



ℓ_p -Norm Multiway Cut

Karthekeyan Chandrasekaran¹ · Weihang Wang¹

Received: 27 October 2021 / Accepted: 28 April 2022

© The Author(s), under exclusive licence to Springer Science+Business Media, LLC, part of Springer Nature 2022

Abstract

We introduce and study ℓ_p -NORM-MULTIWAY-CUT: the input here is an undirected graph with non-negative edge weights along with k terminals and the goal is to find a partition of the vertex set into k parts each containing exactly one terminal so as to minimize the ℓ_p -norm of the cut values of the parts. This is a unified generalization of min-sum multiway cut (when $p = 1$) and min-max multiway cut (when $p = \infty$), both of which are well-studied classic problems in the graph partitioning literature. We show that ℓ_p -NORM-MULTIWAY-CUT is NP-hard for constant number of terminals and is NP-hard in planar graphs. On the algorithmic side, we design an $O(\log^{1.5} n \log^{0.5} k)$ -approximation for all $p \geq 1$. We also show an integrality gap of $\Omega(k^{1-1/p})$ for a natural convex program and an $O(k^{1-1/p-\epsilon})$ -inapproximability for any constant $\epsilon > 0$ assuming the small set expansion hypothesis.

Keywords Multiway cut · Approximation algorithms

1 Introduction

MULTIWAY-CUT is a fundamental problem in combinatorial optimization with both theoretical as well as practical motivations. The input here is an undirected graph $G = (V, E)$ with non-negative edge weights $w : E \rightarrow \mathbb{Q}_+$ along with k specified terminals $T = \{t_1, t_2, \dots, t_k\} \subseteq V$. The goal is to find a partition $\mathcal{P} = (P_1, P_2, \dots, P_k)$ of the vertex set with $t_i \in P_i$ for each $i \in [k]$ so as to minimize the sum of the cut values of the parts, i.e., the objective is to minimize $\sum_{i=1}^k w(\delta(P_i))$, where $\delta(P_i)$ denotes the set of edges with exactly one end-vertex in P_i and $w(\delta(P_i)) := \sum_{e \in \delta(P_i)} w(e)$. On

Supported in part by NSF Grants CCF-1814613 and CCF-1907937. A preliminary version of this work appeared at European Symposium on Algorithms (ESA), 2021.

✉ Weihang Wang
weihang3@illinois.edu

Karthekeyan Chandrasekaran
karthe@illinois.edu

¹ University of Illinois, Urbana-Champaign, Champaign, USA

the practical side, MULTIWAY-CUT has been used to model file-storage in networks as well as partitioning circuit elements among chips—see [13, 22]. On the theoretical side, MULTIWAY-CUT generalizes the min (s, t) -cut problem which is polynomial-time solvable. In contrast to min (s, t) -cut, MULTIWAY-CUT is NP-hard for $k \geq 3$ terminals [13]. The algorithmic study of MULTIWAY-CUT has led to groundbreaking rounding techniques and integrality gap constructions in the field of approximation algorithms [2, 4–7, 11, 16, 17, 21] and novel graph structural techniques in the field of fixed-parameter algorithms [18]. It is known that MULTIWAY-CUT does not admit a $(1.20016 - \epsilon)$ -approximation for any constant $\epsilon > 0$ assuming the Unique Games Conjecture [4] and the current best upper bound on the approximation factor is 1.2965 [21].

Motivated by its connections to partitioning and clustering, Svitkina and Tardos [22] introduced a local part-wise min–max objective for MULTIWAY-CUT—we will denote this problem as MIN-MAX-MULTIWAY-CUT: The input here is the same as MULTIWAY-CUT while the goal is to find a partition $\mathcal{P} = (P_1, P_2, \dots, P_k)$ of the vertex set with $t_i \in P_i$ for each $i \in [k]$ so as to minimize $\max_{i=1}^k w(\delta(P_i))$. We note that MULTIWAY-CUT and MIN-MAX-MULTIWAY-CUT differ only in the objective function—the objective function in MULTIWAY-CUT is to minimize the sum of the cut values of the parts while the objective function in MIN-MAX-MULTIWAY-CUT is to minimize the max of the cut values of the parts. MIN-MAX-MULTIWAY-CUT can be viewed as a fairness inducing multiway cut as it aims to ensure that no part pays too much in cut value. Svitkina and Tardos showed that MIN-MAX-MULTIWAY-CUT is NP-hard for $k \geq 4$ terminals and also that it admits an $O(\log^3 n)$ -approximation.¹ Bansal, Feige, Krauthgamer, Makarychev, Nagarajan, Naor, and Schwartz subsequently improved the approximation factor to $O(\sqrt{\log n \log k})$ (which is $O(\log n)$) [3].

In this work, we study a unified generalization of MULTIWAY-CUT and MIN-MAX-MULTIWAY-CUT that we term as ℓ_p -NORM-MULTIWAY-CUT: In this problem, the input is the same as MULTIWAY-CUT, i.e., we are given an undirected graph $G = (V, E)$ with non-negative edge weights $w : E \rightarrow \mathbb{Q}_+$ along with k specified terminal vertices $T = \{t_1, t_2, \dots, t_k\} \subseteq V$. The goal is to find a partition $\mathcal{P} = (P_1, P_2, \dots, P_k)$ of the vertex set with $t_i \in P_i$ for each $i \in [k]$ so as to minimize the ℓ_p -norm of the cut values of the k parts—formally, we would like to minimize

$$\left(\sum_{i=1}^k \left(\sum_{e \in \delta(P_i)} w(e) \right)^p \right)^{\frac{1}{p}}.$$

Throughout, we will consider $p \geq 1$. We note that ℓ_p -NORM-MULTIWAY-CUT for $p = 1$ corresponds to MULTIWAY-CUT and for $p = \infty$ corresponds to MIN-MAX-MULTIWAY-CUT. We emphasize that ℓ_p -NORM-MULTIWAY-CUT could also be viewed as a multiway cut that aims for a stronger notion of fairness than MULTIWAY-CUT but a weaker notion of fairness than MIN-MAX-MULTIWAY-CUT. For $k = 2$ terminals,

¹ Throughout this work, all NP-hardness results are strong NP-hardness results and all logarithms are to the base 2.

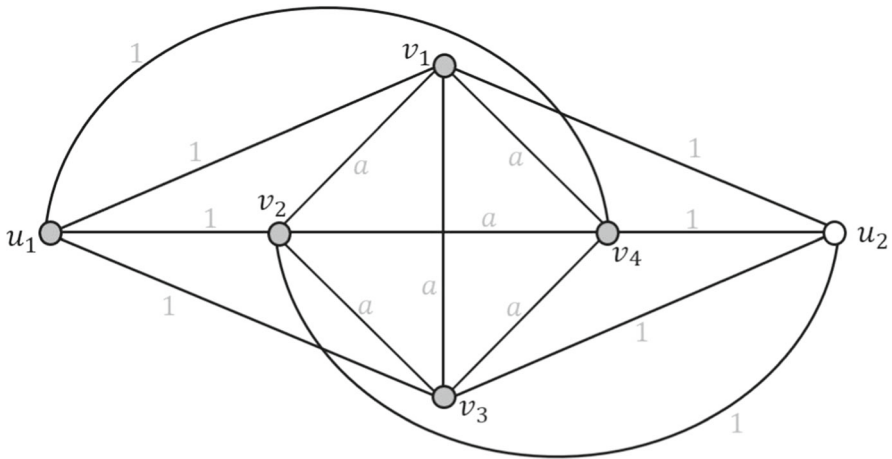


Fig. 1 An example where the unique optimum partition for ℓ_p -NORM-MULTIWAY-CUT for $k = 5$ induces a disconnected part for every $p > 1$. The edge weights are as shown with $a := \lceil 8^{p/(p-1)} \rceil$ and the set of terminals is $\{u_1, v_1, v_2, v_3, v_4\}$. We only need to decide the terminal with which u_2 needs to be grouped. It is cheaper to group u_2 with u_1 than with v_1, v_2, v_3 or v_4 even though u_1 and u_2 are not adjacent: a partition that puts u_2 with one of the terminals in $\{v_1, v_2, v_3, v_4\}$ (and isolates the remaining terminals) has ℓ_p -norm objective value $((3a + 3)^p + 3(3a + 2)^p + 4^p)^{1/p}$ and the partition that puts u_2 with u_1 (and isolates the remaining terminals) has ℓ_p -norm objective value $(4(3a + 2)^p + 8^p)^{1/p}$. To see that the latter is cheaper than the former, we note that the inequality $((3a + 3)^p + 3(3a + 2)^p + 4^p)^{1/p} > (4(3a + 2)^p + 8^p)^{1/p}$ is equivalent to $(3a + 3)^p - (3a + 2)^p > 8^p - 4^p$. The latter is true due to Mean Value Theorem: we have $(3a + 3)^p - (3a + 2)^p \geq p(3a + 2)^{p-1} \geq 8^p > 8^p - 4^p$ for $a = \lceil 8^{p/(p-1)} \rceil$

ℓ_p -NORM-MULTIWAY-CUT reduces to $\min(s, t)$ -cut for all $p \geq 1$ and hence, can be solved in polynomial time.

1.1 Our Results

We begin by remarking that there is a fundamental structural difference between MULTIWAY-CUT and ℓ_p -NORM-MULTIWAY-CUT for $p > 1$ (i.e., between $p = 1$ and $p > 1$). The optimal partition to MULTIWAY-CUT satisfies a nice structural property: assuming that the input graph is connected, every part in an optimal partition for MULTIWAY-CUT will induce a connected subgraph. Consequently, MULTIWAY-CUT is also phrased as the problem of deleting a least weight subset of edges so that the resulting graph contains k connected components with exactly one terminal in each component. However, this nice structural property does not hold for ℓ_p -NORM-MULTIWAY-CUT for $p > 1$ as illustrated by the example in Fig. 1. We remark that Svitkina and Tardos made a similar observation suggesting that the nice structural property fails for MIN-MAX-MULTIWAY-CUT, i.e., for $p = \infty$ —in contrast, our example in Fig. 1 shows that the nice structural property fails to hold for every $p > 1$.

We now discuss our hardness results for ℓ_p -NORM-MULTIWAY-CUT. All NP-hardness results that we mention in this work refer to strong NP-hardness.

Theorem 1.1 *We have the following hardness results for ℓ_p -NORM-MULTIWAY-CUT.*

1. ℓ_p -NORM-MULTIWAY-CUT is NP-hard for every $p > 1$ and every $k \geq 4$.
2. ℓ_p -NORM-MULTIWAY-CUT in planar graphs is NP-hard for every $p > 1$.

We note that the case of $p = 1$ and $p = \infty$ are already known to be hard: MULTIWAY-CUT is NP-hard for $k = 3$ terminals and is NP-hard in planar graphs when k is arbitrary (i.e., when k is not a fixed constant) [13]; MIN-MAX-MULTIWAY-CUT is NP-hard for $k = 4$ terminals and is NP-hard in trees when k is arbitrary [22]. Our NP-hardness in planar graphs result also requires k to be arbitrary.

Given that the problem is NP-hard, we focus on designing approximation algorithms. We show the following result:

Theorem 1.2 *There exists a polynomial-time $O(\log^{1.5} n \log^{0.5} k)$ -approximation for ℓ_p -NORM-MULTIWAY-CUT for every $p \geq 1$, where n is the number of vertices and k is the number of terminals in the input instance.*

Our algorithm to prove Theorem 1.2 is randomized and succeeds with high probability. We note that our approximation factor is $O(\log^2 n)$ since $k \leq n$. While it might be tempting to design an approximation algorithm by solving a convex programming relaxation for ℓ_p -NORM-MULTIWAY-CUT and rounding it, we rule out this approach: the natural convex programming relaxation has an integrality gap of $\Omega(k^{1-1/p})$ —see Sect. 5. Hence, our approach for the approximation algorithm is not based on a convex program but instead based on combinatorial techniques.

