

## LP RELAXATION AND TREE PACKING FOR MINIMUM $k$ -CUT\*

CHANDRA CHEKURI<sup>†</sup>, KENT QUANRUD<sup>‡</sup>, AND CHAO XU<sup>§</sup>

**Abstract.** Karger used spanning tree packings [D. R. Karger, *J. ACM*, 47 (2000), pp. 46–76] to derive a near linear-time randomized algorithm for the global minimum cut problem as well as a bound on the number of approximate minimum cuts. This is a different approach from his well-known random contraction algorithm [D. R. Karger, *Random Sampling in Graph Optimization Problems*, Ph.D. thesis, Stanford University, Stanford, CA, 1995, D. R. Karger and C. Stein, *J. ACM*, 43 (1996), pp. 601–640]. Thorup developed a fast deterministic algorithm for the minimum  $k$ -cut problem via greedy *recursive* tree packings [M. Thorup, *Minimum  $k$ -way cuts via deterministic greedy tree packing*, in Proceedings of the Fortieth Annual ACM Symposium on Theory of Computing, ACM, 2008, pp. 159–166]. In this paper we revisit properties of an LP relaxation for  $k$ -CUT proposed by Naor and Rabani [*Tree packing and approximating  $k$ -cuts*, in Proceedings of the Twelfth Annual ACM-SIAM Symposium on Discrete Algorithms, Vol. 103, SIAM, Philadelphia, 2001, pp. 26–27], and analyzed in [C. Chekuri, S. Guha, and J. Naor, *SIAM J. Discrete Math.*, 20 (2006), pp. 261–271]. We show that the dual of the LP yields a tree packing that, when combined with an upper bound on the integrality gap for the LP, easily and transparently extends Karger’s analysis for mincut to the  $k$ -cut problem. In addition to the simplicity of the algorithm and its analysis, this allows us to improve the running time of Thorup’s algorithm by a factor of  $n$ . We also improve the bound on the number of  $\alpha$ -approximate  $k$ -cuts. Second, we give a simple proof that the integrality gap of the LP is  $2(1 - 1/n)$ . Third, we show that an optimum solution to the LP relaxation, for all values of  $k$ , is fully determined by the principal sequence of partitions of the input graph. This allows us to relate the LP relaxation to the Lagrangean relaxation approach of Barahona [*Oper. Res. Lett.*, 26 (2000), pp. 99–105] and Ravi and Sinha [*European J. Oper. Res.*, 186 (2008), pp. 77–90]; it also shows that the idealized recursive tree packing considered by Thorup gives an optimum dual solution to the LP.

**Key words.**  $k$ -cut, minimum cut, tree packing, approximation

**AMS subject classifications.** 05C40, 68R10, 05C85, 90C27

**DOI.** 10.1137/19M1299359

**1. Introduction.** The global minimum cut problem in graphs (MINCUT) is well known and extensively studied. Given an undirected graph  $G = (V, E)$  with nonnegative edge capacities  $c : E \rightarrow \mathbb{R}_+$ , the goal is to remove a minimum capacity set of edges such that the residual graph has at least two connected components. When all capacities are one, the mincut of a graph is its global edge-connectivity. The  $k$ -CUT problem is a natural generalization. Given a graph  $G = (V, E)$  and an integer  $k \geq 2$ , the goal is to remove a minimum capacity set of edges such that the residual graph has at least  $k$  connected components. MINCUT and  $k$ -CUT have been extensively studied in the literature. Initial algorithms for MINCUT were based on a reduction to the  $s$ - $t$ -mincut problem. However, it was later realized that it can be solved more efficiently and directly. Currently, the best deterministic algorithm for MINCUT runs in  $O(mn + n^2 \log n)$  time [35] and is based on the maximum adjacency ordering approach of Nagamochi and Ibaraki [27]. On the other hand, there is a near-linear time

\*Received by the editors November 13, 2019; accepted for publication (in revised form) March 29, 2020; published electronically June 16, 2020. A preliminary version of this paper appeared in *Proceedings of the Symposium on Simplicity in Algorithms (SOSA)*, 2019.

<https://doi.org/10.1137/19M1299359>

**Funding:** This research was supported in part by NSF grant CCF-1526799.

<sup>†</sup>Department of Computer Science, University of Illinois, Urbana, IL 61801 (chekuri@illinois.edu).

<sup>‡</sup>Department of Computer Science, Purdue University, West Lafayette, IN 47907 (krq@purdue.edu).

<sup>§</sup>Grab, Bellevue, WA 98004 (the.chao.xu@gmail.com, <https://chaoxuprime.com/>).

Monte Carlo randomized algorithm due to Karger [20]. Bridging the gap between the running times for the deterministic and randomized algorithms is a major open problem. Recent work [22, 17] obtained near-linear time deterministic algorithms for *simple* unweighted graphs.

The  $k$ -CUT problem is NP-hard if  $k$  is part of the input [12]; however, there is a polynomial-time algorithm for any fixed  $k$ . Such an algorithm was first devised by Goldschmidt and Hochbaum [12], and subsequently, there have been several different algorithms improving the run-time (see [18] and references therein). The randomized algorithm of Karger and Stein [21] runs in  $\tilde{O}(n^{2(k-1)})$  time and outputs the optimum cut with high probability. The fastest deterministic algorithm, due to Thorup [37], runs in  $\tilde{O}(mn^{2k-2})$  time [37]. Recent work of Gupta, Lee, and Li [13] obtains a faster run-time of  $\tilde{O}(k^{O(k)}n^{(2\omega/3+o(1))k})$  if the graph has small integer weights, where  $\omega$  is the exponent in the run-time of matrix multiplication. It is also known that  $k$ -CUT is  $W[1]$ -hard when parameterized by  $k$  [8]; that is, we do not expect an algorithm with a run-time of  $f(k)n^{O(1)}$ . Several polynomial-time 2-approximation algorithms are known for  $k$ -CUT even when  $k$  is part of the input; the first one is due to Saran and Vazirani and their algorithm is based on repeated minimum-cut computations and yields a  $(2 - 2/k)$ -approximation [33]; the same bound can be achieved by removing the  $(k - 1)$  smallest weight edges in a Gomory–Hu tree of the graph [33]. Nagamochi and Kamidoi showed that using the concept of extreme sets, a  $(2 - 2/k)$ -approximation can be found even faster [28]. Naor and Rabani developed an LP relaxation for  $k$ -CUT [29] and this yields a  $2(1 - 1/n)$ -approximation [4]. Ravi and Sinha [32] obtained another  $2(1 - 1/n)$ -approximation via a Lagrangean relaxation approach which was also considered independently by Barahona [2]. Quanrud obtained a deterministic  $(2 + \epsilon)$ -approximation in near-linear time via the LP relaxation [31]. A factor of 2, for large  $k$ , is the best possible approximation under the Small Set Expansion hypothesis [26]. Recent work has obtained a 1.81 approximation in  $2^{O(k^2)}n^{O(1)}$  time [14]; whether a polynomial time approximation scheme (PTAS) can be obtained in  $f(k)\text{poly}(n)$  time is an interesting open problem.

**Motivation and contributions.** The initial motivation for this work was to understand and simplify Thorup’s tree packing based algorithm for  $k$ -CUT. Karger’s near-linear time algorithm and analysis for the MINCUT problem [20] is based on the well-known theorem of Tutte and Nash-Williams (on the min-max relation for the maximum number of edge-disjoint trees in a graph). It is simple and elegant; the main complexity is in the improved running time which is achieved via a nontrivial dynamic programming based algorithm relying on advanced data structures. Karger also tightened the bound on the number of  $\alpha$ -approximate minimum cuts in a graph via tree packings (originally shown via his random contraction algorithm). In contrast, the main structural result in Thorup’s work on  $k$ -CUT is more difficult to understand. His proof consists of two parts. He shows that an ideal tree packing obtained via a recursive decomposition of the graph, first outlined in [36], has the property that any optimum  $k$ -cut crosses some tree in the packing at most  $2k - 2$  times. The second part argues that a greedy tree packing with sufficiently many trees approximates the ideal tree packing arbitrarily well. The greedy tree packing is closely related to a multiplicative weight update method for solving a basic tree packing linear program; however, no explicit LP is used in Thorup’s analysis. Thus, although Thorup’s algorithm is very simple to describe (and implement), the motivation for the tree packing and analysis are not very transparent.

In this paper we make several contributions which connect Thorup’s tree packing

to the LP relaxation for  $k$ -CUT [29]. We outline the specific contributions below.

- We show that the dual of the LP for  $k$ -CUT gives a tree packing and one can use a simple analysis, very similar to that of Karger, to show that any optimum  $k$ -cut crosses some tree in the packing at most  $(2k - 3)$  times. This improves a bound of  $(2k - 2)$  that Thorup showed for his tree packing and leads to a deterministic algorithm with a running time of  $\tilde{O}(mn^{2k-3})$  to find an optimum  $k$ -cut in a weighted graph with  $n$  nodes and  $m$  edges. For weighted graphs this is currently the fastest known deterministic algorithm. We also obtain an improved bound on the number of  $\alpha$ -approximate  $k$ -cuts.
- We give a new and simple proof that the integrality gap of the LP for  $k$ -CUT is upper bounded by  $2(1 - 1/n)$ . We note that the proof claimed in [29] was incorrect and the proof in [4] is indirect and technical. We also obtain a seemingly new and simple proof of the min-max relation between maximum fractional tree packing and network strength originally due to Tutte and Nash-Williams.
- We show that the optimum solution of the  $k$ -CUT LP, for all values of  $k$ , can be completely characterized by the principal sequence of partitions of the cut function of the given graph. This establishes the connection between the dual of the LP relaxation and the ideal recursive tree packing considered by Thorup. It also shows that the lower bound provided by the LP relaxation is equivalent to the Lagrangean relaxation lower bound considered by Barahona [2] and Ravi and Sinha [32].

