{-1})} = 1 - \frac{\delta}{2}$$

Apply a union bound over successfully returning an estimate of $|Supp(F)|$ and an estimate of $|Deg_{\geq 2}(F)|$, the probability of failure is at most δ. And the error rate is still $1 \pm \varepsilon$ as we are summing up the two terms. Hence, the returned estimate is still a $2(1 \pm \varepsilon)$ approximation of γ.

Lastly, if $|Supp(F)| < \sqrt{n}$, then by Lemma 23, $|\widehat{S}| \leq 2\sqrt{n}$ with probability $1 - 2e^{-\frac{c_1}{3}}$. By Theorem 17, if $\varepsilon < 1/2$, then

$$|\widehat{Deg_{\geq 2}}| \geq (1-\varepsilon)|Deg_{\geq 2}(F)| > \frac{|Deg_{\geq 2}(F)|}{2}$$

with probability $1-\delta/2$. Since we have $|Deg_{\geq 2}(F)| > 12\sqrt{n}$, we can bound the probability of $3|\widehat{S}| \leq |\widehat{Deg_{\geq 2}}|$ as

$$\begin{aligned}
\Pr\left[3|\widehat{S}| \leq |\widehat{Deg_{\geq 2}}|\right] &= \Pr\left[|\widehat{S}| \leq 2\sqrt{n} \wedge |\widehat{Deg_{\geq 2}}| > 6\sqrt{n}\right] \\
&\geq 1 - \Pr\left[|\widehat{S}| > 2\sqrt{n}\right] - \Pr\left[|\widehat{Deg_{\geq 2}}| \leq 6\sqrt{n}\right] \\
&\geq 1 - 2e^{-\frac{c_1}{3}} - \delta/2 \\
&\geq 1 - \delta
\end{aligned}$$

where we apply the union bound in the first inequality. Therefore, with probability at least $1-\delta$, the maximum between the two estimates, $\frac{2|\widehat{Deg_{\geq 2}}|}{3}$ and $\frac{|Deg_{\geq 2}|+|S|}{2}$, is $\frac{2|\widehat{Deg_{\geq 2}}|}{3}$. As shown before, $|\widehat{Deg_{\geq 2}}|$ is a $1 \pm \varepsilon$ estimate of $|Deg_{\geq 2}(F)|$, hence the returned estimate is a $2 \cdot (1 \pm \varepsilon)$ approximation of $\gamma(F)$ with probability $1 - \delta$.

The space usage of Algorithm 6 is the maximum space usage of its three sub-algorithms, which is $\widetilde{\mathcal{O}}(\sqrt{n})$.

◀

C.9 Estimating Matching Number in Forests

C.9.1 2 Approximation

There are known results bound the matching number of forests, ϕ, in terms of the aforementioned graph parameters.

Recall Theorem 18: For every forest F, $\max\left\{ c, \frac{1}{2}(|Deg_{\geq 2}| + c) \right\} \leq \phi \leq |Deg_{\geq 2}| + c$. In Lemma 17, we show that the number of non-leaf vertices, $|Deg_{\geq 2}|$, can be $(1\pm\varepsilon)$-approximated in a turnstile stream with probability $(1-\delta)$ and in $\mathcal{O}(\log^{\mathcal{O}(1)} n \cdot \log \delta^{-1})$ space. Combining Lemma 17 with Theorem 18, we have:

▶ **Theorem 39.** *For every $\varepsilon, \delta \in (0,1)$ and forest F, the matching number ϕ can be $(2 \pm \varepsilon)$-approximated in turnstile streams with probability $(1-\delta)$ and in $\mathcal{O}(\log^{\mathcal{O}(1)} n \cdot \log \delta^{-1})$ space.*

C.9.2 $3/2$ Approximation

With the help of support vertices, *Supp*, we can improve the approximation ratio of matching to $3/2$.

▶ **Theorem 40.** *For every forest F, $\frac{1}{3}\left(|Deg_{\geq 2}| + |Supp|\right) \leq \phi \leq \frac{1}{2}\left(|Deg_{\geq 2}| + |Supp|\right)$.*

Before proving Theorem 40, we prove the following lemma first.

▶ **Lemma 41.** *In every connected graph with $n \geq 3$, there is a maximum matching such that every support vertex is matched by one of its edges between it and its leaves.*

Proof. The first half of the lemma (every support vertex is matched) can be proved via contradiction. Assume we have a maximum matching and a support vertex s is not matched, then we could include an edge between s and one of its leaves into the matching, a contradiction.