For comparison, we state the currently best known approximation factors for $p = 1$ and $p = \infty$: MULTIWAY-CUT admits a 1.2965-approximation via an LP-based algorithm [21] and MIN-MAX-MULTIWAY-CUT admits an $O(\sqrt{\log n \log k})$ -approximation based on a bicriteria approximation for the small-set expansion problem [3].

As a final result, we show that removing the dependence on the number n of vertices in the approximation factor of ℓ_p -NORM-MULTIWAY-CUT is hard assuming the small set expansion hypothesis [20]—see Sect. 6. In particular, we show that achieving a $(k^{1-1/p-\epsilon})$ -approximation for any constant $\epsilon > 0$ is hard. We note that there is a trivial $(k^{p-1} + 1)^{1/p}$ -approximation for ℓ_p -NORM-MULTIWAY-CUT via the well-known isolating cut approach (see Sect. 6.1).

We summarize our results along with known results for MULTIWAY-CUT and MIN-MAX-MULTIWAY-CUT for the sake of comparison in Table 1.

1.2 Outline of Techniques

We briefly outline the techniques underlying our results.

1.2.1 Hardness Results

We show hardness of ℓ_p -NORM-MULTIWAY-CUT for $k = 4$ terminals by a reduction from the graph bisection problem (see Sect. 4.1 for a description of this problem). Our main tool to control the ℓ_p -norm objective in our hardness reduction is the Mean Value Theorem and its consequences (see Propositions 4.1 and 4.2). In order to show NP-hardness of ℓ_p -NORM-MULTIWAY-CUT in planar graphs, we reduce from the 3-partition problem (see Sect. 4.2 for a description of this problem). We do a gadget

Table 1 Summary of results

	ℓ_p -NORM-MULTIWAY-CUT	MIN-MAX-MULTIWAY-CUT	MULTIWAY-CUT
NP-hardness	$k \geq 4$ and on planar graphs (Theorem 1.1)	$k \geq 4$ and on trees [22]	$k \geq 3$ [13]
Approximation factor	$O(\sqrt{\log^3 n \log k})$ (Theorem 1.2)	$O(\sqrt{\log n \log k})$ [3]	1.2965 [21]
Integrality gap lower bound	$\frac{k^{1-1/p}}{2}$ (Sect. 5)	$\frac{k}{2}$ [3]	1.20016 [4, 7]
Inapproximability	No $O(k^{1-1/p-\epsilon})$ -approximation under SSEH (Sect. 6)	No $O(k^{1-\epsilon})$ -approximation under SSEH [3]	No (1.20016 ϵ)-approximation under UGC [4]
Approximation factor of isolating cut approach	$(k^{p-1} + 1)^{\frac{1}{p}}$ (Sect. 6.1)	$2k$	2 [13]

Here, n refers to the number of vertices in the input graph, ϵ is any arbitrary constant, UGC refers to the unique games conjecture, and SSEH refers to the small set expansion hypothesis

based reduction where the gadget is planar. We note that the number of terminals in this reduction is not a constant and is $\Omega(n)$, where n is the number of vertices. Once again, we rely on the Mean Value Theorem and its consequences to control the ℓ_p -norm objective in the reduction. We mention that the starting problems in our hardness reductions are inspired by the hardness results shown by Svitkina and Tardos for MIN-MAX-MULTIWAY-CUT: they showed that MIN-MAX-MULTIWAY-CUT is NP-hard for $k = 4$ terminals by a reduction from the graph bisection problem and that MIN-MAX-MULTIWAY-CUT is NP-hard in trees by a reduction from the 3-partition problem. We also use these same starting problems, but our reductions are more involved owing to the ℓ_p -norm nature of the objective.

1.2.2 Approximation Algorithm

By scaling we may assume that all edge weights are at least one. For the purposes of the algorithm, we will assume knowledge of the optimum value, say OPT —such a value can be guessed within a factor of 2 via binary search. We would run our algorithm for each guess of OPT in the set $\{1, 2, 2^2, 2^3, \dots, 2^{\log(|E| \max_{e \in E} w_e)}\}$. One of the guesses will be within a factor 2 of the true optimum value and our algorithm would achieve the approximation factor for that guess. The overhead in run-time to account for the guess is polynomial in the input size.

Our approximation algorithm proceeds in three steps: (I) *cover*, (II) *uncross*, and (III) *aggregate*. We describe these three steps now.

In the first step of the algorithm—that we term as the *cover* step, we obtain a collection \mathcal{S} of subsets of the vertex set satisfying four properties:

- (1) Each set S in the collection \mathcal{S} has at most one terminal,
- (2) The ℓ_p -norm of the cut values of the sets in the collection raised to the p th power is small, i.e., $\sum_{S \in \mathcal{S}} w(\delta(S))^p = (\beta^p \log n) \text{OPT}^p$ where $\beta = O(\sqrt{\log n \log k})$,
- (3) The number of sets in the collection \mathcal{S} is $O(k \log n)$, and
- (4) The union of the sets in the collection \mathcal{S} is V .

We emphasize that condition (4) is the main reason for calling this step as the *cover* step. We perform this cover step via a multiplicative updates method. For this, we use a bicriteria approximation algorithm for the unbalanced terminal cut problem which was given by Bansal et al. [3] (see Sect. 2 for a description of the unbalanced terminal cut problem and the bicriteria approximation).

Although property (2) above gives a bound on the ℓ_p -norm of the cut values of the sets in the collection \mathcal{S} relative to the optimum, the collection \mathcal{S} does not correspond to a feasible multiway cut: recall that a feasible multiway cut is a partition $\mathcal{P} = (P_1, \dots, P_k)$ of the vertex set where each P_i contains exactly one terminal. The objective of the next two steps is to refine the collection \mathcal{S} to achieve feasibility without blowing up the ℓ_p -norm of the cut values of the parts.

In the second step of the algorithm—that we term as the *uncross* step, we uncross the sets in the collection \mathcal{S} to obtain a partition \mathcal{Q} without increasing the cut values of the sets. We crucially exploit the posimodularity property of the graph cut function to achieve this: posimodularity states that for all subsets $A, B \subseteq V$ of vertices, either

$w(\delta(A)) \geq w(\delta(A - B))$ or $w(\delta(B)) \geq w(\delta(B - A))$. We iteratively consider all pairs of crossing subsets A, B in the collection \mathcal{S} and replace A with $A - B$ if $w(\delta(A)) \geq w(\delta(A - B))$ or replace B with $B - A$ if $w(\delta(B)) \geq w(\delta(B - A))$. The outcome of this step is a partition \mathcal{Q} of the vertex set V satisfying three properties:

- (i) Each part Q in the partition \mathcal{Q} has at most one terminal,
- (ii) The ℓ_p -norm of the cut values of the parts in the partition \mathcal{Q} raised to the p th power is still small, i.e., $\sum_{Q \in \mathcal{Q}} w(\delta(Q))^p = (\beta^p \log n) \text{OPT}^p$, and
- (iii) The number of parts in the partition \mathcal{Q} is $O(k \log n)$.

We encourage the reader to note the similarity between conditions (1), (2), (3), and (4) at the end of the cover step and conditions (i), (ii), and (iii) at the end of the uncross step.

Once again, we observe that the partition \mathcal{Q} at the end of the uncross step may not correspond to a feasible multiway cut: we could have more than k parts in \mathcal{Q} with some of the parts having no terminals. We address this issue in the third step by a careful aggregation.

We term the third step of the algorithm as the *aggregate* step. For the aggregate step, let Q_i be the part in the partition \mathcal{Q} that contains terminal t_i —we have k such parts by property (i)—and let R_1, \dots, R_t be the remaining parts in \mathcal{Q} that contain no terminals. We will aggregate the remaining parts of \mathcal{Q} into the k parts Q_1, \dots, Q_k without blowing up the ℓ_p -norm of the cut value of the parts. By property (iii), the number of remaining parts $t = O(k \log n)$. We create k disjoint buckets B_1, \dots, B_k where B_i contains the union of $O(\log n)$ many parts among R_1, \dots, R_t . Finally, we merge B_i with Q_i . This results in a partition $\mathcal{P} = (Q_1 \cup B_1, \dots, Q_k \cup B_k)$ of V with terminal t_i being in the i th part $Q_i \cup B_i$. The key now is to control the blow-up in the p th power of the ℓ_p -norm of the cut values of the parts in \mathcal{P} : we bound this by a $O(\log^{p-1} n)$ -factor relative to the p th power of the ℓ_p -norm of the cut values of the parts in \mathcal{Q} via Jensen’s inequality; while using Jensen’s inequality, we exploit the fact that each bucket contained $O(\log n)$ many parts. Consequently, using property (ii), the ℓ_p -norm objective value of the cut values of the parts in the partition \mathcal{P} raised to the p th power is still small—we show that $\sum_{P \in \mathcal{P}} w(\delta(P))^p = \beta^p \log^p n \text{OPT}^p$ and hence, we obtain an approximation factor of $O(\beta \log n)$.

The first step of our algorithm is inspired by the $O(\log n)$ -approximation algorithm for MIN-MAX-MULTIWAY-CUT due to Bansal et al. [3]—we modify the multiplicative weights update method and adapt it for ℓ_p -NORM-MULTIWAY-CUT. Our second and third steps differ from that of Bansal et al. We mention that the second and third steps of our algorithm can be adapted to achieve an $O(\beta \log n)$ -approximation factor for ℓ_p -NORM-MULTIWAY-CUT for $p = \infty$, but the resulting approximation factor is only $O(\log^2 n)$ which is weaker than the $O(\log n)$ -factor achieved by Bansal et al. The additional loss of $\log n$ -factor in our algorithm comes from the third step (i.e., the aggregate step). The aggregate step designed in [3] is randomized and saves the $\log n$ -factor in expectation, but it does not generalize to ℓ_p -NORM-MULTIWAY-CUT. As mentioned before, the second step of our algorithm relies on posimodularity. The posimodularity property of the graph cut function has been used in previous works for MIN-MAX-MULTIWAY-CUT in an implicit fashion by a careful and somewhat tedious

edge counting argument [3, 22]. We circumvent the edge counting argument here by the clean posimodularity abstraction. Moreover, the posimodularity abstraction makes the counting considerably easier for our more general problem of ℓ_p -NORM-MULTIWAY-CUT.

1.3 Related Work

ℓ_p -NORM-MULTIWAY-CUT can be viewed as a fairness inducing objective in the context of multiway partitioning problems. Recent works have proposed and studied various fairness inducing objectives for graph cuts and partitioning that are different from ℓ_p -NORM-MULTIWAY-CUT. We briefly discuss these works here. All of the works mentioned in this subsection consider a more general problem known as *correlation clustering*—we discuss these works by specializing to cut and partitioning problems since these specializations are the ones related to our work.

Puleo and Milenkovic [19] introduced a local vertex-wise min–max objective for min (s, t) -cut—here, the goal is to partition the vertex set V of the given edge-weighted undirected graph into two parts $(S, V \setminus S)$ each containing exactly one of the terminals in $\{s, t\}$ so as to minimize $\max_{v \in V} w(\delta(v) \cap \delta(S))$. The motivation behind this objective is that the cut should be fair to every vertex in the graph, i.e., no vertex should pay a lot for the edges in the cut. A result of Chvátal [12] implies that this problem is $(2 - \epsilon)$ -inapproximable for every constant $\epsilon > 0$. Charikar et al. [9] gave an $O(\sqrt{n})$ -approximation for this problem. Reducing the approximability vs inapproximability gap for this problem remains an intriguing open problem. Kalhan et al. [15] considered an ℓ_p -norm version of the objective where the goal is to minimize $(\sum_{v \in V} w(\delta(v) \cap \delta(S))^p)^{1/p}$ and gave an $O(n^{\frac{1}{2} - \frac{1}{2p}} \log^{\frac{1}{2} - \frac{1}{2p}} n)$ -approximation, thus interpolating the best upper bound results for $p = 1$ and $p = \infty$.