Our results help unify and simplify the different approaches to  $k$ -CUT via the LP relaxation and its dual. Since the main motivation for this paper is to simplify and improve the understanding of the tree packing approach, we take a leisurely path and reprove some of Karger's results for the sake of completeness, and to point out the similarity of our argument for  $k$ -CUT to the case of MINCUT. Readers familiar with [20] may wish to skip section 3.

**Other related work.** There have been several developments on the  $k$ -CUT problem since the publication of a preliminary version of this paper [14, 13, 31, 16, 24]. Building on a sequence of developments on  $k$ -CUT, Gupta, Lee, and Li [15] show an improved analysis of the Karger–Stein algorithm which shows that (i) the number of minimum  $k$ -cuts is  $n^{(1+o(1))k}$  and (ii) the running time is  $n^{(1+o(1))k}$ . These bounds are essentially tight modulo the small additive term in the exponent. We refer the reader to these recent papers for more details.

Computing a minimum  $k$ -cut in a hypergraph is also an interesting problem, whose complexity was open until recently. Chandrasekaran, Xu, and Yu obtained a randomized polynomial time algorithm for any fixed  $k$  [3]. When  $k$  is part of the input, in contrast to the case of graphs, the  $k$ -CUT problem in hypergraphs is hard to approximate to within large factors under the Exponential Time Hypothesis [5, 25].

**Organization.** Section 2 sets up some basic notation and definitions. Section 3 discusses Karger's approach for MINCUT via tree packings with some connections to recent developments on approximately solving tree packings. Section 4 describes the tree packing obtained from the dual of the LP relaxation for  $k$ -CUT and how it can be used to extend Karger's approach to  $k$ -CUT. Section 5 gives a new proof that the LP integrality gap for  $k$ -CUT is  $2(1 - \frac{1}{n})$ . In section 6 we show that the optimum LP solution for all values of  $k$  can be characterized by a recursive decomposition of the input graph.

**2. Preliminaries.** We use  $n$  and  $m$  to denote the number of nodes and edges in a given graph. For a graph  $G = (V, E)$ , let  $\mathcal{T}(G)$  denote the set of spanning trees of  $G$ . For a graph  $G = (V, E)$  with edge capacities  $c : E \rightarrow \mathbb{R}_+$  the *fractional spanning tree packing number*, denoted by  $\tau(G)$ , is the optimum value of a simple linear program shown in Figure 1 whose variables are  $y_T, T \in \mathcal{T}(G)$ . The LP has an exponential number of variables but is still polynomial time solvable. There are several ways to see this, and efficient strongly combinatorial algorithms are also known [10]. We also observe that there is an optimum solution to the LP whose support has at most  $m$  trees since the number of nontrivial constraints in the LP is at most  $m$  (one per each edge).

$$\begin{aligned} \max \quad & \sum_{T \in \mathcal{T}(G)} y_T \\ & \sum_{T \ni e} y_T \leq c_e \quad e \in E \\ & y_T \geq 0 \quad T \in \mathcal{T}(G) \end{aligned}$$

FIG. 1. LP relaxation defining  $\tau(G)$ .

There is a min-max formula for  $\tau(G)$  which is a special case of the min-max formula for matroid base packing due to Tutte and Nash-Williams [34]. To state this theorem we introduce some notation. For a partition  $\mathcal{P}$  of the vertex set  $V$  let  $E(\mathcal{P})$  denote the set of edges that cross the partition (that is, have end points in two different parts), and let  $|\mathcal{P}|$  denote the number of parts of  $\mathcal{P}$ . A  $k$ -cut is  $E(\mathcal{P})$  for some partition  $\mathcal{P}$  such that  $|\mathcal{P}| \geq k$ . A *cut* is a 2-cut. The value of the minimum cut of  $G$  is denoted as  $\lambda(G)$ . It is not hard to see that for any partition  $\mathcal{P}$  of the vertex set  $V$ ,  $\tau(G) \leq \frac{c(E(\mathcal{P}))}{|\mathcal{P}|-1}$  since every spanning tree of  $G$  contains at least  $|\mathcal{P}| - 1$  edges from  $E(\mathcal{P})$ . The minimum over all partitions of the quantity,  $\frac{c(E(\mathcal{P}))}{|\mathcal{P}|-1}$ , is also referred to as the *strength* of  $G$  (denoted by  $\sigma(G)$ ), and turns out to be equal to  $\tau(G)$ .

**THEOREM 2.1** (Tutte and Nash-Williams (see [34])). *For any undirected edge capacitated graph  $G$ ,*

$$\tau(G) = \min_{\mathcal{P}} \frac{c(E(\mathcal{P}))}{|\mathcal{P}| - 1}.$$

Tutte and Nash-Williams proved the integer packing version of the preceding theorem which is harder; they showed that the maximum number of edge-disjoint spanning trees in a graph  $G$  with integer capacities  $c$  is given by  $\min_{\mathcal{P}} \lfloor \frac{c(E(\mathcal{P}))}{|\mathcal{P}|-1} \rfloor$ . The theorem is, in fact, a special case of matroid base packing theorem and can also be derived via the matroid union theorem of Edmonds; we refer the reader to [34].

A useful and well-known corollary of the preceding theorem is given below.

**COROLLARY 2.2.** *For any graph  $G$ ,  $\tau(G) \geq \frac{n}{2(n-1)} \cdot \lambda(G)$ .*

*Proof.* Consider the partition  $\mathcal{P}^*$  that achieves the minimum in the min-max formula. We have  $c(E(\mathcal{P}^*)) \geq |\mathcal{P}^*| \lambda(G) / 2$  since the capacity of edges leaving each

part of  $\mathcal{P}^*$  is at least  $\lambda(G)$  and an edge in  $E(\mathcal{P}^*)$  crosses exactly two parts. Thus,

$$\tau(G) = \frac{c(E(\mathcal{P}^*))}{|\mathcal{P}^*| - 1} \geq \frac{|\mathcal{P}^*|\lambda(G)}{2(|\mathcal{P}^*| - 1)} \geq \frac{n}{2(n-1)}\lambda(G)$$

since the ratio  $\frac{|\mathcal{P}^*|}{|\mathcal{P}^*| - 1}$  decreases as  $|\mathcal{P}^*|$  increases, and  $|\mathcal{P}^*| \leq n$ .  $\square$

We say that a tree packing  $y : \mathcal{T}(G) \rightarrow \mathbb{R}_+$  is  $(1-\epsilon)$ -approximate if  $\sum_{T \in \mathcal{T}(G)} y_T \geq (1-\epsilon)\tau(G)$ . Note that we typically want a compact tree packing that can either be specified explicitly via a small number of trees or even implicitly via a data structure representing a collection of trees. Approximate spanning tree packings have been obtained via greedy spanning tree packings which can be viewed as applying the multiplicative weight update method. Recently, [6] obtained the following result.

**THEOREM 2.3** (see [6]). *There is a deterministic algorithm that, given an edge-capacitated undirected graph on  $m$  edges and an  $\epsilon \in (0, 1/2)$ , runs in  $O(m \log^3 n/\epsilon^2)$  time and outputs an implicit representation of a  $(1-\epsilon)$ -approximate tree packing.*

**3. Tree packings and minimum cuts.** We review some of Karger's observations and results connecting tree packings and minimum cuts [20] which follow relatively easily via Corollary 2.2. In this section, we restrict the definition of a cut to an edge set  $E(\mathcal{P})$  for a partition  $\mathcal{P} = \{S, V \setminus S\}$  with *exactly* two parts. Hence, we use  $\delta(S)$  to uniquely identify a cut. However, as one can see later in section 4, the results will also hold for the more general definition of a cut, where  $\mathcal{P}$  has *at least* two parts. We rephrase Karger's results and arguments with a slightly different notation. Given a spanning tree  $T$  and a cut  $A \subseteq E$ , following Karger, we say that  $T$   $h$ -respects  $A$  for some integer  $h \geq 1$  if  $|E(T) \cap A| \leq h$ .

Karger proved that a constant fraction of trees (in the weighted sense) of an optimum packing 2-respect any fixed mincut. In fact this holds for a  $(1-\epsilon)$ -approximate tree packing for sufficiently small  $\epsilon$ . The proof, as follows, is an easy consequence of Corollary 2.2 and an averaging argument. It is convenient to view a tree packing as a probability distribution. Let  $p_T = y_T/\tau(G)$ . We then have  $\sum_T p_T = 1$  for an exact tree packing and for a  $(1-\epsilon)$ -packing we have  $\sum_T p_T \in (1-\epsilon, 1]$ . Let  $\delta(S)$  be a fixed minimum cut whose capacity is  $\lambda(G)$ . Let  $\ell_T = |E(T) \cap \delta(S)|$  be the number of edges of  $T$  that cross  $S$ . Let  $q = \sum_{T: \ell_T \leq 2} p_T$  be the fraction of trees that 2-respect  $\delta(S)$ . Since each tree crosses  $S$  at least once we have

$$\sum_T p_T \ell_T = \sum_{T: \ell_T \leq 2} p_T \ell_T + \sum_{T: \ell_T \geq 3} p_T \ell_T \geq q + 3(1-\epsilon-q).$$

Because  $y$  is a valid packing,

$$\tau(G) \sum_T p_T \ell_T = \sum_T y_T \ell_T \leq c(\delta(S)) = \lambda(G).$$

Putting the two inequalities together and using Corollary 2.2,

$$3(1-\epsilon) - 2q \leq \lambda(G)/\tau(G) \leq 2(n-1)/n,$$

which implies that

$$q \geq \frac{3}{2}(1-\epsilon) - (1-1/n) = \frac{1}{2} + \frac{1}{n} - \frac{3\epsilon}{2}.$$

If  $\epsilon = 0$ , this implies that at least half the fraction of trees 2-respect any minimum cut. Let  $q'$  be the fraction of trees that 1-respect a minimum cut. One can do similar calculations as above to conclude that

$$q' \geq 2(1 - \epsilon) - 2(1 - 1/n) \geq 2\left(\frac{1}{n} - \epsilon\right).$$

Thus,  $q' > 0$  as long as  $\epsilon < 1/n$ . In an optimum packing there is always a tree in the support that 1-respects a mincut. The preceding argument can be generalized in a direct fashion to yield the following useful lemma on  $\alpha$ -approximate cuts.