The latter half can be proved via alternating edges in every given maximum matching without decreasing its size. Given a maximum matching M. If M satisfies the condition, then we are done. If there is a support vertex s that is matched through an edge between it and a non-leaf vertex, then by the definition of matching, none of its leaves are matched. Hence, we can remove the matching edge of s and include one of the edges between s and its leaves into the matching. This is valid as the leaf is not matched before, and s is matched only by the newly added edge. The size of maximum matching does not decrease. ◀

We may now prove an upper bound on ϕ,

▶ **Lemma 42.** *For every tree T with $n \geq 2$,*

$$\phi \leq \frac{|Deg_{\geq 2}(T)| + |Supp(T)|}{2}$$

Proof. We prove the lemma by induction on $n = |T|$. If $|T| = 2$, then T is an edge and we have $\phi(T) = 1, |Deg_{\geq 2}(T)| = 0$ and $|S| = 2$: hence the inequality holds. Henceforth, we assume that $n > 2$.

Rearranging the upper bound, we have

$$|Supp(T)| + \frac{|Deg_{\geq 2}(T)| - |Supp(T)|}{2}.$$

By Lemma 41, all support vertices are matched by the edges between them and one of their leaves. Hence, removing all support vertices and their leaves removes $|S|$ edges in the matching. The remaining graph is a forest F with $n(F) = |Deg_{\geq 2}(T)| - |S|$, and all edges in F can be added into the matching as they do not share a common endpoint with edges between support vertices and leaves (i.e., the matching edges that are removed). Hence it remains to prove that for every forest, there is a matching of size at most $\frac{n(F)}{2}$. Since every forest is bipartite, and in a bipartite graph with total order of n, the matching number is at most $\frac{n}{2}$. This is because one side has at most $\frac{n}{2}$ nodes. By the Pigeonhole Principle, if there is a matching of size greater than $\frac{n}{2}$, at least one node has more than one matching edges incident on it, a contradiction. ◀

The upper bound of Lemma 42 is tight: consider the tree T obtained by taking a star graph H with r leaves and attaching a new leaf for each vertex in H (including the center). It is easy to see that $|Deg_{\geq 2}(T)| = |Supp(T)| = r + 1$, as all vertices in H become a support vertex and have more than 1 neighbor. And it is not hard to see that $\phi = r + 1$, as the maximum matching is the edge set $E(T) \setminus E(H)$.

The next lemma gives a lower bound for the matching number:

▶ **Lemma 43.** *For every tree T with $|T| \geq 2$,*

$$\phi(T) \geq \frac{|Deg_{\geq 2}(T)| + |Supp(T)|}{3}$$

Proof. We prove the lemma by induction on $n = |T|$. We consider trees with $2 \leq n \leq 4$ as the base cases, which can be easily verified by hand. We say that a pair of vertices are twins if and only if they are leaves and adjacent to a common (support) vertex. There are two cases to consider depending on whether not T contains twins.

Case 1: There exist twins in T: Let x, y be the leaves which are twins and let s be the common vertex that they are adjacent to. Remove either of x or y, and let T' be the tree after the removal. Note that in T', s is a leaf if and only if $d(s) = 2$ in T, in this case $n = 3$, which falls into our base case. Hence, we have $|Deg_{\geq 2}(T)| = Deg_{\geq 2}(T')$ and $|Supp(T)| = |Supp(T')|$. Moreover, we claim that $\phi(T) = \phi(T')$: this is because only one of s's edges can be included into matching, if the removed leaf is on this edge, then we could add another edge between s and the other one of twin. By the induction hypothesis, this inequality holds.

Case 2: There are no twins in T: Consider the longest path in T, denote one of its endpoint as x, the parent of x as y, the parent of y as w, and the parent of w as z. By the induction hypothesis, T has more than 4 nodes and no twins, hence z is guaranteed to be a non-leaf. This can be seen by contradiction. Assuming z is leaf, by our choice of longest path, the neighbors of w and y can only be leaves. Otherwise, there exists a longer path. Also, since there is no twin, w and y both have degree 2, thus there is exactly 4 nodes, a contradiction. By similar argument, it is not hard to see that y is guaranteed to have degree 2.