Ahmadi et al. [1] introduced a min–max version of multicut: the input consists of an undirected graph $G = (V, E)$ with edge weights $w : E \rightarrow \mathbb{Q}_+$ along with source-sink terminal pairs $(s_1, t_1), \dots, (s_k, t_k)$. The goal is to find a partition $\mathcal{P} = (P_1, \dots, P_r)$ of the vertex set with all source-sink pairs separated by the partition so as to minimize $\max_{i \in [r]} w(\delta(P_i))$. We emphasize that the number of parts here—namely, r —is not constrained by the input and hence, could be arbitrary. Ahmadi, Khuller, and Saha gave an $O(\sqrt{\log n \max\{\log k, \log T\}})$ -approximation for this problem, where T is the number of parts in the optimal solution. Kalhan et al. [15] improved the approximation factor to $2 + \epsilon$.

1.3.1 Organization

We begin with preliminaries in Sect. 2. We present the complete details of our approximation algorithm and prove Theorem 1.2 in Sect. 3. We show the hardness results and prove Theorem 1.1 in Sect. 4. We discuss a convex program and its integrality gap in Sect. 5. We discuss the inapproximability and present a trivial $O(k^{1-1/p})$ -approximation in Sect. 6. We conclude with a few open problems in Sect. 7.

2 Preliminaries

We start with notations that will be used throughout. Let $G = (V, E)$ be an undirected graph with edge weight function $w : E \rightarrow \mathbb{Q}_+$ and vertex weight function $y : V \rightarrow \mathbb{Q}_+$. For every subset $S \subseteq V$, we use $\delta_G(S)$ to denote the set of edges that have exactly one end-vertex in S (we will drop the subscript G when the graph is clear from context), and we write $w(\delta(S)) := \sum_{e \in \delta(S)} w(e)$. Moreover, we will use $y(S)$ to refer to $\sum_{v \in S} y(v)$. We will denote an instance of ℓ_p -NORM-MULTIWAY-CUT by (G, w, T) , where $G = (V, E)$ is the input graph, $w : E \rightarrow \mathbb{Q}_+$ is the edge weight function, and $T \subseteq V$ is the set of terminal vertices. We will call a partition $\tilde{P} = (P_1, \dots, P_r)$ of the vertex set to be a multiway cut if $r = k$ and $t_i \in P_i$ for each $i \in [k]$ and denote the ℓ_p -norm of the cut values of the parts (i.e., $(\sum_{i=1}^k w(\delta(P_i))^p)^{1/p}$) as the ℓ_p -norm objective value of the multiway cut \tilde{P} .

We note that the function $\mu : \mathbb{R} \rightarrow \mathbb{R}$ defined by $\mu(x) := x^p$ is convex for every $p \geq 1$. We will use Jensen’s inequality as stated below in our approximation algorithm as well as our hardness reductions.

Lemma 2.1 (Jensen) *Let $\mu : \mathbb{R} \rightarrow \mathbb{R}$ be a convex function. For arbitrary $x_1, \dots, x_t \in \mathbb{R}$, we have*

$$\mu\left(\frac{1}{t} \sum_{i=1}^t x_i\right) \leq \frac{1}{t} \sum_{i=1}^t \mu(x_i).$$

Our algorithm relies on the graph cut function being symmetric and submodular. We recall that the graph cut function $f : 2^V \rightarrow \mathbb{R}_+$ is given by $f(S) := w(\delta(S))$ for all $S \subseteq V$. Let $f : 2^V \rightarrow \mathbb{R}_+$ be a set function. The function f is symmetric if $f(S) = f(V \setminus S)$ for all $S \subseteq V$, submodular if $f(A) + f(B) \geq f(A \cap B) + f(A \cup B)$ for all $A, B \subseteq V$, and posimodular if $f(A) + f(B) \geq f(A - B) + f(B - A)$ for all $A, B \subseteq V$. Symmetric submodular functions are also posimodular (see Proposition 2.1)—this fact has been used implicitly [3, 22] and explicitly [8, 10] before.

Proposition 2.1 *Symmetric submodular functions are posimodular.*

Proof Let $f : 2^V \rightarrow \mathbb{R}$ be a symmetric submodular set function on a set V , and let $A, B \subseteq V$ be two arbitrary subsets. Then, we have

$$\begin{aligned} f(A) + f(B) &= f(V - A) + f(B) \geq f((V - A) \cup B) + f((V - A) \cap B) \\ &= f(V - (A - B)) + f(B - A) = f(A - B) + f(B - A). \end{aligned}$$

In the above, the first and last equations follow by symmetry and the inequality follows by submodularity. □

Our algorithm for ℓ_p -NORM-MULTIWAY-CUT relies on an intermediate problem, namely the Unbalanced Terminal Cut problem that we introduce now. In Unbalanced Terminal Cut (UTC), the input (G, w, y, τ, T) consists of an undirected graph $G =$

(V, E) , an edge weight function $w : E \rightarrow \mathbb{Q}_+$, a vertex weight function $y : V \rightarrow \mathbb{Q}_+$, a real value $\tau \in [0, 1]$, and a set $T \subseteq V$ of terminal vertices. The goal is to compute

$$\text{UTC}(G, w, y, \tau, T) := \min \{w(\delta(S)) : S \subseteq V, y(S) \geq \tau \cdot y(V), |S \cap T| \leq 1\}.$$

In words, the goal is to find a vertex subset $S \subseteq V$ with minimum cut value such that the total vertex weight of S is at least a τ -fraction of the total vertex weight of V and S contains at most one terminal vertex. Bansal et al. gave a bicriteria approximation for UTC that is summarized in the theorem below.

Theorem 2.1 [3] *There exists a randomized algorithm UTC-BICRIT-ALGO that takes as input (G, w, y, τ, T) consisting of an undirected graph $G = (V, E)$, an edge weight function $w : E \rightarrow \mathbb{Q}_+$, a vertex weight function $y : V \rightarrow \mathbb{Q}_+$, a number $\tau \in [0, 1]$, and a set $T \subseteq V$ of terminal vertices and runs in polynomial time to return a set $S \subseteq V$ satisfying the following three conditions with high probability:*

1. $|S \cap T| \leq 1$,
2. $y(S) = \Omega(\tau)y(V)$, and
3. $w(\delta(S)) \leq \alpha \text{UTC}(G, w, y, \tau, T)$, where $\alpha \leq C\sqrt{\log n \log(1/\tau)}$ for some constant C and $n = |V|$.

3 Approximation Algorithm

In this section, we prove Theorem 1.2. We restate the theorem below for the sake of completeness.

Theorem 1.2 *There exists a polynomial-time $O(\log^{1.5} n \log^{0.5} k)$ -approximation for ℓ_p -NORM-MULTIWAY-CUT for every $p \geq 1$, where n is the number of vertices and k is the number of terminals in the input instance.*

Let OPT be the optimal ℓ_p -norm objective value of a multiway cut in the given instance. For the purposes of the algorithm, we will assume knowledge of a value D such that $D \geq \text{OPT}^p$ —such a value can be guessed via binary search.

We prove Theorem 1.2 via a three step algorithm. In the first step of the algorithm termed as the *cover* step, we will obtain a collection \mathcal{S} of $O(k \log n)$ sets whose union is the vertex set V such that each set in the collection has at most one terminal, the cut value of each set is not too large relative to D , and the ℓ_p -norm of the cut values of the sets in the collection is within a polylog(n) factor of D (see Lemma 3.1). Although the collection \mathcal{S} has low ℓ_p -norm value relative to D , the collection \mathcal{S} may not be a feasible multiway cut. In the second step of the algorithm termed as the *uncross* step, we uncross the sets in the collection \mathcal{S} without increasing the ℓ_p -norm of the cut values of the sets in the collection (see Lemma 3.2). After uncrossing, we obtain a partition, but we could have more than k sets. We address this in our third step termed as the *aggregate* step, where we aggregate parts to ensure that we obtain exactly k parts (see Lemma 3.3). We rely on Jensen’s inequality to ensure that the aggregation does not blow-up the ℓ_p -norm of the cut values of the sets in the partition.

3.1 Cover Step

We begin with the cover step of the algorithm in Lemma 3.1.

Lemma 3.1 *There exists a randomized algorithm that takes as input an undirected graph $G = (V, E)$, an edge weight function $w : E \rightarrow \mathbb{Q}_+$, k distinct terminal vertices $T := \{t_1, \dots, t_k\} \subseteq V$ and a value $D > 0$ such that there exists a partition (P_1^*, \dots, P_k^*) of V with $t_i \in P_i^*$ for all $i \in [k]$ and $\sum_{i=1}^k w(\delta(P_i^*))^p \leq D$, and runs in polynomial time to return a collection of sets $\mathcal{S} \subseteq 2^V$ satisfying the following three conditions with high probability:*

1. $|S \cap T| \leq 1$ and $w(\delta(S)) \leq \beta(2D)^{1/p}$ for every $S \in \mathcal{S}$,
2. $\sum_{S \in \mathcal{S}} w(\delta(S))^p = O(\beta^p \log n)D$, and
3. $|\mathcal{S}| = O(k \log n)$ and $|\{S \in \mathcal{S} : v \in S\}| \geq \log n$ for each $v \in V$,

where $\beta = C\sqrt{\log n \log(2k)}$ (with C being the constant that appears in Theorem 2.1).

Proof We will use Algorithm 1 to obtain the desired collection \mathcal{S} . The correctness of Algorithm 1 is based on Claims 3.1, 3.2 and 3.3.

Algorithm 1 Cover via multiplicative weights update

```

Initialize  $t \leftarrow 1, \mathcal{S} \leftarrow \emptyset, y^1(v) = 1$  for each  $v \in V, Y^1 = \sum_{v \in V} y^1(v)$  and  $\beta = O(\sqrt{\log n \log k})$ 
while  $Y^t > \frac{1}{n}$  do
  for  $i = 1, 2, \dots, \log(2k)$  do
    Execute UTC-BICRIT-ALGO( $G, w, y^t, 2^{-i}, T$ ) to obtain a subset  $S^t(i) \subseteq V$ 
    if  $w(\delta(S^t(i))) \leq \beta(\frac{4D}{2^i})^{1/p}$  then
      Set  $S^t = S^t(i)$  and BREAK
    end if
  end for
   $\mathcal{S} \leftarrow \mathcal{S} \cup \{S^t\}$ .
  for  $v \in V$  do
    Set  $y^{t+1}(v) = \begin{cases} y^t(v)/2 & \text{if } v \in S^t, \\ y^t(v) & \text{if } v \in V \setminus S^t. \end{cases}$ 
  end for
  Set  $Y^{t+1} = \sum_{v \in V} y^{t+1}(v)$ .
  Set  $t \leftarrow t + 1$ .
end while
Return  $\mathcal{S}$ 

```

Our first claim will help in showing that the set S^t added in each iteration of the while loop satisfies certain nice properties.

Claim 3.1 *For every iteration t of the while loop of Algorithm 1, there exists $i \in \{1, 2, \dots, \log(2k)\}$ such that the set $S^t(i)$ satisfies the following conditions:*

1. $|S^t(i) \cap T| \leq 1$,
2. $y^t(S^t(i)) = \Omega(\frac{Y^t}{2^i})$, and
3. $w(\delta(S^t(i))) \leq \beta(\frac{4D}{2^i})^{1/p}$.