**LEMMA 3.1.** *Let  $y$  be a  $(1 - \epsilon)$ -approximate tree packing. Let  $\delta(S)$  be an  $\alpha$ -approximate minimum cut (i.e.,  $c(\delta(S)) \leq \alpha\lambda(G)$ ) for some fixed  $\alpha \geq 1$ . For a fixed integer  $h \geq 1$ , let  $q_h$  denote the fraction of trees in the packing  $y$  that  $h$ -respect  $\delta(S)$ . Then*

$$q_h \geq (1 - \epsilon) \left(1 + \frac{1}{h}\right) - \frac{2\alpha}{h} \left(1 - \frac{1}{n}\right).$$

**3.1. Number of approximate minimum cuts.** Karger showed that the number of  $\alpha$ -approximate minimum cuts is at most  $O(n^{2\alpha})$  via his random contraction algorithm [19]. He improved the bound to  $O(n^{\lfloor 2\alpha \rfloor})$  (for any fixed  $\alpha$ ) via tree packings in [20]. We review the latter argument.

For any cut  $\delta(S)$ , we associate with it the subset of edges of  $T$  that cross  $S$ ,  $E(T) \cap \delta(S)$ . In the other direction, removing a set of edges  $A \subseteq E(T)$  induces several components in  $T - A$ , which induces a unique cut in  $G$  where any two components of  $T - A$  adjacent in  $T$  lie on opposite sides of the cut. This gives a bijection between cuts induced by edge removals in  $T$ , and cuts in the graph.

Fix  $\alpha > 1$ , and let  $h = \lfloor 2\alpha \rfloor$ . Let  $y$  be a fixed optimum tree packing supported by some  $m' \leq m$  trees. For any  $\alpha$ -approximate mincut  $\delta(S)$ , by Lemma 3.1 and some simplification, the fraction of trees in the packing  $y$  that  $h$ -respects  $\delta(S)$  is at least

$$q_{h,\alpha} \equiv \frac{1}{\lfloor 2\alpha \rfloor} (1 - (2\alpha - \lfloor 2\alpha \rfloor)(1 - 1/n)).$$

Observing that  $q_{h,\alpha} > 0$ , an easy counting argument for approximate mincuts is the following. For each  $\alpha$ -approximate mincut, there is at least one tree in the (support of the) packing  $y$  which crosses it at most  $h$  times. Hence each  $\alpha$ -approximate cut can be mapped to a distinct combination of a tree in the packing and at most  $h$  edges from that tree. The total number of these combinations is  $m' \sum_{i=1}^h \binom{n-1}{i} = O(mn^h)$ .

We can avoid the factor of  $m$  by leveraging the fact that  $q_{h,\alpha}$  is a constant for every fixed  $\alpha$ . We give an informal argument here. A tree can  $h$ -respect at most  $2^{h+1}n^h$  distinct cuts (removing  $\leq h$  edges from  $T$  creates  $\leq h + 1$  components and these induce at most  $2^{h+1}$  cuts), while each  $\alpha$ -approximate minimum cut is  $h$ -respected by a (constant)  $q_{h,\alpha}$ -fraction of the tree packing; some of these details will be more clear in section 4.1. It follows by a packing argument that the total number of  $\alpha$ -approximate mincuts is at most  $\frac{2^{h+1}n^h}{q_{h,\alpha}} = O(n^h)$  for any fixed  $\alpha$ .

**3.2. Algorithm for minimum cut via tree packings.** Karger used tree packings to obtain a randomized near-linear time algorithm for the global minimum cut. The algorithm is based on combining the following two steps.

- Given a graph  $G$  there is a randomized algorithm that outputs  $O(\log n)$  trees in  $\tilde{O}(m)$  time such that with high probability there is a global minimum cut that 2-respects one of the trees in the packing.

- There is a deterministic algorithm that, given a graph  $G$  and a spanning tree  $T$ , in  $\tilde{O}(m)$  time finds the cut of minimum capacity in  $G$  that 2-respects  $T$ . This is based on a clever dynamic programming algorithm that utilizes dynamic tree data structures.

Only the first step of the algorithm is randomized. Karger solves the first step as follows. Given a capacitated graph  $G$  and an  $\epsilon > 0$ , he sparsifies the graph  $G$  to obtain an unweighted skeleton graph  $H$  via random sampling such that (i)  $H$  has  $O(n \log n / \epsilon^2)$  edges, (ii)  $\lambda(H) = \Theta(\log n / \epsilon^2)$ , and (iii) a minimum cut of  $G$  corresponds to a  $(1 + \epsilon)$ -approximate minimum cut of  $H$  in that the cuts induce the same vertex partition. Karger then uses greedy tree packing in  $H$  to obtain a  $(1 - \epsilon')$ -tree packing in  $H$  with  $O(\log n / \epsilon'^2)$  trees, and via Corollary 2.2 argues that one of the trees in the packing 2-respects a mincut of  $G$ ; here  $\epsilon$  and  $\epsilon'$  are chosen to be sufficiently small but fixed constants.

We observe that Theorem 2.3 can be used in place of the sparsification step of Karger. The deterministic algorithm implied by the theorem can be used to find an implicit  $(1 - \epsilon)$ -approximate tree packing in near-linear time for any fixed  $\epsilon > 0$ . For sufficiently small but fixed  $\epsilon$ , a constant fraction of the trees in the tree packing 2-respect any fixed minimum cut. Thus, if we sample a tree  $T$  from the tree packing, and then apply Karger's deterministic algorithm for finding the smallest cut that 2-respects  $T$ , we can find a minimum cut with constant probability. We can repeat the sampling  $\Theta(\log n)$  times to obtain a high probability bound.

Karger raised the following question in his paper. Can the dynamic programming algorithm for finding the minimum cut that 2-respects a tree be made *dynamic*? That is, suppose  $T$  is altered via edge swaps to yield a tree  $T' = T - e + e'$ , where  $e \in E(T)$  is removed and replaced by a new edge  $e'$ . Can the solution for  $T$  be updated quickly to obtain a solution for  $T'$ ? Note that  $G$  is static; only the tree is changing. The tree packing from Theorem 2.3 finds an implicit packing via  $\tilde{O}(m)$  edge swap operations from a starting tree  $T_0$ . Suppose there is a dynamic version of Karger's dynamic program that handles updates to the tree in amortized  $g(n)$  time per update. This would yield a *deterministic* algorithm for the global mincut with a total time of  $\tilde{O}(mg(n))$ . We note that the best deterministic algorithm for capacitated graphs is  $O(mn + n^2 \log n)$  [35]. This would be improved by any  $g(n) = o(n)$ .

**4. Tree packings for  $k$ -Cut via the LP relaxation.** In this section we consider the  $k$ -CUT problem. Thorup [36] constructed a probability distribution over spanning trees which were obtained via a recursive greedy tree packing and showed that there is a tree  $T$  in the support of the distribution such that a minimum weight  $k$ -cut contains at most  $2(k - 1)$  edges of  $T$ . He then showed that greedy tree packing with  $O(mk^3 \log n)$  trees closely approximates the ideal distribution. With this approach, he derived the currently fastest known deterministic algorithm to find the minimum  $k$ -CUT in  $\tilde{O}(mn^{2k-2})$  time. This is only slightly slower than the randomized Monte Carlo algorithm of Karger and Stein [21] whose algorithm runs in  $\tilde{O}(n^{2k-2})$  time. Thorup's algorithm is fairly simple. However, the proofs are somewhat complex since they rely on the recursive tree packing and its subtle properties. Arguing that the greedy tree packing approximates the recursive tree packing is also technical.

Here we consider a different tree packing for  $k$ -CUT that arises from the LP relaxation for  $k$ -CUT considered by Naor and Rabani [29]. This LP relaxation is shown in Figure 2. The variables are  $x_e \in [0, 1]$ ,  $e \in E$ , which indicate whether an edge  $e$  is cut or not. There is a constraint for each spanning tree  $T \in \mathcal{T}(G)$ ; at least  $k - 1$  edges from  $T$  need to be chosen in a valid  $k$ -cut. We note that for  $k > 2$  the

$$\begin{aligned}
 & \min \sum_{e \in E} c_e x_e \\
 & \text{s.t. } \sum_{e \in T} x_e \geq k - 1 \text{ for all } T \in \mathcal{T}(G) \\
 & \quad x_e \leq 1 \text{ for all } e \in E \\
 & \quad x_e \geq 0 \text{ for all } e \in E
 \end{aligned}$$

FIG. 2. An LP relaxation for the  $k$ -CUT problem from [29].

upper bound constraint  $x_e \leq 1$  is necessary.

The dual of the LP is given in Figure 3. Naor and Rabani claimed an integrality gap of 2 for the  $k$ -CUT LP. Their proof was incomplete and a correct proof was given in [4] in the context of a more general problem called the Steiner  $k$ -CUT problem. Let  $\lambda_k(G)$  denote the minimum  $k$ -cut capacity in  $G$ .

**THEOREM 4.1** (see [4]). *The worst case integrality gap of the LP for  $k$ -CUT in Figure 2 is  $2(1 - 1/n)$ .*

**COROLLARY 4.2.** *Let  $(y, z)$  be an optimum solution for the dual LP for  $k$ -CUT shown in Figure 3. Then*

$$(k - 1) \sum_T y_T \geq \frac{n\lambda_k(G)}{2(n - 1)} + z(E).$$

Note that Corollary 2.2 is a special case of the preceding corollary.