Case 2.1: $\mathbf{d(w) > 2}$ We remove both x and y, denote the resulting tree as T'. Since $d(w) > 2$, removing x and y does not make w leaf, thus $|Deg_{\geq 2}(T)| = |Deg_{\geq 2}(T')|+1$. Also, w is a support vertex in T' if and only if it is a support vertex in T, we have $|Supp(T)| = |Supp(T')|+1$. By Lemma 41, we know that $e(x, y)$ is in the maximum matching of T, but not $e(y, w)$. Hence, the removal of x and y does not change the matching status of vertices in T': this implies that $\phi(T) = \phi(T') + 1$. By the induction hypothesis, the inequality holds.

Case 2.2: $\mathbf{d(w) = 2}$ We have two cases to consider:

Case 2.2.1: $\mathbf{z \notin} \mathit{Supp}(\mathbf{T})$ We remove x and y. Similar to Case 2.1, we have $\phi(T) = \phi(T') + 1$. However, since $d_w = 2$ and z is not a support vertex in T, the removal of x and y makes w a leaf and z a support vertex. Hence $|Deg_{\geq 2}(T)| = |Deg_{\geq 2}(T')| + 2$, and $|Supp(T)| = |Supp(T')|$. By the induction hypothesis, the inequality holds.

Case 2.2.2: $\mathbf{z} \in \mathit{Supp}(\mathbf{T})$ Consider T' as the tree after removing x, y, and w from T. Since z is a support vertex, according to Lemma 41, z must be matched and $e(x, y)$ is in the maximum matching. Hence, we know that $e(y, w)$ and $e(w, z)$ are not in the matching, and the removal of x, y, and w does not change the matching status of vertices in T'. Therefore, $\phi(T) = \phi(T') + 1$. Moreover, removing these three vertices does not make z a leaf or make it not a support vertex in T', unless z itself has degree 2. In the former case, we

have $|Deg_{\geq 2}(T)| = |Deg_{\geq 2}(T')| + 2$, and $|Supp(T)| = |Supp(T')| + 1$. And by the induction hypothesis, the inequality holds. In the later case, we have a path on 5 veryices since x is a leaf, y, w, and z have degree 2, and z is also a support vertex. It can be easily verified by hand that the inequality holds for $P5$.

◀

The lower bound is asymptotically tight. Consider a graph G constructed as follows. Start with a star graph, H, with r leaves, and replace each leaf with a path of length 3. G has $2r + 1$ non-leaf vertices, including the star center and two of the vertices on each of the paths. Also $|Supp(G)| = r$ as only the middle vertices on the paths are support vertices. Lastly, a maximum matching can be found by adding all edges between support vertices and leaves, and one edge from the edges incident on star center, thus $\phi(G) = r + 1$.

Theorem 40 follows if we combine Lemma 42 and Lemma 43. Moreover, combining Theorem 18 and Theorem 40, if $|Supp(F)| \leq \frac{|Deg_{\geq 2}(F)|}{3}$, we have:

$$\phi(F) \leq \frac{|Deg_{\geq 2}(F)| + |Supp(F)|}{2} \leq \frac{3(|Deg_{\geq 2}(F)| + c(F))}{4}$$

Since by Theorem 18, $\phi(F) \geq \max\{c(F), \frac{|Deg_{\geq 2}(F)|+c(F)}{2}\}$, Corollary 44 follows.

▶ **Corollary 44.** *If* $|Supp| \leq \frac{1}{2}|Deg_{\geq 2}|$, *then* $\max\left\{c, \frac{|Deg_{\geq 2}|+c}{2}\right\} \leq \phi \leq \frac{3}{4} \cdot \left(|Deg_{\geq 2}| + c\right)$.

As shown in Section 4.2, $|Supp|$ can be $(1 \pm \varepsilon)$-estimated in small spaces when it is large (Subroutine 3), and can be exactly estimated in small spaces when both it and $|Deg_{\geq 2}|$ are small (Subroutine 4). When it does not fall into either case, Corollary 37 suggests we can drop it without losing approximation ratio. Hence, using a similar approach to Algorithm 5 and Algorithm 6, we can approximate the matching number with the desired ratio. For simplicity, we omit the main algorithm and the proof of the main theorem (Theorem 45) here.

▶ **Theorem 45.** [⋆] *For every* $\varepsilon, \delta \in (0, 1)$ *and forest* F *(with no isolated vertices), Algorithm 5 estimates the matching number* $\phi(F)$ *in a turnstile stream within factor* $3/2 \cdot (1 \pm \varepsilon)$ *with probability* $(1 - \delta)$ *in* $\widetilde{\mathcal{O}}(\sqrt{n})$ *space and two passes.*