Proof We have that $\sum_{i=1}^k y^t(P_i^*) = y^t(V)$ and

$$\sum_{i=1}^k w(\delta(P_i^*))^p \leq D.$$

Let L be the subset of indices of parts for which the cut value is relatively low:

$$L := \left\{ j \in [k] : w(\delta(P_j^*))^p \leq \frac{2y^t(P_j^*)}{Y^t} \cdot D \right\}.$$

It follows that

$$\sum_{j \in [k] \setminus L} y^t(P_j^*) < \sum_{j \in [k] \setminus L} \frac{w(\delta(P_j^*))^p Y^t}{2D} \leq \frac{Y^t}{2}$$

and hence,

$$\sum_{j \in L} y^t(P_j^*) = Y^t - \sum_{j \in [k] \setminus L} y^t(P_j^*) > Y^t - \frac{Y^t}{2} = \frac{Y^t}{2}.$$

Since $|L| \leq k$, there exists an index $q \in L$ such that $y^t(P_q^*) > Y^t/(2k)$. Let us fix i_0 to be an integer such that $y^t(P_q^*) \in (Y^t \cdot 2^{-i_0}, Y^t \cdot 2^{-i_0+1}]$. Then, we must have $i_0 \leq \log(2k)$. We note that the set P_q^* satisfies $|P_q^* \cap T| = 1$ and $y^t(P_q^*) > Y^t/(2k) = y^t(V)/(2k)$. This implies P_q^* is feasible to the UTC problem on input $(G, w, y^t, 1/2^{i_0}, T)$. Therefore, according to Theorem 2.1, the set $S^t(i_0)$ has the following properties: Firstly, $|S^t(i_0) \cap T| \leq 1$. Secondly, $y^t(S^t(i_0)) = \Omega(1/2^{i_0})y^t(V) = \Omega(Y^t/2^{i_0})$. Finally,

$$w(\delta(S^t(i_0))) \leq C \sqrt{\log n \log(2k)} \cdot \text{UTC} \left(G, w, y^t, \frac{1}{2^{i_0}}, T \right) \leq \beta \cdot w(\delta(P_q^*)) \tag{1a}$$

$$\leq \beta \cdot \left(\frac{2y^t(P_q^*)}{Y^t} \cdot D \right)^{\frac{1}{p}} \tag{1b}$$

$$\leq \beta \cdot \left(\frac{2 \cdot Y^t \cdot 2^{-i_0+1}}{Y^t} \right) = \beta \cdot \left(\frac{4D}{2^{i_0}} \right)^{\frac{1}{p}}. \tag{1c}$$

Inequality (1a) follows from the fact that P_q^* is feasible to $\text{UTC}(G, w, y^t, 1/2^{i_0}, T)$, Inequality (1b) follows from the fact that $q \in L$, and Inequality (1c) follows from the choice of i_0 . This completes the proof of Claim 3.1. \square

For the rest of the proof, we will use the following notation: In the t 'th iteration of the while loop of Algorithm 1, we will fix $i_t \in \{1, 2, \dots, \log(2k)\}$ to be the integer such that $S^t = S^t(i_t)$. We will use ℓ to denote the total number of iterations of the while loop. For each $v \in V$, we define $N_v := |\{t \in [\ell] : v \in S^t\}|$ to be the number of sets in the collection \mathcal{S} that contain the vertex v .

We observe that for each $v \in V$, we have $y^{\ell+1}(v) = 2^{-N_v}$. Claim 3.1 and Theorem 2.1 together imply that the t 'th iteration of the while loop leads to a set S^t being added to the collection \mathcal{S} such that the following three properties hold:

1. $|S^t \cap T| \leq 1$,
2. $y^t(S^t) = \Omega\left(\frac{Y^t}{2^{i_t}}\right)$, and
3. $w(\delta(S^t)) \leq \beta\left(\frac{4D}{2^{i_t}}\right)^{1/p}$.

Our next claim shows that the number of iterations of the while loop executed in Algorithm 1 is small. Moreover, the union of the sets in the collection \mathcal{S} is the vertex set V .

Claim 3.2 *The number of iterations ℓ of the while loop satisfies $\ell = O(k \log n)$. Moreover, $N_v \geq \log n$ for each $v \in V$.*

Proof Upon termination of Algorithm 1, we must have $Y^{\ell+1} \leq 1/n$. Combining with the earlier observation that $y^{\ell+1}(v) = 2^{-N_v}$ for every $v \in V$, we have that

$$2^{-N_v} = y^{\ell+1}(v) \leq Y^{\ell+1} \leq \frac{1}{n},$$

which implies that $N_v \geq \log n$ for every $v \in V$.

It remains to show that $\ell = O(k \log n)$. Consider the t th iteration of the while loop for an arbitrary $t \in [\ell]$. By property 2 of the set S^t stated above, we have that $y^t(S^t) \geq cY^t/2^{i_t} \geq cY^t/(2k)$ for some constant $c > 0$. Consequently,

$$Y^{t+1} = Y^t - \frac{y^t(S^t)}{2} \leq Y^t - \frac{cY^t}{4k} = \left(1 - \frac{c}{4k}\right) Y^t.$$

Due to the termination condition of the while loop, we know that $Y^\ell > 1/n$. Hence,

$$\frac{1}{n} < Y^\ell \leq \left(1 - \frac{c}{4k}\right)^{\ell-1} Y^1 = \left(1 - \frac{c}{4k}\right)^{\ell-1} n \leq \exp\left(-\frac{c(\ell-1)}{4k}\right) n.$$

Therefore, $\frac{c(\ell-1)}{4k} = O(\log n)$ which implies that $\ell = O(k \log n)$. This completes the proof of Claim 3.2. □

The next claim bounds the ℓ_p -norm of the cut values of the sets in the collection \mathcal{S} .

Claim 3.3 *The collection \mathcal{S} returned by Algorithm 1 satisfies $\sum_{S \in \mathcal{S}} w(\delta(S))^p = O(\beta^p \log n) \cdot D$.*

Proof Consider the t th iteration of the while loop for an arbitrary $t \in [\ell]$. By property 3 of the set S^t stated above, we have that $w(\delta(S^t)) \leq \beta(4D/2^{i_t})^{1/p}$ and consequently, $2^{i_t} \leq 4D\beta^p \cdot w(\delta(S^t))^{-p}$. Moreover, by property 2 of the set S^t stated above, we have that $y^t(S^t) \geq cY^t/2^{i_t}$ for some constant $c > 0$. Hence,

$$y^t(S^t) \geq \frac{cY^t}{2^{i_t}} \geq \frac{cY^t \cdot w(\delta(S^t))^p}{\beta^p \cdot 4D}.$$

Therefore,

$$Y^{t+1} = Y^t - \frac{y^t(S^t)}{2} \leq \left(1 - \frac{c \cdot w(\delta(S^t))^p}{\beta^p \cdot 8D}\right) Y^t.$$

Using the fact that $Y^\ell > 1/n$, we observe that

$$\begin{aligned} \frac{1}{n} < Y^\ell &\leq Y^1 \cdot \prod_{t=1}^{\ell-1} \left(1 - \frac{c \cdot w(\delta(S^t))^p}{\beta^p \cdot 8D}\right) = n \cdot \prod_{t=1}^{\ell-1} \left(1 - \frac{c \cdot w(\delta(S^t))^p}{\beta^p \cdot 8D}\right) \\ &\leq n \cdot \prod_{i=1}^{\ell-1} \exp\left(-\frac{c \cdot w(\delta(S^t))^p}{\beta^p \cdot 8D}\right) = n \cdot \exp\left(-\frac{c \cdot \sum_{i=1}^{\ell-1} w(\delta(S^t))^p}{\beta^p \cdot 8D}\right). \end{aligned}$$

This implies that $\frac{c \cdot \sum_{i=1}^{\ell-1} w(\delta(S^t))^p}{\beta^p \cdot 8D} = O(\log n)$, and hence $\sum_{i=1}^{\ell-1} w(\delta(S^t))^p = O(\beta^p \log n) \cdot D$.

In the ℓ 'th iteration of the while loop, we have $w(\delta(S^\ell)) \leq \beta(4D/2^{i_\ell})^{1/p}$ by property 3 of the set S^t stated above and hence $w(\delta(S^\ell))^p \leq \beta^p \cdot 4D/2^{i_\ell} \leq O(\beta^p D)$. Consequently, $\sum_{i=1}^\ell w(\delta(S^t))^p = O(\beta^p \log n) \cdot D$. This completes the proof of Claim 3.3. \square

We now show correctness of our algorithm to complete the proof of Lemma 3.1. Firstly, we note that every $S \in \mathcal{S}$ satisfies $|S \cap T| \leq 1$ by property 1 of the set S^t stated above. Moreover, there exists $i \in \{1, 2, \dots, \log(2k)\}$ such that $w(\delta(S)) \leq \beta(4D/2^i)^{1/p} \leq \beta(2D)^{1/p}$, which implies Conclusion 1 of Lemma 3.1. Secondly, Conclusion 2 of Lemma 3.1 is implied by Claim 3.3. Finally, Conclusion 3 of Lemma 3.1 is implied by Claim 3.2 because each iteration of the while loop adds exactly one new set to the collection \mathcal{S} .

We now bound the run time of Algorithm 1. Each iteration of the while loop takes polynomial time due to Theorem 2.1, and the number of iterations of the while loop is $O(k \log n)$. This implies that the total run time of Algorithm 1 is indeed polynomial in the size of the input. \square

We note that the our algorithm to prove Lemma 3.1 is randomized only because of the use of the randomized algorithm given in Theorem 2.1.

3.2 Uncross Step

The collection \mathcal{S} that we obtain in Lemma 3.1 may not be a partition. Our next lemma will uncross the collection \mathcal{S} obtained from Lemma 3.1 to obtain a partition without increasing the cut values of the sets.

Lemma 3.2 *There exists an algorithm that takes as input a collection $\mathcal{S} \subseteq 2^V$ of subsets of vertices satisfying the conclusions in Lemma 3.1 and runs in polynomial time to return a partition $\tilde{\mathcal{Q}}$ of V such that*

1. $|Q \cap T| \leq 1$ for each $Q \in \tilde{\mathcal{Q}}$,
2. $\sum_{Q \in \tilde{\mathcal{Q}}} w(\delta(Q))^p \leq \sum_{S \in \mathcal{S}} w(\delta(S))^p$, and
3. the number of parts in $\tilde{\mathcal{Q}}$ is $O(k \log n)$.

Proof For convenience, we will define $f : 2^V \rightarrow \mathbb{R}_+$ by $f(S) := w(\delta(S))$ for all $S \subseteq V$. We will use Algorithm 2 to obtain the desired partition $\tilde{\mathcal{Q}}$ of V .