*Remark 1.* We note that the LP relaxation in Figure 2 assumes that  $G$  is connected. This is easy to ensure by adding dummy edges of zero cost to make  $G$  connected. However, it is useful to consider the general case when the number of connected components in  $G$  is  $h$  where we assume for simplicity that  $h < k$  (if  $h \geq k$ , the problem is trivial). In this case we need to consider the maximal forests in  $G$ , each of which has exactly  $n - h$  edges; to avoid notational overload we use  $\mathcal{T}(G)$  to denote the set of maximal forests of  $G$ . The LP constraint now changes to

$$\sum_{e \in T} x_e \geq k - h, \quad T \in \mathcal{T}(G).$$

**Tree packing interpretation of the dual LP.** The dual LP has two types of variables. For each edge  $e$  there is a variable  $z_e$  and for each spanning tree  $T$  there is a variable  $y_T$ . The dual seeks to add capacity  $z : E \rightarrow \mathbb{R}_+$  to the original capacities  $c$ , and then find a maximum tree packing  $y : \mathcal{T}(G) \rightarrow \mathbb{R}_+$  within the augmented capacities  $c + z$ . The objective is  $(k - 1) \sum_T y_T - \sum_{e \in E} z_e$ . Note that for  $k = 2$ , there is an optimum solution with  $z = 0$ ; this can be seen by the fact that for  $k = 2$  the primal LP can omit the constraints  $x_e \leq 1$  for  $e \in E$ . For  $k > 2$  it may be advantageous to add capacity to some bottleneck edges (say from a minimum cut) to increase the tree packing value, which is multiplied by  $(k - 1)$ .

Our goal is to show that one can transparently carry over the arguments for global minimum cut via tree packings to the  $k$ -CUT setting via (optimum) solutions  $y, z$  to

$$\begin{aligned}
& \max (k-1) \sum_{T \in \mathcal{T}(G)} y_T - \sum_{e \in E} z_e \\
& \text{s.t. } \sum_{T \ni e} y_T \leq c_e + z_e \text{ for all } e \in E \\
& \quad y_T \geq 0 \text{ for all } T \in \mathcal{T}(G)
\end{aligned}$$

FIG. 3. Dual of the LP relaxation from Figure 2.

the dual LP. Theorem 4.1 plays the role of Corollary 2.2. The key lemma below is analogous to Lemma 3.1.

LEMMA 4.3. *Let  $y, z$  be an optimum solution to the dual LP for  $k$ -CUT shown in Figure 3. Let  $E' \subseteq E$  be any subset of edges such that  $c(E') \leq \alpha \lambda_k(G)$  for some  $\alpha \geq 1$ . For each tree  $T$ , let  $\ell_T = |E' \cap E(T)|$  denote the number of edges in both  $T$  and  $E'$ . For an integer  $h \geq (k-1)$ , let  $q_h$  denote the fraction of the trees in the packing induced by  $y, z$  that contain at most  $h$  edges from  $E'$ . Then*

$$q_h \geq 1 - \frac{2\alpha(k-1)(1 - \frac{1}{n})}{h+1}.$$

*Proof.* Let  $\tau_k(G)$  denote  $\sum_T y_T$ , and let  $p_T = y_T / \tau_k(G)$ . Then  $q_h = \sum_{T: \ell_T \leq h} p_T$ .

$$\sum_T p_T \ell_T = \sum_{T: \ell_T \leq h} p_T \ell_T + \sum_{T: \ell_T \geq (h+1)} p_T \ell_T \geq (h+1)(1 - q_h).$$

Because  $y$  is a valid tree packing in capacities  $c + z$ ,

$$\tau_k(G) \sum_T p_T \ell_T = \sum_T y_T \ell_T \leq c(E') + z(E') \leq \alpha \lambda_k(G) + z(E') \leq \alpha \cdot (\lambda_k(G) + z(E)).$$

The last inequality uses the fact that  $\alpha \geq 1$  and  $z \geq 0$ . Putting the preceding inequalities together, we have

$$(h+1)(1 - q_h) \leq \frac{1}{\tau_k(G)} \alpha (\lambda_k(G) + z(E)).$$

We rearrange and simplify the inequality in Corollary 4.2 as

$$2 \left(1 - \frac{1}{n}\right) (k-1) \sum_T y_T \geq \lambda_k(G) + 2 \left(1 - \frac{1}{n}\right) z(E) \geq \lambda_k(G) + z(E).$$

Plugging this inequality into the preceding one yields

$$(h+1)(1 - q_h) \leq 2\alpha(k-1) \left(1 - \frac{1}{n}\right),$$

which implies that

$$q_h \geq 1 - \frac{2\alpha(k-1)(1 - \frac{1}{n})}{h+1}. \quad \square$$

*Remark 2.* Lemma 4.3 does not require  $E'$  to be a  $k$ -cut. Ultimately, we will only apply Lemma 4.3 to  $k$ -cuts.

**COROLLARY 4.4.** *Let  $(y, z)$  be an optimum solution to the dual LP. For every optimum  $k$ -cut  $A \subseteq E$  there is a tree  $T$  in the support of  $y$  such that  $|E(T) \cap A| \leq 2k - 3$ .*

*Proof.* We apply Lemma 4.3 with  $h = 2k - 3$  and  $\alpha = 1$  and observe that  $q_h > 0$ , which implies the desired statement.  $\square$

**COROLLARY 4.5.** *Let  $(y, z)$  be a  $(1 - \epsilon)$ -approximate solution to the dual LP where  $\epsilon < \frac{1}{2k-1}$ . For every optimum  $k$ -cut  $A \subseteq E$  there is a tree  $T$  in the support of  $y$  such that  $|E(T) \cap A| \leq 2k - 2$ .*

*Proof.* Let  $q_h$  be defined as in Lemma 4.3. If  $(y, z)$  is a  $(1 - \epsilon)$ -approximate solution to the dual LP, we would have

$$(4.1) \quad (k - 1) \sum_T y_T \geq (1 - \epsilon) \frac{n\lambda_k(G)}{2(n - 1)} + z(E) \geq (1 - \epsilon) \frac{\lambda_k(G)}{2} + z(E)$$

in place of Corollary 4.2.

Examining the proof of Lemma 4.3, we see that the optimality of  $(y, z)$  is not used in the proof except when invoking Corollary 4.2. Repeating the proof of Lemma 4.3, except using (4.1) instead of Corollary 4.2 and setting  $\alpha = 1$ , we obtain the bound

$$q_h \geq 1 - \frac{2(k - 1)}{(h + 1)(1 - \epsilon)}.$$

We observe that  $q_h > 0$  for  $h = 2k - 2$  and  $\epsilon < \frac{1}{2k-1}$ .  $\square$

**4.1. Number of approximate  $k$ -cuts.** We now prove the following theorem.

**THEOREM 4.6.** *Let  $G = (V, E)$  be an undirected edge-weighted graph, and let  $k$  be a fixed integer. For  $\alpha \geq 1$ , the number of  $\alpha$ -approximate minimum  $k$ -cuts is  $O(n^{\lfloor 2\alpha(k-1) \rfloor})$ .*

*Proof.* Let  $y$  be an optimum solution to the dual LP for  $k$ -CUT. Let  $h = \lfloor 2\alpha(k - 1) \rfloor$ , and let

$$q_h = 1 - \frac{2\alpha(k - 1) \left(1 - \frac{1}{n}\right)}{h + 1} > 0.$$

Fix a  $k$ -cut consisting of a set of edges  $A$  such that  $c(A) \leq \alpha\lambda_k(G)$ . From Lemma 4.3, at least a  $q_h$ -weighted fraction of trees in the packing  $y$  have the property that at most  $h$  edges of  $A$  are in the tree.

For a given tree  $T$ , consider the number of distinct  $k$ -cuts that contain  $h$  or fewer edges in  $T$ . There are at most  $n^h$  subsets of the tree's edges where each subset has cardinality at most  $h$ . Each such subset induces at most  $k^{h+1}$  distinct partitions of the vertex set into  $k$  parts. Thus there are at most  $k^{h+1}n^h$  distinct  $k$ -cuts containing  $h$  or fewer edges from  $T$ .

If there are (strictly) more than  $k^{h+1}n^h/q_h$  distinct cuts with weight at most  $\alpha\lambda_k(G)$ , then by the pigeonhole principle there exists a tree  $T$  in the packing that induces strictly more than  $k^{h+1}n^h$  different cuts with  $h$  or fewer edges—a contradiction. Thus the total number of  $k$ -cuts with cost at most  $\alpha\lambda_k(G)$  is  $O(n^{\lfloor 2\alpha(k-1) \rfloor})$  where the asymptotic notation hides a constant that depends on  $\alpha$  and  $k$ .  $\square$

*Remark 3.* As mentioned earlier, Gupta, Lee, and Li [15] have recently shown that the number of minimum  $k$ -cuts is at most  $n^{1+o(1)k}$ . It is unclear whether their approach yields an improvement for  $\alpha$ -approximate  $k$ -cuts for all values of  $\alpha > 1$  although it seems likely that it would give an improvement for  $\alpha$  close to 1.

**4.2. Enumerating all minimum  $k$ -cuts.** We briefly describe how to enumerate all  $k$ -cuts via Lemma 4.3. The argument is basically the same as that of Karger and Thorup. First, we compute an optimum solution  $(y^*, z^*)$  to the dual LP. We can do this via the Ellipsoid method or other ways. Let  $\beta(n, m)$  be the running time to find  $(y^*, z^*)$ . Moreover, we find a basic feasible solution to the dual LP such that the support of  $y$  has at most  $m$  distinct trees. Now Lemma 4.3 guarantees that for every minimum  $k$ -cut  $A \subseteq E$  there is a tree  $T$  such that  $y_T > 0$  and  $T$   $(2k - 3)$ -respects  $A$ . Thus, to enumerate all minimum  $k$ -cuts the following procedure suffices. For each of the trees  $T$  in the optimum packing we enumerate all  $k$ -cuts induced by removing  $h = 2k - 3$  edges from  $T$ . With appropriate care and data structures (see [20, 36]) this can be done for a single tree  $T$  in  $\tilde{O}(n^{2k-3} + m)$  time. The total time over all  $m$  trees in the support of  $y$  is  $\tilde{O}(mn^{2k-3})$  for  $k > 2$ . This gives the following theorem.