Algorithm 2 Uncross via posimodularity

```

Initialize  $\tilde{\mathcal{Q}} \leftarrow \mathcal{S}$ 
while there exist distinct sets  $A, B \in \tilde{\mathcal{Q}}$  such that  $A \cap B \neq \emptyset$  do
  if  $f(A) \geq f(A - B)$  then
    Set  $A \leftarrow A - B$ 
  else
    Set  $B \leftarrow B - A$ 
  end if
end while
Return  $\tilde{\mathcal{Q}}$ 

```

We now prove the correctness of Algorithm 2. We begin by observing that Algorithm 2 indeed outputs a partition of the vertex set: Firstly, the while loop enforces that the output $\tilde{\mathcal{Q}}$ satisfies $A \cap B = \emptyset$ for all distinct $A, B \in \tilde{\mathcal{Q}}$. Secondly, during each iteration of the while loop, the set $\bigcup_{Q \in \tilde{\mathcal{Q}}} Q$ remains unchanged: In the iteration of the while loop that uncrosses $A, B \in \tilde{\mathcal{Q}}$, let A' and B' denote the updated sets at the end of the while loop, respectively. Then we must have $A' \cup B' = (A - B) \cup B = A \cup B$ or $A' \cup B' = A \cup (B - A) = A \cup B$. In either case, since $A' \cup B' = A \cup B$, the set $\bigcup_{Q \in \tilde{\mathcal{Q}}} Q$ remains unchanged after the update. Therefore, we have $\bigcup_{Q \in \tilde{\mathcal{Q}}} Q = \bigcup_{S \in \mathcal{S}} S$. We recall that $\bigcup_{S \in \mathcal{S}} S = V$ by conclusion 3 of Lemma 3.1. Hence, $\tilde{\mathcal{Q}}$ is indeed a partition of V .

Furthermore, each set Q in the output $\tilde{\mathcal{Q}}$ is a subset of some set $S \in \mathcal{S}$. This implies $|Q \cap T| \leq |S \cap T| \leq 1$, thus proving the first conclusion.

To prove the second conclusion, we use posimodularity of f as shown in Proposition 2.1. Namely, for every $A, B \subseteq V$,

$$f(A) + f(B) \geq f(A - B) + f(B - A).$$

Therefore, at least one of the following two hold: either $f(A) \geq f(A - B)$ or $f(B) \geq f(B - A)$. This implies that, by the choice of the algorithm, $\sum_{Q \in \tilde{Q}} f(Q)^p$ does not increase.

To see the third conclusion, we note that after each iteration of the while loop, the size of \tilde{Q} is unchanged. Therefore, at the end Algorithm 2, we have $|\tilde{Q}| = |\mathcal{S}| = O(k \log n)$ by Lemma 3.1.

Finally, we bound the run time as follows. At initialization, there are $O((k \log n)^2)$ pairs $(A, B) \in \tilde{Q}^2$ such that $A \cap B \neq \emptyset$. After each iteration of the while loop, the number of such pairs decreases by at least 1. Therefore, the total number of iterations of the while loop is $O((k \log n)^2)$. Hence, Algorithm 2 indeed runs in polynomial time. \square

3.3 Aggregation Step

The partition \tilde{Q} that we obtain in Lemma 3.2 may contain more than k parts and hence, some of the parts may not contain any terminals. Our next lemma will aggregate the parts in \tilde{Q} from Lemma 3.2 to obtain a k -partition that contains exactly one terminal in each part while controlling the increase in the ℓ_p -norm of the cut value of the parts.

Lemma 3.3 *There exists an algorithm that takes as input a partition \tilde{Q} of V satisfying the conclusions in Lemma 3.2 and runs in polynomial time to return a partition (P_1, P_2, \dots, P_k) of V such that*

1. $t_i \in P_i$ for each $i \in [k]$, and
2. $\sum_{i=1}^k w(\delta(P_i))^p = O((\beta \log n)^p) \cdot D$,

where $\beta = C\sqrt{\log n \log(2k)}$ (with C being the constant that appears in Theorem 2.1).

Proof We will use Algorithm 3 on input \tilde{P} to obtain the desired partition.

Algorithm 3 Aggregate via bucketing

```

Let  $\mathcal{F} = \{Q \in \tilde{Q} : Q \cap T = \emptyset\}$ 
Let  $\mathcal{P}' = \{Q \in \tilde{Q} : Q \cap T \neq \emptyset\} = \{Q'_1, \dots, Q'_k\}$ , where  $t_i \in Q'_i$  for each  $i \in [k]$ 
Partition the sets in  $\mathcal{F}$  into  $k$  buckets  $B_1, \dots, B_k$  such that  $|B_i| = O(\log n)$  for each  $i \in [k]$  (arbitrarily)
for  $i = 1, 2, \dots, k$  do
    Set  $P_i \leftarrow Q'_i \cup \left(\bigcup_{A \in B_i} A\right)$ 
end for
Return  $(P_1, \dots, P_k)$ 

```

The run time of Algorithm 3 is linear in its input size. We now argue the correctness. We note that the third step in Algorithm 3 is possible because $|\mathcal{F}| \leq |\tilde{Q}| = O(k \log n)$.

Since $|Q \cap T| \leq 1$ for each $Q \in \tilde{Q}$, the tuple (P_1, \dots, P_k) returned by Algorithm 3 is indeed a partition of V satisfying $t_i \in P_i$ for all $i \in [k]$. We will now bound $\sum_{i=1}^k f(P_i)^p$, where $f : 2^V \rightarrow \mathbb{R}_+$ is given by $f(S) := w(\delta(S))$ for all $S \subseteq V$. We have that

$$\sum_{i=1}^k f(P_i)^p = \sum_{i=1}^k f\left(Q'_i \cup \left(\bigcup_{A \in B_i} A\right)\right)^p \leq \sum_{i=1}^k \left(f(Q'_i) + \sum_{A \in B_i} f(A)\right)^p.$$

Since the number of sets in B_i is $O(\log n)$, we have the following using Jensen’s inequality (Lemma 2.1) for each $i \in [k]$:

$$\begin{aligned} \left(f(Q'_i) + \sum_{A \in B_i} f(A)\right)^p &\leq (|B_i| + 1)^{p-1} \left(f(Q'_i)^p + \sum_{A \in B_i} f(A)^p\right) \\ &= O(\log^{p-1} n) \left(f(Q'_i)^p + \sum_{A \in B_i} f(A)^p\right). \end{aligned}$$

Hence,

$$\begin{aligned} \sum_{i=1}^k f(P_i)^p &= \sum_{i=1}^k O(\log^{p-1} n) \left(f(Q'_i)^p + \sum_{A \in B_i} f(A)^p\right) \\ &= O(\log^{p-1} n) \sum_{Q \in \tilde{\mathcal{Q}}} f(Q)^p \\ &= O(\log^{p-1} n) \sum_{S \in \mathcal{S}} f(S)^p = \beta^p O(\log^p n) D. \end{aligned}$$

The last but one equality above is due to Conclusion 2 of Lemma 3.2 and the last equality is due to Conclusion 2 of Lemma 3.1. Hence, $\sum_{i=1}^k w(\delta(P_i))^p = \sum_{i=1}^k f(P_i)^p = O((\beta \log n)^p) D$. \square

3.4 Proof of Theorem 1.2

Lemmas 3.1, 3.2, and 3.3 together lead to an algorithm that takes as input an undirected graph $G = (V, E)$, an edge weight function $w : E \rightarrow \mathbb{R}_+$, k distinct terminal vertices $T := \{t_1, \dots, t_k\} \subseteq V$, and a value $D > 0$ with the property that there exists a partition (P_1^*, \dots, P_k^*) of V with $t_i \in P_i^*$ for all $i \in [k]$ and $\sum_{i=1}^k w(\delta(P_i^*))^p \leq D$. This algorithm runs in polynomial time to return a multiway cut $\mathcal{P} = (P_1, \dots, P_k)$ such that

$$\begin{aligned} \left(\sum_{i=1}^k w(\delta(P_i))^p\right)^{\frac{1}{p}} &= (O((\beta \log n)^p) D)^{\frac{1}{p}} = O(\beta \log n) D^{\frac{1}{p}} \\ &= O(\log^{1.5} n \log^{0.5} k) D^{\frac{1}{p}}. \end{aligned}$$

In order to prove Theorem 1.2, we may use binary search to guess $D \in [\text{OPT}^p, (2\text{OPT})^p]$ and run the above algorithm to obtain a multiway cut $\mathcal{P} = (P_1, \dots, P_k)$ such that

$$\left(\sum_{i=1}^k w(\delta(P_i))^p \right)^{\frac{1}{p}} = O(\log^{1.5} n \log^{0.5} k) D^{\frac{1}{p}} = O(\log^{1.5} n \log^{0.5} k) \text{OPT}.$$

This completes the proof of Theorem 1.2.

4 NP-Hardness

In this section, we show NP-hardness results for ℓ_p -NORM-MULTIWAY-CUT thereby proving Theorem 1.1. In Sect. 4.1, we show that ℓ_p -NORM-MULTIWAY-CUT is NP-hard for $k = 4$ terminals for every $p > 1$ by a reduction from graph bisection. In Sect. 4.2, we show that ℓ_p -NORM-MULTIWAY-CUT is NP-hard in planar graphs for every $p > 1$ by a reduction from 3-partition. In our reductions, we will frequently use the following two consequences of the Mean Value Theorem. We recall that the function $\mu(x) = x^p$ is differentiable.

Proposition 4.1 *For a differentiable function $\mu : [0, \infty) \rightarrow \mathbb{R}$, and two real values $0 < x \leq y$, we have*

$$(y - x) \inf_{z \in [x, y]} \mu'(z) \leq \mu(x) - \mu(y) \leq (y - x) \sup_{z \in [x, y]} \mu'(z).$$

Proposition 4.2 *For $p \geq 1$ and real values $0 < x_1 \leq x_2 \leq x_3 \leq x_4$ such that $x_2 + x_3 = x_1 + x_4$, we have $x_2^p + x_3^p \leq x_1^p + x_4^p$.*

Proof We have

$$x_4^p - x_3^p \geq p(x_4 - x_3)x_3^{p-1} = p(x_2 - x_1)x_3^{p-1} \geq p(x_2 - x_1)x_2^{p-1} \geq x_2^p - x_1^p.$$

The first and last inequalities above are by Proposition 4.1. □

4.1 NP-Hardness for Constant Number of Terminals

The following is the main result of this section.

Theorem 4.1 *ℓ_p -NORM-MULTIWAY-CUT is NP-hard for every $p \geq 1$ and $k \geq 4$.*

Proof We note that when $p = 1$, ℓ_p -NORM-MULTIWAY-CUT corresponds to MULTIWAY-CUT and is known to be NP-hard for every $k \geq 3$ [13]. For the rest of our proof, we will fix $p > 1$.

Our hardness reduction is from BISECTION which is known to be NP-complete. BISECTION is defined as follows: Given an undirected graph $G = (V, E)$ where $|V| =: n$ is even and an integer C , the goal is to decide if there exists a subset $S \subseteq V$ such that $|S| = n/2$ and $|\delta_G(S)| \leq C$.

Given an instance $(G = (V, E), C)$ of BISECTION, we construct an instance (G', w', T) of ℓ_p -NORM-MULTIWAY-CUT consisting of a graph $G' = (V', E')$, an edge weight function $w' : E' \rightarrow \mathbb{R}_+$, and a set $T \subseteq V'$ of 4 terminals vertices as follows:

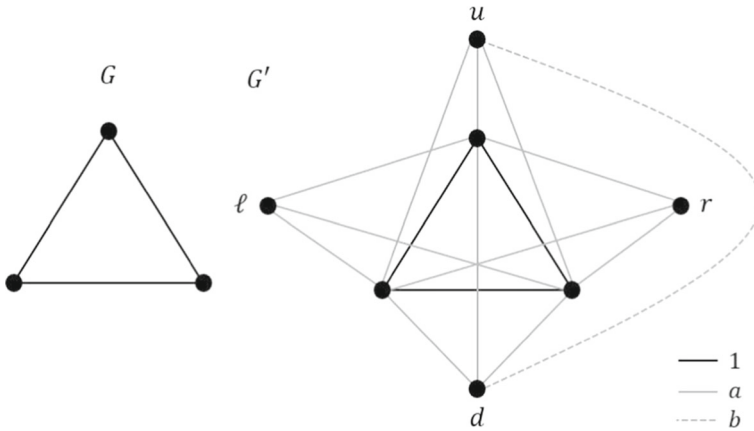


Fig. 2 An example showing the construction of G' from G that is used in the proof of Theorem 4.1. The graph G is a 3-cycle, and the graph G' is shown on the right hand side. The solid black edges are each of weight 1, the solid grey edges are each of weight a , and the dashed grey edge is of weight b

$$\begin{aligned}
 V' &:= V \cup \{u, d, \ell, r\}, \\
 E' &:= E \cup \{ud\} \cup \{vu, vd, vl, vr : v \in V\}, \\
 T &:= \{u, d, \ell, r\}, \\
 w'(e) &:= \begin{cases} 1 & \text{if } e \in E \\ a & \text{if } e \in \{vu, vd, vl, vr : v \in V\}, \\ b & \text{if } e = ud \end{cases}
 \end{aligned}$$

where the parameters a and b are given by

$$a := \max \left\{ 1, \left\lceil \frac{8n^3}{p-1} \right\rceil, 2C + 1 \right\}, \quad b := 1 + \max \left\{ 1, \left\lceil (2an + C)^{\frac{p}{p-1}} \right\rceil, 3an \right\}.$$

We show an example construction of G' from G in Fig. 2.

We note that for every fixed $p > 1$, the size of (G', w', T) is polynomial in the size of (G, C) . The following lemma completes the proof of the theorem. \square

Lemma 4.1 *There exists a subset $S \subseteq V$ such that $|S| = n/2$ and $|\delta_G(S)| \leq C$ if and only if (G', w', T) has a multiway cut whose ℓ_p -norm objective value is at most*

$$(2(b + an)^p + 2(2an + C)^p)^{\frac{1}{p}}.$$

Proof We start by showing the forward direction.

Claim 4.1 *If there exists a subset $S \subseteq V$ such that $|S| = n/2$ and $|\delta_G(S)| \leq C$, then (G', w', T) has a multiway cut whose ℓ_p -norm objective value is at most $(2(b + an)^p + 2(2an + C)^p)^{1/p}$.*

Proof Let $S \subseteq V$ satisfy $|S| = n/2$ and $|\delta_G(S)| \leq C$. Then the ℓ_p -norm objective value of the multiway cut $(\{u\}, \{d\}, S \cup \{\ell\}, (V \setminus S) \cup \{r\})$ raised to the p th power is at most

$$(b + an)^p + (b + an)^p + (3a|S| + a|V \setminus S| + C)^p + (3a|V \setminus S| + a|S| + C)^p = 2(b + an)^p + 2(2an + C)^p.$$

□

In order to show the reverse direction, we need the following structural result on multiway cuts of (G', w', T) with cheap ℓ_p -norm objective value.

Claim 4.2 *If G' has a multiway cut \mathcal{P} whose ℓ_p -norm objective value is at most $(2(b + an)^p + 2(2an + C)^p)^{1/p}$, then the parts of \mathcal{P} containing u and d are singletons.*

Proof Let $\mathcal{P} = (U \cup \{u\}, D \cup \{d\}, L \cup \{\ell\}, R \cup \{r\})$ be a multiway cut of G' whose ℓ_p -norm objective value raised to p th power is at most $2(b + an)^p + 2(2an + C)^p$, where $U \cup D \cup L \cup R = V$. Without loss of generality, suppose that U is non-empty. Then we have

$$w'(\delta(U \cup \{u\})) \geq b + 3a|U| + a(n - |U|) = b + 2a|U| + an \geq b + 2a + an, \\ w'(\delta(D \cup \{d\})) \geq b + 3a|D| + a(n - |D|) = b + 2a|D| + an \geq b + an.$$

This implies that the ℓ_p -norm objective value of \mathcal{P} raised to the p th power is at least $(b + 2a + an)^p + (b + an)^p$. By assumption, the ℓ_p -norm objective value of \mathcal{P} raised to the p th power is at most $2(b + an)^p + 2(2an + C)^p$. Thus,

$$0 \geq (b + 2a + an)^p + (b + an)^p - (2(b + an)^p + 2(2an + C)^p) = (b + 2a + an)^p - (b + an)^p - 2(2an + C)^p$$

Setting $\mu(z) = z^p$, $x = b + an$ and $y = b + 2a + an$ in Proposition 4.1, we observe that

$$(b + 2a + an)^p - (b + an)^p \geq 2a \cdot \inf_{z \in [b+an, b+2a+an]} pz^{p-1} = 2ap(b + an)^{p-1}.$$

Therefore,

$$0 \geq (b + 2a + an)^p - (b + an)^p - 2(2an + C)^p \geq 2ap(b + an)^{p-1} - 2(2an + C)^p > 2ap \left((2an + C)^{\frac{p}{p-1}} + an \right)^{p-1} - 2(2an + C)^p \geq 2ap \left((2an + C)^{\frac{p}{p-1}} \right)^{p-1} - 2(2an + C)^p = 2ap(2an + C)^p - 2(2an + C)^p \geq 0.$$

Here the strict inequality follows from our choice of $b > (2an + C)^{p/(p-1)}$ and the last inequality follows from $a \geq 1$ and $p > 1$. This is a contradiction since one of the inequalities in the above sequence is strict. Hence, we must have $U = D = \emptyset$. \square

The following claim proves the reverse direction of the lemma by showing that a multiway cut of (G', w', T) that is cheap in ℓ_p -norm objective value can be used to recover a cheap bisection.

Claim 4.3 *If a multiway cut $\mathcal{P} = (U \cup \{u\}, D \cup \{d\}, L \cup \{\ell\}, R \cup \{r\})$ of (G', w', T) has ℓ_p -norm objective value at most $(2(b + an)^p + 2(2an + C)^p)^{1/p}$, then $L \cup R = V$, $|L| = |R| = n/2$, and $|\delta_G(L)| \leq C$.*

Proof By Claim 4.2, we know that $U = D = \emptyset$, and hence $L \cup R = V$. We note that in this case, we have

$$\begin{aligned} w'(\delta(U \cup \{u\})) &= w'(\delta(D \cup \{d\})) = b + an, \\ w'(\delta(L \cup \{\ell\})) &\geq 3a|L| + a(n - |L|) = 2a|L| + an, \\ w'(\delta(R \cup \{r\})) &\geq 3a|R| + a(n - |R|) = 2a|R| + an. \end{aligned}$$

This implies that the ℓ_p -norm objective value of \mathcal{P} raised to the p th power is at least $2(b + an)^p + (2a|L| + an)^p + (2a|R| + an)^p$. By assumption, the ℓ_p -norm objective value of \mathcal{P} raised to the p th power at most $2(b + an)^p + 2(2an + C)^p$. Hence,

$$\begin{aligned} 0 &\geq 2(b + an)^p + (2a|L| + an)^p + (2a|R| + an)^p - (2(b + an)^p + 2(2an + C)^p) \\ &= (2a|L| + an)^p + (2a|R| + an)^p - 2(2an + C)^p. \end{aligned} \tag{2}$$

For the sake of contradiction, suppose that $|L| \neq n/2$. Without loss of generality, let $|L| \geq n/2 + 1$ and $|R| \leq n/2 - 1$. We note that $|L| + |R| = n = (n/2 - 1) + (n/2 + 1)$.

In Proposition 4.2, by setting

$$x_1 = 2a|R| + an, \quad x_2 = 2a\left(\frac{n}{2} - 1\right) + an, \quad x_3 = 2a\left(\frac{n}{2} + 1\right) + an, \quad x_4 = 2a|L| + an,$$

we get

$$(2a|L| + an)^p + (2a|R| + an)^p \geq \left(2a\left(\frac{n}{2} + 1\right) + an\right)^p + \left(2a\left(\frac{n}{2} - 1\right) + an\right)^p.$$

Substituting this in inequality (2), we get that

$$\begin{aligned} 0 &\geq (2a|L| + an)^p + (2a|R| + an)^p - 2(2an + C)^p \\ &\geq \left(2a\left(\frac{n}{2} + 1\right) + an\right)^p + \left(2a\left(\frac{n}{2} - 1\right) + an\right)^p - 2(2an + C)^p \\ &= (2an + 2a)^p + (2an - 2a)^p - 2(2an + C)^p \\ &= ((2an + 2a)^p - (2an + a)^p) + ((2an + a)^p - (2an + C)^p) \\ &\quad - ((2an + C)^p - (2an)^p) - ((2an)^p - (2an - 2a)^p) \end{aligned} \tag{3}$$

By applying Proposition 4.1 four times, we get that

$$\begin{aligned} (2an + 2a)^p - (2an + a)^p &\geq ap(2an + a)^{p-1}, \\ (2an + a)^p - (2an + C)^p &\geq (a - C)p(2an + C)^{p-1}, \\ (2an + C)^p - (2an)^p &\leq Cp(2an + C)^{p-1}, \\ (2an)^p - (2an - 2a)^p &\leq 2ap(2an)^{p-1}. \end{aligned}$$

Substituting these in inequality (3), we get that

$$\begin{aligned} 0 &\geq ap(2an + a)^{p-1} + (a - C)p(2an + C)^{p-1} - Cp(2an + C)^{p-1} - 2ap(2an)^{p-1} \\ &= ap(2an + a)^{p-1} + (a - 2C)p(2an + C)^{p-1} - 2ap(2an)^{p-1} \\ &\geq ap(2an + a)^{p-1} + (a - 2C)p(2an)^{p-1} - 2ap(2an)^{p-1}. \end{aligned}$$

Let $\epsilon := (p - 1)/(8n)$. Since $a \geq 8n^3/(p - 1)$, we have that $2C < n^2 \leq \epsilon a$. This implies

$$\begin{aligned} 0 &\geq ap(2an + a)^{p-1} + (a - 2C)p(2an)^{p-1} - 2ap(2an)^{p-1} \\ &> ap(2an + a)^{p-1} + (1 - \epsilon)ap(2an)^{p-1} - 2ap(2an)^{p-1}. \end{aligned}$$

This inequality is equivalent to

$$0 > (2n + 1)^{p-1} + (1 - \epsilon)(2n)^{p-1} - 2(2n)^{p-1} = (2n + 1)^{p-1} - (1 + \epsilon)(2n)^{p-1},$$

which further implies

$$\epsilon > \left(1 + \frac{1}{2n}\right)^{p-1} - 1.$$

Applying Proposition 4.1 again, we get

$$\epsilon > \left(1 + \frac{1}{2n}\right)^{p-1} - 1 \geq \frac{1}{2n} \cdot \inf_{z \in [1, 1 + \frac{1}{2n}]} (p - 1)z^{p-2} = \frac{p - 1}{2n} \inf_{z \in [1, 1 + \frac{1}{2n}]} z^{p-2}.$$

If $p \geq 2$, the infimum z^{p-2} for $z \in [1, 1 + 1/(2n)]$ is attained at $z = 1$, and thus $z^{p-2} \geq 1$ for all $z \in [1, 1 + 1/(2n)]$. If $p \in (1, 2)$, the infimum of z^{p-2} for $z \in [1, 1 + 1/(2n)]$ is attained at $z = 1 + 1/(2n)$, and thus $z^{p-2} \geq (1 + 1/(2n))^{p-2} > 2^{p-2} > 1/2$ for all $z \in [1, 1 + 1/(2n)]$. Hence,

$$\epsilon > \frac{p - 1}{2n} \cdot \inf_{z \in [1, 1 + \frac{1}{2n}]} z^{p-2} > \frac{p - 1}{2n} \cdot \frac{1}{2} > \frac{p - 1}{8n} = \epsilon.$$

This leads to a contradiction since one of the inequalities in the above sequence is strict. Hence, we must have $|L| = |R| = n/2$. Finally, we prove the last conclusion

that $|\delta_G(L)| \leq C$: since $|L| = |R| = n/2$, the four parts of \mathcal{P} are $\{u\}$, $\{d\}$, L , and R . Consequently, the ℓ_p -norm objective value of \mathcal{P} raised to p th power is

$$2(b + an)^p + 2(2an + |\delta_G(L)|)^p$$

which is known to be at most $2(b + an)^p + 2(2an + C)^p$. Hence, $|\delta_G(L)| \leq C$ as claimed. \square

4.2 NP-Hardness in Planar Graphs

The following is the main result of this section.