**THEOREM 4.7.** *For  $k > 2$  all of the minimum  $k$ -cuts of a graph with  $n$  nodes and  $m$  edges can be computed in time  $\tilde{O}(mn^{2k-3} + \beta(m, n))$ , where  $\beta(m, n)$  is the time to find an optimum solution to the LP for  $k$ -cut.*

We observe that Thorup's algorithm [36] runs in time  $\tilde{O}(mn^{2k-2})$ . Thorup uses greedy tree packing in place of solving the LP. The optimality of the LP solution was crucial in using the bound of  $2k - 3$  instead of  $2k - 2$ . Thus, even though we obtain a slightly faster algorithm than Thorup, we need to find an optimum solution to the LP which can be done via the Ellipsoid method. The Ellipsoid method is not quite practical. Below we discuss a different approach.

In recent work Quanrud showed that a  $(1 - \epsilon)$ -approximate solution to the dual LP can be computed in near-linear time. We state his theorem below.

**THEOREM 4.8** (see [31]). *There is an algorithm that computes a  $(1 - \epsilon)$ -approximate solution  $(y, z)$  the dual LP in  $O(m \log^3 n / \epsilon^2)$  time.*

We observe that the preceding theorem guarantees  $O(m \log^3 n / \epsilon^2)$  trees in the support of  $y$  and also implicitly stores them in  $O(m \log^3 n / \epsilon^2)$  space. If we choose  $\epsilon < 1/(2k - 1)$ , then, via Corollary 4.5, for every minimum  $k$ -cut  $A \subseteq E$  there is a tree  $T$  in the support of  $y$  that  $(2k - 2)$ -respects  $A$ . This leads to an algorithm that in  $\tilde{O}(mn^{2k-2})$  time enumerates all minimum  $k$ -cuts and recovers Thorup's running time. However, we note that the trees generated by the algorithm in the preceding theorem are implicit, and can be stored in small space. It may be possible to use this additional structure to match or improve the run-time achieved by Theorem 4.7.

*Remark 4.* For unweighted graphs with  $\tilde{O}(\frac{m}{n-k} \frac{1}{\epsilon^2})$  trees [31] guarantees a  $(1 - \epsilon)$ -approximation. This improves the running time to  $\tilde{O}(mn^{2k-3})$  for unweighted graphs.

We briefly discuss a *potential* approach to speed up the computation further. Recall that Karger describes an algorithm that given a spanning tree  $T$  of a graph  $G$  finds the minimum cut that 2-respects  $T$  in  $\tilde{O}(m)$  time, speeding up the easier  $\tilde{O}(n^2)$  time algorithm. We can leverage this as follows. In the case of  $k > 2$  we are given  $T$  and  $G$  and wish to find the minimum  $k$ -cut induced by the removal of at most  $t$  edges where  $t$  is either  $2k - 3$  or  $2k - 2$  depending on the tree packing we use. Suppose  $A$  is a set of  $t - 2$  edges of  $T$ . Removing them from  $T$  yields a forest with  $t - 1$  components. We can then apply Karger's algorithm in each of these components with

an appropriate graph. This results in a running time of  $\tilde{O}(mn^{t-2})$  per tree rather than  $\tilde{O}(n^t)$ . We can try to build on Karger’s ideas to improve the running time to find the best 3-cut induced by removing at most four edges from  $T$ . We could then potentially leverage this for larger values of  $k$ .

**5. A new proof of the LP integrality gap for  $k$ -Cut.** The proof of Theorem 4.1 in [4] is based on the primal-dual algorithm and analysis of Agarwal, Klein, and Ravi [1], and Goemans and Williamson [11] for the Steiner tree problem. For this reason the proof is technical and indirect. Further, the proof from [4] is described for the Steiner  $k$ -cut problem which has additional complexity. Here we give a different and intuitive proof for  $k$ -CUT. Unlike the proof in [4], the proof here relies on optimality properties of the LP solution and hence is less useful algorithmically. We note that [31] uses Theorem 4.8 and a fast implementation of the algorithmic proof in [4] to obtain a near-linear time  $(2 + \epsilon)$ -approximation for  $k$ -CUT.

Let  $G = (V, E)$  be a graph with nonnegative edge capacities  $c_e, e \in E$ . We let  $\text{deg}(v) = \sum_{e \in \delta(v)} c_e$  denote the capacitated degree of node  $v$ . We will assume without loss of generality that  $V = \{v_1, v_2, \dots, v_n\}$  and that the nodes are sorted in increasing order of degrees, that is,  $\text{deg}(v_1) \leq \text{deg}(v_2) \leq \dots \leq \text{deg}(v_n)$ . We observe that  $\text{deg}(v_1) + \text{deg}(v_2) + \dots + \text{deg}(v_{k-1})$  is an upper bound on the value of an optimum  $k$ -CUT; removing all the edges incident to  $v_1, v_2, \dots, v_{k-1}$  gives a feasible solution in which the components are the  $k - 1$  isolated vertices  $\{v_1\}, \{v_2\}, \dots, \{v_{k-1}\}$ , and a component consisting of the remaining nodes of the graph.

The key lemma is the following which proves the integrality gap in a special case.

LEMMA 5.1. *Let  $G$  be a connected graph, and let  $x^*$  be an optimum solution to the  $k$ -CUT LP such that  $x_e^* \in (0, 1)$  for each  $e \in E$  (in other words  $x^*$  is fully fractional). Then  $\sum_{i=1}^{k-1} \text{deg}(v_i) \leq 2(1 - 1/n) \sum_e c_e x_e^*$ .*

*Proof.* Since  $G$  is connected, the LP shown in Figure 2 is a valid relaxation and we can consider its dual. Let  $(y^*, z^*)$  be any fixed optimum solution to the dual LP. Complementary slackness gives the following two properties:

- $z_e^* = 0$  for each  $e \in E$ , for if  $z_e^* > 0$ , we would have  $x_e^* = 1$ .
- For each  $e \in E$ ,  $\sum_{T \ni e} y_T^* = c_e$  since  $x_e^* > 0$ .

From the second property above, and the fact that each spanning tree has exactly  $(n - 1)$  edges, we conclude that

$$(5.1) \quad (n - 1) \sum_T y_T^* = \sum_{e \in E} c_e.$$

Since the degrees are sorted,

$$(5.2) \quad \sum_{i=1}^{k-1} \text{deg}(v_i) \leq \frac{k-1}{n} \sum_{i=1}^n \text{deg}(v_i) = 2 \frac{k-1}{n} \sum_e c_e.$$

Putting the two preceding inequalities together,

$$\sum_{i=1}^{k-1} \text{deg}(v_i) \leq 2 \left(1 - \frac{1}{n}\right) (k-1) \sum_T y_T^* = 2 \left(1 - \frac{1}{n}\right) \sum_e c_e x_e^*,$$

where the last equality is based on strong duality and the fact that  $z^* = 0$ . □

The preceding lemma can be easily generalized to the case when  $G$  has  $h$  connected components following the remark in the preceding section on the  $k$ -CUT LP. This gives us the following.

**COROLLARY 5.2.** *Let  $G$  be a graph with  $h$  connected components, and let  $x^*$  be an optimum solution to the  $k$ -CUT LP such that  $x_e^* \in (0, 1)$  for each  $e \in E$ . Then  $\sum_{i=1}^{k-h} \deg(v_i) \leq 2(1 - 1/n) \sum_e c_e x_e^*$ .*

Now we consider the general case when the optimum solution  $x^*$  to the  $k$ -CUT LP is not necessarily fully fractional as needed in Lemma 5.1. The following claim is easy.

**CLAIM 1.** *Let  $x_e^* = 0$ , where  $e = uv$ . Let  $G'$  be the graph obtained from  $G$  by contracting  $u$  and  $v$  into a single node. Then there is a feasible solution  $x'$  to the  $k$ -CUT LP in  $G'$  of the same cost as that of  $x^*$ . Moreover, a feasible  $k$ -cut in  $G'$  induces a feasible  $k$ -cut in  $G$  of the same cost.*

Using the preceding claim we can assume without loss of generality that  $x_e^* > 0$  for each  $e \in E$ . Let  $F = \{e \in E \mid x_e^* = 1\}$ . Since the LP solution  $x^*$  paid for the full cost of the edges in  $F$ , we can recurse on  $G' = G - F$  and the fractional solution  $x'$  obtained by restricting  $x^*$  to  $E \setminus F$ . If  $G'$  is connected, then  $x'$  is an optimum solution to the  $k$ -CUT LP for  $G'$ , and is fully fractional, and we can apply Lemma 5.1. However,  $G'$  can be disconnected. Let  $h$  be the number of connected components in  $G'$ . If  $h \geq k$ , we are done since  $F$  is a feasible  $k$ -cut and  $c(F) \leq \sum_e c_e x_e^*$ . The interesting case is when  $2 \leq h < k$ . In this case we apply Corollary 5.2 based on the following claim which is intuitive and whose formal proof we omit.

**CLAIM 2.** *Let  $x'$  be the restriction of  $x^*$  to  $E \setminus F$ . Then for any maximal forest  $T$  in  $G'$  we have  $\sum_{e \in T} x'(e) \geq k - h$ . Moreover,  $x'$  is an optimum solution to the  $k$ -CUT LP for  $G'$ .*

From Corollary 5.2 we can find  $E' \subset E \setminus F$  such that  $G' - E'$  induces a  $k$ -cut in  $G'$  such that

$$c(E') \leq 2 \left(1 - \frac{1}{n}\right) \sum_{e \in E \setminus F} c_e x'_e = 2 \left(1 - \frac{1}{n}\right) \sum_{e \in E \setminus F} c_e x_e^*.$$

Therefore,  $F \cup E'$  is a  $k$ -cut in  $G$  and we have that

$$c(F \cup E') = c(F) + c(E') \leq \sum_{e \in F} c_e x_e^* + 2 \left(1 - \frac{1}{n}\right) \sum_{e \in E \setminus F} c_e x_e^* \leq 2 \left(1 - \frac{1}{n}\right) \sum_{e \in E} c_e x_e^*.$$

This finishes the proof. Note that the proof also gives a very simple rounding algorithm assuming we have an optimum solution  $x^*$  for the LP. Contract all edges with  $x_e^* = 0$ , remove all edges  $e$  with  $x_e^* = 1$ , and use Lemma 5.1 in the residual graph to find the  $(k - h)$  smallest degree vertices.