Theorem 4.2 ℓ_p -NORM-MULTIWAY-CUT in planar graphs is NP-hard for every $p \geq 1$.

Proof We note that when $p = 1$, ℓ_p -NORM-MULTIWAY-CUT corresponds to MULTIWAY-CUT and is known to be NP-hard in planar graphs [13]. For the rest of our proof, we will fix $p > 1$.

Our hardness reduction is from 3- PARTITION which is known to be NP-hard [14]. 3- PARTITION is defined as follows: Given a set $S = [3m]$, a sequence of weights a_1, a_2, \dots, a_{3m} , and a bound B satisfying $\sum_{i=1}^{3m} a_i = mB$ and $B/4 < a_i < B/2$ for all $i \in [3m]$, the goal is to decide whether there exists a partition of S into m subsets S_1, S_2, \dots, S_m such that $\sum_{i \in S_j} a_i = B$ for every $j \in [m]$. Due to the constraints on a_1, a_2, \dots, a_{3m} , we note that a feasible partition S_1, S_2, \dots, S_m would have exactly 3 elements in each subset S_j .

Given an instance of 3- PARTITION by a set $S = [3m]$, weights a_1, a_2, \dots, a_{3m} , and bound B , we construct an instance (G, w, T) of ℓ_p -NORM-MULTIWAY-CUT as follows: we start with an empty graph G , and for each $i \in [3m]$, we add to G a subgraph as shown in Fig. 3. The edge weights are labelled near the corresponding edges, where $d := \lceil (12m + 12)^{\frac{1}{p-1}} \rceil$. These $3m$ subgraphs are disjoint from each other. Finally, we add m isolated vertices t_1, \dots, t_m to G . The terminal set T is given by $\{x'_i : i \in [3m], r \in [3]\} \cup \{t_1, \dots, t_m\}$. We observe that the graph G constructed this way is planar. We note that the size of (G, w, T) is polynomial in the size of the 3- PARTITION instance and the edge weights are polynomial in m, B , and a_1, \dots, a_{3m} . We emphasize that the number of terminals in this reduction is not a constant. The following lemma completes the proof of Theorem 4.2. \square

Lemma 4.2 *There exists a partition of S into S_1, \dots, S_m such that $\sum_{i \in S_j} a_i = B$ for each $j \in [m]$ if and only if (G, w, T) has a multiway cut whose ℓ_p -norm objective value is at most*

$$(9m(2dB)^p + mB^p)^{\frac{1}{p}}.$$

Proof We start by showing the forward direction.

Claim 4.4 *If S can be partitioned into S_1, \dots, S_m such that $\sum_{i \in S_j} a_i = B$ for each $j \in [m]$, then G has a multiway cut whose ℓ_p -norm objective value is $(9m(2dB)^p + mB^p)^{1/p}$.*

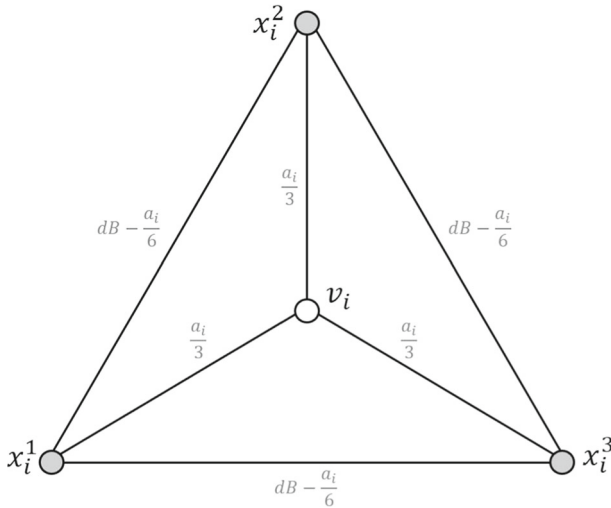


Fig. 3 The i th subgraph in the reduction from 3- PARTITION

Proof Consider the multiway cut of G defined by

$$\{x_i^r : i \in [3m], r \in [3]\} \cup \{t_j\} \cup \{v_i : i \in S_j\} : j \in [m].$$

The p th power of the ℓ_p -norm objective value of this multiway cut is

$$3m \cdot 3 \cdot (2dB)^p + \sum_{j \in [m]} \left(\sum_{i \in S_j} \frac{a_i}{3} \cdot 3 \right)^p = 9m(2dB)^p + \sum_{j \in [m]} B^p = 9m(2dB)^p + mB^p.$$

Therefore, we have a multiway cut with ℓ_p -norm objective value $(9m(2dB)^p + mB^p)^{1/p}$, completing the proof of the claim. \square

For the backward direction, we will start with a structural property regarding multiway cuts of (G, w, T) with cheap ℓ_p -norm objective value: each non-terminal vertex v_i will not be contained in a part that contains any of the x_i^r terminals.

Claim 4.5 *Let \mathcal{P} be a multiway cut in (G, w, T) with ℓ_p -norm objective value at most $(9m(2dB)^p + mB^p)^{1/p}$. Then, for every $i \in [3m]$, the vertex v_i will be in a part of \mathcal{P} that contains terminal t_j for some $j \in [m]$.*

Proof We will use X_i^r to denote the part in \mathcal{P} containing x_i^r for each $i \in [3m]$ and $r \in [3]$, and T_j to denote the part in \mathcal{P} containing t_j for each $j \in [m]$. Let $I := \{i \in [3m] : v_i \in \bigcup_{i' \in [3m], r \in [3]} X_{i'}^r\}$ be the indices of vertices v_i that are contained in some part that contains an x_i^r terminal. Suppose for the sake of contradiction that $I \neq \emptyset$.

Let $i \in [3m], r \in [3]$. Then, we have

$$w(\delta(X_i^r)) \geq 2dB + \sum_{i' \in [3m]: v_{i'} \in X_i^r} \frac{a_{i'}}{3}. \tag{4}$$

For an arbitrarily fixed X_i^r such that $|X_i^r| \geq 3$, let $X_i^r = \{x_i^r, v_{i_1}, v_{i_2}, \dots, v_{i_\ell}\}$ for some $i_1, \dots, i_\ell \in [3m]$, where $\ell \geq 2$. In Proposition 4.2, if we choose

$$\begin{aligned} x_1 &= 2dB, \quad x_4 = 2dB + \sum_{q=1}^{\ell} \frac{a_{i_q}}{3}, \\ x_2 &= \min \left\{ 2dB + \frac{a_{i_\ell}}{3}, 2dB + \sum_{q=1}^{\ell-1} \frac{a_{i_q}}{3} \right\}, \\ x_3 &= \max \left\{ 2dB + \frac{a_{i_\ell}}{3}, 2dB + \sum_{q=1}^{\ell-1} \frac{a_{i_q}}{3} \right\}, \end{aligned}$$

then we have

$$\left(2dB + \sum_{q=1}^{\ell} \frac{a_{i_q}}{3} \right)^p + (2dB)^p \geq \left(2dB + \sum_{q=1}^{\ell-1} \frac{a_{i_q}}{3} \right)^p + \left(2dB + \frac{a_{i_\ell}}{3} \right)^p.$$

By applying this argument $\ell - 1$ times, we get

$$\begin{aligned} &\left(2dB + \sum_{q=1}^{\ell} \frac{a_{i_q}}{3} \right)^p + (\ell - 1)(2dB)^p \\ &\geq \left(2dB + \sum_{q=1}^{\ell-1} \frac{a_{i_q}}{3} \right)^p + \left(2dB + \frac{a_{i_\ell}}{3} \right)^p + (\ell - 2)(2dB)^p \\ &\geq \dots \\ &\geq \left(2dB + \sum_{q=1}^2 \frac{a_{i_q}}{3} \right)^p + \left(2dB + \frac{a_{i_3}}{3} \right)^p + \dots + \left(2dB + \frac{a_{i_\ell}}{3} \right)^p + (2dB)^p \\ &\geq \sum_{q=1}^{\ell} \left(2dB + \frac{a_{i_q}}{3} \right)^p. \end{aligned} \tag{5}$$

We will divide parts X_i^r into three categories by defining

$$\begin{aligned} \mathcal{X}_1 &:= \{X_i^r : i \in [3m], r \in [3], |X_i^r| = 1\}, \\ \mathcal{X}_2 &:= \{X_i^r : i \in [3m], r \in [3], |X_i^r| = 2\}, \end{aligned}$$

$$\mathcal{X}_3 := \{X_i^r : i \in [3m], r \in [3], |X_i^r| \geq 3\}.$$

Moreover, let us define two subsets of I by

$$I_2 := \left\{ i' \in I : v_{i'} \in \bigcup_{X_i^r \in \mathcal{X}_2} X_i^r \right\}, \quad I_3 := \left\{ i' \in I : v_{i'} \in \bigcup_{X_i^r \in \mathcal{X}_3} X_i^r \right\}.$$

We note that (I_2, I_3) form a partition of I , $|I_2| = |\mathcal{X}_2|$, and $|\mathcal{X}_1| + |\mathcal{X}_2| + |\mathcal{X}_3| = 9m$.

Then the contribution of sets X_i^r to the p th power of the ℓ_p -norm objective value is given by

$$\begin{aligned} \sum_{i \in [3m], r \in [3]} w(\delta(X_i^r))^p &= \sum_{X_i^r \in \mathcal{X}_1} w(\delta(X_i^r))^p + \sum_{X_i^r \in \mathcal{X}_2} w(\delta(X_i^r))^p + \sum_{X_i^r \in \mathcal{X}_3} w(\delta(X_i^r))^p \\ &= |\mathcal{X}_1|(2dB)^p + \sum_{X_i^r \in \mathcal{X}_2} w(\delta(X_i^r))^p + \sum_{X_i^r \in \mathcal{X}_3} w(\delta(X_i^r))^p. \end{aligned} \tag{6}$$

By applying (4) and (5) to members of \mathcal{X}_3 , we get

$$\begin{aligned} \sum_{X_i^r \in \mathcal{X}_3} w(\delta(X_i^r))^p &\geq \sum_{X_i^r \in \mathcal{X}_3} \left(2dB + \sum_{i' \in [3m]: v_{i'} \in X_i^r} \frac{a_{i'}}{3} \right)^p \tag{by(4)} \\ &= \sum_{X_i^r \in \mathcal{X}_3} \left(\left(2dB + \sum_{i' \in [3m]: v_{i'} \in X_i^r} \frac{a_{i'}}{3} \right)^p + (|X_i^r| - 2)(2dB)^p \right) \\ &\quad - \sum_{X_i^r \in \mathcal{X}_3} (|X_i^r| - 2)(2dB)^p \\ &\geq \sum_{X_i^r \in \mathcal{X}_3} \sum_{i' \in [3m]: v_{i'} \in X_i^r} \left(2dB + \frac{a_{i'}}{3} \right)^p - \sum_{X_i^r \in \mathcal{X}_3} (|X_i^r| - 2)(2dB)^p \tag{by(5)} \\ &= \sum_{i' \in I_3} \left(2dB + \frac{a_{i'}}{3} \right)^p - \sum_{X_i^r \in \mathcal{X}_3} (|X_i^r| - 1)(2dB)^p + |\mathcal{X}_3|(2dB)^p \\ &= \sum_{i' \in I_3} \left(2dB + \frac{a_{i'}}{3} \right)^p - |I_3|(2dB)^p + |\mathcal{X}_3|(2dB)^p. \end{aligned}$$