**An LP-based proof of Theorem 2.1.** The preceding proof idea also yields a proof of Theorem 2.1, which we sketch here. Recall that  $\tau(G)$  is the optimum solution value to the tree packing LP, which corresponds to the dual of the  $k$ -CUT LP when  $k = 2$ . We had already observed that  $\tau(G)$  is at most the strength of the graph. Here we will prove equality. When  $k = 2$ , as we remarked, the LP does not require the upper bound constraints  $x(e) \leq 1$  which implies that the dual tree packing LP does not have the  $z$  variables. Following Lemma 5.1 we consider a fully fractional optimum solution  $x^*$  to the  $k$ -CUT LP with  $k = 2$  and an optimum dual solution  $y^*$  to the dual tree packing LP. We have

$$(n - 1) \sum_T y_T^* = (n - 1) \tau(G) = \sum_{e \in E} c_e.$$

Consider the partition  $\mathcal{P}$  consisting of all the singleton vertices; all edges cross this partition, and hence  $c(\mathcal{P}) = \sum_e c_e$  and  $|\mathcal{P}| = n$ . Since  $\tau(G) = \frac{\sum_e c_e}{n-1} = \frac{c(\mathcal{P})}{|\mathcal{P}|-1}$  it must be the case that  $\frac{\sum_e c_e}{n-1}$  is the network strength which equals the tree packing value. When  $x^*$  is not fully fractional we can contract edges with  $x_e^* = 0$  and apply the preceding argument. A similar argument can be used to prove the corresponding min-max relation for the fractional packing of bases of a matroid.

**6. Characterizing the optimum LP solution.** We have seen that the dual of the LP relaxation for  $k$ -CUT yields a tree packing that can be used in place of Thorup’s recursive tree packing. In this section we show that the two are the same by characterizing the optimum LP solution for a given graph through a recursive partitioning procedure. This yields a nested sequence of partitions of the vertex set of the graph. This sequence is called the principal sequence of partitions of a graph and is better understood in the more general context of submodular functions [30]. We refer the reader to Fujishige’s article for more on this topic [9], and to [7, 23] for algorithmic aspects in the setting of graphs. We also connect the LP relaxation with the Lagrangean relaxation approach for  $k$ -CUT considered by Barahona [2] and Ravi and Sinha [32]. Their approach is also built upon the principal sequence of partitions. In order to keep the discussion simple we mainly follow the notation and approach of [32].

Given  $G = (V, E)$  and an edge set  $A \subseteq E$ , let  $\kappa(A)$  denote the number of connected components in  $G - A$ . Recall that the strength of a capacitated graph  $G$ , denoted by  $\sigma(G)$  is defined as  $\min_{A \subseteq E} \frac{c(A)}{\kappa(A)-1}$ . The  $k$ -CUT problem can be phrased as  $\min_{A: \kappa(A) \geq k} c(A)$ . The constraint is that  $\kappa(A) \geq k$  is not straightforward to work with although  $\kappa(A)$  is a supermodular set function over the ground set  $E$ . A Lagrangean relaxation approach was considered in [2, 32]. To set this up we define, for any fixed edge set  $A$ , a function  $g_A : \mathbb{R}_+ \rightarrow \mathbb{R}$  as  $g_A(b) = c(A) - b(\kappa(A) - 1)$ . We then obtain the function  $g : \mathbb{R}_+ \rightarrow \mathbb{R}$  where  $g(b) = \min_{A \subseteq E} c(A) - b(\kappa(A) - 1)$ . The quantity  $g(b)$  is the attack value of the graph for parameter  $b$  and was considered by Cunningham [7] in his algorithm to compute the strength of the graph. (Note that  $g(b)$  for any fixed value of  $b$  can be computed in polynomial-time via submodular function minimization.) Then, as noted in [2, 32],

$$\min_{A: \kappa(A) \geq k} c(A) \geq \max_{b \geq 0} \min_{A \subseteq E} c(A) + b(k - \kappa(A)) = \max_{b \geq 0} g(b) + b(k - 1).$$

Consider the function  $g' : \mathbb{R}_+ \rightarrow \mathbb{R}$  where  $g'(b) = g(b) + b(k - 1)$ . From the preceding discussion, for any  $b \geq 0$ ,  $g'(b)$  is a lower bound on the minimum  $k$ -CUT value. [32] describes structural properties of the function  $g$ , several of which are explicit or implicit in [7]. We state them below.

- The functions  $g$  and  $g'$  are continuous, concave, and piecewise linear and have no more than  $n - 1$  breakpoints. The function  $g$  is nonincreasing in  $b$ .
- Under a nondegeneracy assumption on the graph, which is easy to ensure, the following holds. If  $b$  is not a breakpoint, then there is a unique edge set  $A$  such that  $g_A(b) = g(b)$ . If  $b$  is a breakpoint, then there are exactly two distinct edge sets  $A, B$  such that  $g_A(b) = g_B(b) = g(b)$ .
- If  $b_0$  is a breakpoint of  $g'$  induced by edge sets  $A$  and  $B$  with  $\kappa(A) > \kappa(B)$ , then  $B \subset A$ . In particular  $A \setminus B$  is contained in some connected component of  $G' = (V, E \setminus B)$ .
- Let  $b_0$  be a breakpoint of  $g'$  induced by edge set  $A$ . Then the next breakpoint is induced by the edge set which is the solution to the strength problem on

the smallest strength component of  $G' = (V, E \setminus A)$ .

The above properties show that the breakpoints induce a sequence of partitions of  $V$  which are refinements. Alternatively, we consider the sequence of edge sets  $A_1, A_2, \dots$ , obtained by the following algorithm. We will assume that  $G$  is connected. Let  $A_0 = \emptyset$ , and let  $\mathcal{P}_0 = \{V\}$  which consists of a trivial partition of  $V$  induced by removing  $A_0$ . Given  $A_{i-1}$  we obtain  $A_i \supseteq A_{i-1}$  as follows. Let  $G_i = (V, E \setminus A_{i-1})$ . If  $G_i$  has no edges, we stop. Otherwise, let  $C_i$  be the minimum strength connected component of  $G_i$  and  $B_i$  be a maximal minimum strength edge set of  $C_i$ . We define  $A_i = A_{i-1} \cup B_i$ , and  $\tau_i$  to be the strength of the component  $C_i$ . That is,  $\tau_i = \frac{c(A_i) - c(A_{i-1})}{\kappa(A_i) - \kappa(A_{i-1})} = \frac{c(B_i)}{\kappa(A_i) - \kappa(A_{i-1})}$ . The process stops when  $A_h = E$ . Let  $\mathcal{P}_i$  denote the partition of  $V$  induced by  $A_i$ . Note that  $\mathcal{P}_i$  is obtained from  $\mathcal{P}_{i-1}$  by replacing  $C_i$  by a minimum strength partition of  $C_i$ , thus  $\mathcal{P}_i$  is a refinement of  $\mathcal{P}_{i-1}$  and  $\mathcal{P}_h$  consists of singleton nodes. Note that Thorup's ideal tree packing is also based on the same recursive decomposition. Let the  $i$ th breakpoint of  $g'$  to be  $b_i$ . It was shown that  $b_i = \tau_i$  is precisely the  $i$ th breakpoint of the function  $g'$  [32].

Ravi and Sinha obtained a 2-approximation for  $k$ -CUT as follows. Given the preceding decomposition of  $G$ , they consider the smallest  $j$  such that  $|\mathcal{P}_j| \geq k$ . If  $|\mathcal{P}_j| = k$ , they output it and can argue that it is an optimum solution. Otherwise, they do the following. Recall  $\mathcal{P}_j$  is obtained from  $\mathcal{P}_{j-1}$  by replacing the component  $C_j$  in  $G - A_{j-1}$  by a minimum strength decomposition of  $C_j$ . Let  $k' = k - |\mathcal{P}_{j-1}|$ . Consider the minimum strength partition of  $C_j$ , and let  $H_1, H_2, \dots, H_r$  be the connected components of this partition ordered such that  $|\delta_{C_j}(H_1)| \leq |\delta_{C_j}(H_2)| \leq \dots \leq |\delta_{C_j}(H_r)|$ . Note that  $r \geq k'$ . Output the cut  $A_{j-1} \cup (\cup_{\ell=1}^{k'} \delta(H_\ell))$ .

**6.1. An optimum LP solution from the decomposition.** Given  $k$ , as before let  $j$  be the smallest index such that  $\kappa(A_j) \geq k$ . We consider the following solution to the LP:

- $x_e = 1$  for each  $e \in A_{j-1}$ .
- $x_e = \alpha$  for each  $e \in A_j \setminus A_{j-1}$ , where  $\alpha = \frac{k - \kappa(A_{j-1})}{\kappa(A_j) - \kappa(A_{j-1})}$ .
- $x_e = 0$  for each  $e \in E \setminus A_j$ .

LEMMA 6.1. *The solution  $x$  is feasible and has objective value*

$$c(A_{j-1}) + \alpha c(B_j) = c(A_{j-1}) + (k - \kappa(A_{j-1})) b_j.$$

*Proof.* Let  $T$  be any spanning tree. We want to show that  $\sum_{e \in T} x_e \geq k - 1$ . For each  $j$ , let  $\kappa_j = \kappa(A_j)$ , and let  $\ell_j = |T \cap A_j|$ . Then  $T$  has  $\ell_{j-1}$  edges of value  $x_e = 1$ , and  $\ell_j - \ell_{j-1}$  edges of value  $\alpha$ . We have

$$\begin{aligned} \sum_{e \in T} x_e &= \ell_{j-1} + (\ell_j - \ell_{j-1})\alpha \\ &\geq \kappa_{j-1} - 1 + (\kappa_j - \kappa_{j-1})\alpha = k - 1, \end{aligned}$$

where we observe that the right-hand side of the first line is decreasing in both  $\ell_j$  and  $\ell_{j-1}$ ,  $\ell_j \geq \kappa_j - 1$ , and  $\ell_{j-1} \geq \kappa_{j-1} - 1$ . Recall that  $B_j = A_j \setminus A_{j-1}$  and that  $b_j = \tau_j = \frac{c(B_j)}{\kappa_j - \kappa_{j-1}}$ . Therefore,  $\alpha c(B_j) = (k - \kappa_{j-1})b_j$ . To calculate the objective value of  $x$ , we have

$$\begin{aligned} \sum_{e \in E} c_e x_e &= \sum_{e \in A_{j-1}} c_e + \sum_{e \in A_j \setminus A_{j-1}} \alpha c_e = c(A_{j-1}) + \alpha c(B_j) \\ &= c(A_{j-1}) + (k - \kappa_{j-1})b_j. \quad \square \end{aligned}$$

The harder part is captured by the next lemma.

LEMMA 6.2. *The solution  $x$  attains the optimum value to the LP relaxation.*

*Proof.* We prove the claim by constructing a dual solution of equal value. See Figure 3 for the dual LP.

Recall the definitions of  $\mathcal{P}_i$ ,  $A_i$ ,  $B_i$ , and  $C_i$  from earlier. For each  $i$ , let  $\kappa_i = \kappa(A_i) = |\mathcal{P}_i|$  be the number of components in the  $i$ th partition. Recall that  $b_i$  is a breakpoint of  $g'$  and  $b_i = \tau_i$  is the strength of  $C_i$ . Let  $Q_i$  be the partition of the component  $C_i$  obtained by removing  $B_i$  from the graph  $G[C_i]$ . An ideal tree packing, following [37], is a convex combination of trees  $p : \mathcal{T}(G) \rightarrow [0, 1]$  s.t.  $\sum_T p_T = 1$  with the following properties.

1. For each  $i$ , every tree  $T$  supported by  $p$  induces a tree in the graph  $G/\mathcal{P}_i$  obtained by contracting each component of  $\mathcal{P}_i$ .
2. For each  $i$  and each edge  $e \in B_i$ ,  $\sum_{T \ni e} p_T = c_e/b_i$ .

One can interpret  $p$  as a probability distribution over spanning trees of  $G$ . Every graph has an ideal tree packing, and can be constructed recursively as follows. If  $G$  consists of a single vertex  $v$ , then there is only one spanning tree consisting of the vertex  $v$  and is picked with probability 1, and the desired properties are easily satisfied. For the general case we construct the distribution by following the partitioning process. Recall that  $G_1 = G$  and  $B_1 \subseteq E$  is the edge set that induces a minimum strength partitioning of  $G_1$ , and  $Q_1$  is a partition of  $C_1 = V$  induced by removing the edge set  $B_1$ ; note that  $Q_1 = \mathcal{P}_1$ . Let  $H$  be the graph obtained by shrinking each part of  $Q_1 = \mathcal{P}_1$  into a single vertex; the edge set of  $H$  is  $B_1$  and  $B_1$  is also a minimum strength edge partitioning of  $H$  and strength value is equal to  $b_1$ . Consider a maximum tree packing in  $H$ , and let  $y_T \geq 0, T \in \mathcal{T}(H)$ , denote this packing. We have that  $\sum_T y_T = c(B_1)/(|\mathcal{P}_1| - 1) = b_1$  and every edge  $e \in B_1$  is saturated by the packing, that is,  $\sum_{T \ni e} y_T = c_e$ . We can obtain a probability distribution  $p'$  over trees  $\mathcal{T}(H)$  by setting  $p'_T = y_T/b_1$  (since  $\sum_T y_T = b_1$ ). If a tree  $T$  is chosen according to this distribution  $p'$ , then the probability that an edge  $e$  from  $B_1$  is in the chosen tree is equal to  $c(e)/b_1$ . To obtain the desired probability distribution for the spanning trees of the entire graph  $G$ , we independently and recursively pick one tree  $T_j$  from each graph  $G[V_j]$  where  $V_j$  is a part in  $\mathcal{P}_1$ , and concatenate the edges of these chosen trees with a tree chosen independently from the distribution  $p'$  we described for  $H$ . This results in a distribution  $p$  over the spanning trees of  $G$ , and it can be verified that this is the desired ideal tree packing.

Let  $p : \mathcal{T}(G) \rightarrow [0, 1]$  be an ideal tree packing. Recall that  $j$  is the smallest index such that  $|\mathcal{P}_j| \geq k$ . To construct our dual solution, we define nonnegative edge potentials  $z_e \geq 0$  and a tree packing  $y(t) \geq 0$  (packing into  $c + z$ ) s.t.

$$y_T = b_j p_T \quad \text{for all } T \in \mathcal{T}(G),$$

$$c_e + z_e = \begin{cases} \frac{b_j}{b_i} c_e & \text{for all } e \in B_i \text{ for } i < j, \\ c_e & \text{otherwise.} \end{cases}$$

We first claim that  $(y, z)$  is feasible in the dual LP; that is,  $y$  is a feasible tree packing with respect to the augmented capacities  $c + z$ . Observe that for any edge  $e$ ,  $y$  uses capacity  $b_j$  times the capacity by  $p$ . We need to show the capacity used by  $y$  along any edge  $e$  is at most  $c_e + z_e$ . We have two cases.

1. If  $e \in B_i$  for some  $i < j$ , then  $p$  uses capacity  $\frac{c_e}{b_i}$ . In turn,  $y$  uses capacity  $\frac{b_j}{b_i} c_e$ . By choice of  $z_e$ , we have  $c_e + z_e = \frac{b_j}{b_i} c_e$ , as desired.

2. If  $e \in E \setminus A_{j-1}$ , then  $p$  uses capacity at most  $\frac{c_e}{b_i}$ . In turn,  $y$  uses capacity at most  $\frac{b_i}{b_i} c_e$ . But  $b_j \leq b_i$ , so the capacity used by  $y$  is  $\leq c_e$ .

We now analyze the objective value of our dual solution. We first observe that since each tree supported by  $y$  is a tree in  $G/\mathcal{P}_j$ , we have

$$\begin{aligned} (k-1) \sum_T y_T &= \frac{k-1}{\kappa_j-1} \sum_T y_T |T \cap A_j| = \frac{k-1}{\kappa_j-1} \sum_{e \in A_j} \sum_{T \ni e} y_T \\ &= \frac{k-1}{\kappa_j-1} b_j \sum_{i \leq j} \frac{1}{b_i} \sum_{e \in B_i} c_e = \frac{k-1}{\kappa_j-1} b_j \sum_{i \leq j} (\kappa_i - \kappa_{i-1}) \\ &= \frac{k-1}{\kappa_j-1} b_j (\kappa_j - 1) = (k-1) b_j. \end{aligned}$$

On the other hand, when subtracting the augmented capacities, we have

$$\begin{aligned} \sum_{e \in E} z_e &= \sum_{i < j} \sum_{e \in B_i} \left( \frac{b_j}{b_i} - 1 \right) c_e = b_j \left( \sum_{i < j} \frac{1}{b_i} \sum_{e \in B_i} c_e \right) - \sum_{e \in A_{j-1}} c_e \\ &= b_j \sum_{i < j} (\kappa_i - \kappa_{i-1}) - \sum_{e \in A_{j-1}} c_e = b_j (\kappa_{j-1} - 1) - \sum_{e \in A_{j-1}} c_e. \end{aligned}$$

Thus the total objective value of our solution, as a function of  $b_j$ , is

$$(k-1) \sum_T y_T - \sum_{e \in E} z_e = (k - \kappa_{j-1}) b_j + \sum_{e \in A_{j-1}} c_e,$$

as desired. □

*Remark 5.* One can also verify the optimality of  $(x, y, z)$  in the proof above by complimentary slackness conditions. Recall that  $x$  and  $(y, z)$  satisfy the complimentary slackness conditions if

1.  $z_e > 0$  only if  $x_e = 1$ ,
2.  $y_T > 0$  only if  $\sum_{e \in T} x_e = k - 1$ ,
3.  $x_e > 0$  only if  $\sum_{T \ni e} y_e = c_e + z_e$ .

We address these individually.

1.  $z_e > 0$  only if  $e \in B_i$  for some  $i < j$ . In this case,  $e \in A_{j-1}$  so  $x_e = 1$ .
2. If  $y_T > 0$ , then  $T$  is in the support of the ideal tree packing. In particular,  $T$  contains exactly  $\kappa_j - 1$  edges from  $A_j$  and  $\kappa_j - 1$  edges from  $A_{j-1}$ , so we have

$$\sum_{e \in T} x_e = \sum_{e \in T \cap A_{j-1}} 1 + \sum_{e \in T \cap (A_j \setminus A_{j-1})} \frac{k - \kappa_{j-1}}{\kappa_j - \kappa_{j-1}} = \kappa_{j-1} - 1 + k - \kappa_{j-1} = k - 1,$$

as desired.

3. If  $x_e > 0$ , then  $e \in B_i$  for some  $i \leq j$ , so  $y$  uses  $\frac{b_i}{b_i} c_e = c_e + z_e$  units of capacity of  $e$ , as desired.

**6.2. Implications of the characterization.** We now outline some implications of the preceding characterization of the optimum LP solution.

Ravi and Sinha showed that Lagrangean relaxation lower bound is no weaker than the one provided by LP relaxation. Here we show that they are equivalent.

**THEOREM 6.3.** *The Lagrangean relaxation value is the same as the LP value.*

*Proof.* Let  $\kappa_i = \kappa(A_i) = |\mathcal{P}_i|$  be the number of components after removing  $A_i$ . If  $\kappa_j = k$  for some  $j$ , then the Lagrangean relaxation value is the min  $k$ -cut value. Hence it matches the LP value. Otherwise, assume  $\kappa_{j-1} < k < \kappa_j$ . Since  $g'$  is concave, continuous, and piecewise linear, one can see that the function  $g'$  maximizes at one of the breakpoints, and it is precisely  $b_j$ . Indeed,  $\kappa_j > k$ , so  $b_{j+1}(k - \kappa_j) < 0$  is a negative slope. We have

$$g'(b_j) = c(A_{j-1}) - b_j(\kappa_{j-1} - 1) + b_j(k - 1) = c(A_{j-1}) + (k - \kappa_{j-1})b_j.$$

This is precisely the value of the LP in Lemma 6.1.  $\square$

The preceding also gives yet another proof that the integrality gap of the LP is  $2(1 - 1/n)$ .

Second, as we saw, for any value of  $k$ , an optimum dual solution to the  $k$ -CUT LP can be derived from the ideal tree packing [36, 37]. The last issue is the connection between greedy tree packing and the dual LP. At the high-level it is tempting to conjecture that greedy tree packing is essentially approximating the dual LP via the standard multiplicative weight update (MWU) approach. Proving the conjecture formally may require a fair amount of technical work and we leave it for future work. We believe that some insights obtained in [31] could be useful in this context; [31] recasts the LP relaxation for  $k$ -CUT into a pure covering LP, and the dual as a pure packing LP that packs forests instead of trees.

**Acknowledgments.** We thank Karthik Chandrasekaran for helpful comments on this paper. We also thank several reviewers for comments and corrections that improved the presentation and clarity of this paper.

#### REFERENCES

- [1] A. AGRAWAL, P. KLEIN, AND R. RAVI, *When trees collide: An approximation algorithm for the generalized Steiner problem on networks*, SIAM J. Comput., 24 (1995), pp. 440–456, <https://doi.org/10.1137/S0097539792236237>.
- [2] F. BARAHONA, *On the  $k$ -cut problem*, Oper. Res. Lett., 26 (2000), pp. 99–105.
- [3] K. CHANDRASEKARAN, C. XU, AND X. YU, *Hypergraph  $k$ -cut in randomized polynomial time*, in Proceedings of the Twenty-Ninth Annual ACM-SIAM Symposium on Discrete Algorithms, SIAM, Philadelphia, 2018, pp. 1426–1438.
- [4] C. CHEKURI, S. GUHA, AND J. NAOR, *The Steiner  $k$ -cut problem*, SIAM J. Discrete Math., 20 (2006), pp. 261–271, <https://doi.org/10.1137/S0895480104445095>.
- [5] C. CHEKURI AND S. LI, *A Note on the Hardness of the  $k$ -way Hypergraph Cut Problem*, unpublished manuscript available at <http://chekuri.cs.illinois.edu/papers/hypergraph-kcut.pdf>, 2015.
- [6] C. CHEKURI AND K. QUANRUD, *Near-linear time approximation schemes for some implicit fractional packing problems*, in Proceedings of the Twenty-Eighth Annual ACM-SIAM Symposium on Discrete Algorithms, SIAM, Philadelphia, 2017, pp. 801–820, <https://doi.org/10.1137/1.9781611974782.51>.
- [7] W. H. CUNNINGHAM, *Optimal attack and reinforcement of a network*, J. ACM, 32 (1985), pp. 549–561.
- [8] R. G. DOWNEY, V. ESTIVILL-CASTRO, M. R. FELLOWS, E. PRIETO-RODRIGUEZ, AND F. A. ROSAMOND, *Cutting up is hard to do: The parameterized complexity of  $k$ -cut and related problems*, Electr. Notes Theor. Comput. Sci., 78 (2003), pp. 209–222.
- [9] S. FUJISHIGE, *Theory of principal partitions revisited*, in Research Trends in Combinatorial Optimization, Springer, Berlin, 2009, pp. 127–162.
- [10] H. N. GABOW AND K. S. MANU, *Packing algorithms for arborescences (and spanning trees) in capacitated graphs*, Math. Programming, 82 (1998), pp. 83–109.

- [11] M. X. GOEMANS AND D. P. WILLIAMSON, *A general approximation technique for constrained forest problems*, SIAM J. Comput., 24 (1995), pp. 296–317, <https://doi.org/10.1137/S0097539793242618>.
- [12] O. GOLDSCHMIDT AND D. HOCHBAUM, *A polynomial algorithm for the  $k$ -cut problem for fixed  $k$* , Math. Oper. Res., 19 (1994), pp. 24–37.
- [13] A. GUPTA, E. LEE, AND J. LI, *Faster exact and approximate algorithms for  $k$ -cut*, in Proceedings of the 59th Annual IEEE Symposium on Foundations of Computer Science—FOCS, IEEE Computer Soc., Los Alamitos, CA, 2018, pp. 113–123.
- [14] A. GUPTA, E. LEE, AND J. LI, *An FPT algorithm beating 2-approximation for  $k$ -cut*, in Proceedings of the Twenty-Ninth Annual ACM-SIAM Symposium on Discrete Algorithms, SODA 2018, New Orleans, LA, 2018, A. Czumaj, ed., SIAM, Philadelphia, 2018, pp. 2821–2837, <https://doi.org/10.1137/1.9781611975031.179>.
- [15] A. GUPTA, E. LEE, AND J. LI, *The Karger-Stein Algorithm is Optimal for  $k$ -cut*, preprint, <https://arxiv.org/abs/1911.09165>, 2019, to appear in ACM STOC 2020.
- [16] A. GUPTA, E. LEE, AND J. LI, *The number of minimum  $k$ -cuts: Improving the Karger-Stein bound*, in Proceedings of the 51st Annual ACM SIGACT Symposium on Theory of Computing, ACM, 2019, pp. 229–240.
- [17] M. HENZINGER, S. RAO, AND D. WANG, *Local flow partitioning for faster edge connectivity*, in Proceedings of the Twenty-Eighth Annual ACM-SIAM Symposium on Discrete Algorithms, SODA, P. N. Klein, ed., SIAM, Philadelphia, 2017, pp. 1919–1938, <https://doi.org/10.1137/1.9781611974782.125>.
- [18] Y. KAMIDOI, N. YOSHIDA, AND H. NAGAMUCHI, *A deterministic algorithm for finding all minimum  $k$ -way cuts*, SIAM J. Comput., 36 (2007), pp. 1329–1341, <https://doi.org/10.1137/050631616>.
- [19] D. R. KARGER, *Random Sampling in Graph Optimization Problems*, Ph.D. thesis, Stanford University, Stanford, CA, 1995.
- [20] D. R. KARGER, *Minimum cuts in near-linear time*, J. ACM, 47 (2000), pp. 46–76, <https://doi.org/10.1145/331605.331608>.
- [21] D. R. KARGER AND C. STEIN, *A new approach to the minimum cut problem*, J. ACM, 43 (1996), pp. 601–640.
- [22] K. KAWARABAYASHI AND M. THORUP, *Deterministic global minimum cut of a simple graph in near-linear time*, in Proceedings of the Forty-Seventh Annual ACM on Symposium on Theory of Computing, STOC 2015, Portland, OR, 2015, R. A. Servedio and R. Rubinfeld, eds., ACM, 2015, pp. 665–674.
- [23] V. KOLMOGOROV, *A faster algorithm for computing the principal sequence of partitions of a graph*, Algorithmica, 56 (2010), pp. 394–412.
- [24] J. LI, *Faster minimum  $k$ -cut of a simple graph*, in Proceedings of the 60th Annual IEEE Symposium on Foundations of Computer Science, 2019, pp. 1056–1077.
- [25] P. MANURANGSI, *Almost-polynomial ratio  $\epsilon$ -hardness of approximating densest  $k$ -subgraph*, in Proceedings of the 49th Annual ACM SIGACT Symposium on Theory of Computing, STOC 2017, Montreal, QC, Canada, 2017, pp. 954–961, <https://doi.org/10.1145/3055399.3055412>.
- [26] P. MANURANGSI, *Inapproximability of maximum edge biclique, maximum balanced biclique and minimum  $k$ -cut from the small set expansion hypothesis*, in Proceedings of ICALP, LIPIcs 80, Schloss Dagstuhl-Leibniz-Zentrum fuer Informatik, 2017, 79.
- [27] H. NAGAMUCHI AND T. IBARAKI, *A linear-time algorithm for finding a sparse  $k$ -connected spanning subgraph of a  $k$ -connected graph*, Algorithmica, 7 (1992), pp. 583–596, <https://doi.org/10.1007/BF01758778>.
- [28] H. NAGAMUCHI AND Y. KAMIDOI, *Minimum cost subpartitions in graphs*, Inform. Process. Lett., 102 (2007), pp. 79–84, <https://doi.org/10.1016/j.ipl.2006.11.011>.
- [29] J. NAOR AND Y. RABANI, *Tree packing and approximating  $k$ -cuts*, in Proceedings of the Twelfth Annual ACM-SIAM Symposium on Discrete Algorithms, vol. 103, SIAM, Philadelphia, 2001, pp. 26–27.
- [30] H. NARAYANAN, *The principal lattice of partitions of a submodular function*, Linear Algebra Appl., 144 (1991), pp. 179–216.
- [31] K. QUANRUD, *Fast and Deterministic Approximations for  $k$ -cut*, preprint, <https://arxiv.org/abs/1807.07143>, 2018, in Proceedings of APPROX 2019.
- [32] R. RAVI AND A. SINHA, *Approximating  $k$ -cuts using network strength as a Lagrangean relaxation*, European J. Oper. Res., 186 (2008), pp. 77–90.
- [33] H. SARAN AND V. V. VAZIRANI, *Finding  $k$ -cuts within twice the optimal*, SIAM J. Comput., 24 (1995), pp. 101–108, <https://doi.org/10.1137/S0097539792251730>.

- [34] A. SCHRIJVER, *Combinatorial Optimization: Polyhedra and Efficiency*, Algorithms and Combinatorics 24, Springer-Verlag, Berlin, 2003.
- [35] M. STOER AND F. WAGNER, *A simple min-cut algorithm*, J. ACM, 44 (1997), pp. 585–591, <https://doi.org/10.1145/263867.263872>.
- [36] M. THORUP, *Fully-dynamic min-cut*, Combinatorica, 27 (2007), pp. 91–127.
- [37] M. THORUP, *Minimum  $k$ -way cuts via deterministic greedy tree packing*, in Proceedings of the Fortieth Annual ACM Symposium on Theory of Computing, ACM, 2008, pp. 159–166.