Moreover, we observe that by Eq. (4) and definition of I_2 ,

$$\sum_{X_i^r \in \mathcal{X}_2} w(\delta(X_i^r))^p \geq \sum_{X_i^r \in \mathcal{X}_2} \left(2dB + \sum_{i' \in [3m]: v_{i'} \in X_i^r} \frac{a_{i'}}{3} \right)^p = \sum_{i' \in I_2} \left(2dB + \frac{a_{i'}}{3} \right)^p.$$

Therefore, (6) implies that

$$\begin{aligned}
 & \sum_{i \in [3m], r \in [3]} w(\delta(X_i^r))^p \\
 &= |\mathcal{X}_1|(2dB)^p + \sum_{X_i^r \in \mathcal{X}_2} w(\delta(X_i^r))^p + \sum_{X_i^r \in \mathcal{X}_3} w(\delta(X_i^r))^p \\
 &\geq |\mathcal{X}_1|(2dB)^p + \sum_{i' \in I_2} \left(2dB + \frac{a_{i'}}{3}\right)^p \\
 &\quad + \sum_{i' \in I_3} \left(2dB + \frac{a_{i'}}{3}\right)^p - |I_3|(2dB)^p + |\mathcal{X}_3|(2dB)^p \\
 &= |\mathcal{X}_1|(2dB)^p + \sum_{i' \in I} \left(2dB + \frac{a_{i'}}{3}\right)^p - |I_3|(2dB)^p + |\mathcal{X}_3|(2dB)^p \\
 &= \sum_{i' \in I} \left(2dB + \frac{a_{i'}}{3}\right)^p + (|\mathcal{X}_1| + |\mathcal{X}_3|)(2dB)^p - |I_3|(2dB)^p \\
 &= \sum_{i' \in I} \left(2dB + \frac{a_{i'}}{3}\right)^p + (9m - |\mathcal{X}_2|)(2dB)^p - |I_3|(2dB)^p \\
 &= \sum_{i' \in I} \left(2dB + \frac{a_{i'}}{3}\right)^p + (9m - |I_2|)(2dB)^p - |I_3|(2dB)^p \\
 &= \sum_{i' \in I} \left(2dB + \frac{a_{i'}}{3}\right)^p + (9m - |I|)(2dB)^p \\
 &\geq \sum_{i' \in I} \left(2dB + \frac{B}{12}\right)^p + (9m - |I|)(2dB)^p \quad (\text{since } a_{i'} \geq \frac{B}{4} \forall i' \in [3m]) \\
 &= |I| \left(2d + \frac{1}{12}\right)^p B^p + (9m - |I|)(2dB)^p.
 \end{aligned}$$

Assuming $I \neq \emptyset$, the p th power of the ℓ_p -norm objective value of this multiway cut is

$$\begin{aligned}
 & \sum_{i \in [3m], r \in [3]} w(\delta(X_i^r))^p + \sum_{j=1}^m w(\delta(T_j))^p \\
 &\geq |I| \left(2d + \frac{1}{12}\right)^p B^p + (9m - |I|)(2dB)^p + \sum_{i \notin I} w(\delta(v_i))^p \\
 &= |I| \left(2d + \frac{1}{12}\right)^p B^p + (9m - |I|)(2dB)^p + \sum_{i \notin I} a_i^p \\
 &\geq |I| \left(2d + \frac{1}{12}\right)^p B^p + (9m - |I|)(2dB)^p + (3m - |I|) \left(\frac{B}{4}\right)^p \\
 &= |I| \left(\left(2d + \frac{1}{12}\right)^p - (2d)^p - \left(\frac{1}{4}\right)^p \right) B^p + 9m(2d)^p B^p + 3m \left(\frac{1}{4}\right)^p B^p.
 \end{aligned}$$

Since we assumed that the ℓ_p -norm objective value of this multiway cut is at most $(9m(2dB)^p + mB^p)^{1/p}$, we have

$$\begin{aligned} 0 &\geq \left(\sum_{i \in [3m], r \in [3]} w(\delta(X_i^r))^p + \sum_{j=1}^m w(\delta(T_j))^p - (9m(2dB)^p + mB^p) \right) B^{-p} \\ &\geq |I| \left(\left(2d + \frac{1}{12}\right)^p - (2d)^p - \left(\frac{1}{4}\right)^p \right) + 3m \left(\frac{1}{4}\right)^p - m \\ &> |I| \left(\left(2d + \frac{1}{12}\right)^p - (2d)^p - \left(\frac{1}{4}\right)^p \right) - m. \end{aligned}$$

We note that due to Proposition 4.1, we have

$$\left(2d + \frac{1}{12}\right)^p - (2d)^p \geq \frac{p(2d)^{p-1}}{12} \geq 2^{p-1} \cdot p(m+1).$$

This implies

$$0 > |I| \left(2^{p-1} \cdot p(m+1) - \left(\frac{1}{4}\right)^p \right) - m \geq 2^{p-1} \cdot p(m+1) - \left(\frac{1}{4}\right)^p - m \geq 0,$$

yielding a contradiction since one of the inequalities in the sequence is strict. Therefore, we must have $I = \emptyset$. □

We complete the proof of the backward direction by showing the following claim which derives a YES certificate for 3-PARTITION from an optimal multiway cut whose ℓ_p -norm objective value is at most $(9m(2dB)^p + mB^p)^{1/p}$. □

Claim 4.6 *Given a multiway cut \mathcal{P} of (G, w, T) whose ℓ_p -norm objective value is at most $(9m(2dB)^p + mB^p)^{1/p}$, let $S_j := \{i \in [3m] : v_i \text{ is in the same part as } t_j \text{ in } \mathcal{P}\}$ for each $j \in [m]$. Then (S_1, S_2, \dots, S_m) is a partition of $S = [3m]$ such that $\sum_{i \in S_j} a_i = B$ for each $j \in [m]$.*

Proof By Claim 4.5, we know that S_1, \dots, S_j must form a partition of S . This also implies that for each $i \in [3m], r \in [3]$, the set $\{x_i^r\}$ is a part in the multiway cut \mathcal{P} . Therefore, the ℓ_p -norm objective value of \mathcal{P} is at least

$$\sum_{i \in [3m], r \in [3]} (2dB)^p + \sum_{j=1}^m w(\delta(S_j))^p = 9m(2dB)^p + \sum_{j=1}^m w(\delta(S_j))^p.$$

Since we know that the ℓ_p -norm objective value of \mathcal{P} is at most $(9m(2dB)^p + mB^p)^{1/p}$, it follows that $\sum_{j=1}^m w(\delta(S_j))^p \leq mB^p$. By Jensen’s inequality, we observe that

$$mB^p \geq \sum_{j=1}^m w(\delta(S_j))^p = \sum_{j=1}^m \left(\sum_{i \in S_j} a_i \right)^p \geq m \cdot \left(\frac{1}{m} \sum_{j=1}^m \sum_{i \in S_j} a_i \right)^p = mB^p.$$

Hence, all inequalities above should be equations. This happens only when $\sum_{i \in S_j} a_i = B$ for all $j \in [m]$. □

5 Convex Program and Integrality Gap

The following is a natural convex programming relaxation for ℓ_p -NORM-MULTIWAY-CUT on instance (G, w, T) where $T = \{t_1, \dots, t_k\}$ are the terminal vertices (the objective function can be convexified by introducing additional variables and constraints):

$$\begin{aligned} &\text{Minimize } \left(\sum_{i=1}^k \left(\sum_{uv \in E} w(uv) \cdot |x(u, i) - x(v, i)| \right)^p \right)^{1/p} \quad \text{subject to} \\ &\sum_{i=1}^k x(v, i) = 1 \quad \forall v \in V, \\ &x(t_i, i) = 1 \quad \forall i \in [k], \\ &x(v, i) \geq 0 \quad \forall v \in V, \forall i \in [k]. \end{aligned} \tag{7}$$

The above convex program is inspired by the formulation given by Călinescu et al. [7] for the case of $p = 1$. We note that determining the precise integrality gap of the formulation for $p = 1$ is still an open problem.

Lemma 5.1 *The convex program in (7) has an integrality gap of at least $k^{1-1/p}/2$.*

Proof Consider the star graph that has k leaves $\{t_1, \dots, t_k\}$ and a center vertex v with all edge weights being 1. Let the terminal vertices be the k leaves. The optimum ℓ_p -norm objective value of a multiway cut is

$$((k - 1)^p + k - 1)^{\frac{1}{p}},$$

and it corresponds to the partition $(\{t_1, v\}, \{t_2\}, \{t_3\}, \dots, \{t_k\})$. A feasible solution to the convex program (7) is given by $x(v, i) = 1/k$ for all $i \in [k]$ and $x(t_i, j) = 1[i = j]$ for all $i, j \in [k]$, which yields an objective of

$$\left(k \cdot \left(\frac{k - 1}{k} + (k - 1) \cdot \frac{1}{k} \right)^p \right)^{\frac{1}{p}} = \frac{2k - 2}{k} \cdot k^{\frac{1}{p}}.$$

This results in an integrality gap of at least

$$\frac{((k - 1)^p + k - 1)^{\frac{1}{p}}}{\frac{2k - 2}{k} \cdot k^{\frac{1}{p}}} \geq \frac{k - 1}{2k - 2} = \frac{k^{1 - \frac{1}{p}}}{2}.$$

□

Bansal et al. give an SDP relaxation for MIN-MAX-MULTIWAY-CUT and show that the star graph has an integrality gap of $\Omega(k)$ for this SDP relaxation. This SDP relaxation can be generalized in a natural fashion to ℓ_p -NORM-MULTIWAY-CUT. The star graph still exhibits an integrality gap of $\Omega(k^{1-1/p})$ for the generalized SDP relaxation for ℓ_p -NORM-MULTIWAY-CUT.

6 Inapproximability

In this section, we show that ℓ_p -NORM-MULTIWAY-CUT does not admit a $k^{1-1/p-\epsilon}$ -approximation assuming the small set expansion hypothesis. In contrast, there is a trivial $O(k^{1-1/p})$ -approximation (see Sect. 6.1). We mention that the inapproximability result in this section is similar to the result of Bansal et al. [3] who showed that MIN-MAX-MULTIWAY-CUT does not admit a $k^{1-\epsilon}$ -approximation assuming the small set expansion hypothesis. We adapt the same ideas for ℓ_p -NORM-MULTIWAY-CUT.

To prove our results, we consider MIN-SUM-EQUI-K-PARTITIONING: the input to this problem is a graph $G = (V, E)$ (where $n := |V|$), an edge weight function $w : E \rightarrow \mathbb{Q}_+$, and an integer $k \leq n$ such that k divides n . The goal is to partition V into k sets P_1, \dots, P_k such that $|P_i| = n/k$ for all $i \in [k]$ so as to minimize $\sum_{i=1}^k w(\delta(P_i))$. We will use λ to denote the optimum objective value of MIN-SUM-EQUI-K-PARTITIONING. A partition (P_1, \dots, P_k) of V is an (α, β) -bicriteria approximation for MIN-SUM-EQUI-K-PARTITIONING if $\sum_{i=1}^k w(\delta(P_i)) \leq \alpha\lambda$ and $|P_i| \leq \beta(n/k)$ for all $i \in [k]$. For a sufficiently large constant k , it is known that a (α, β) -bicriteria approximation for MIN-SUM-EQUI-K-PARTITIONING where α is a constant that could depend on k and β is an absolute constant is at least as hard as small set expansion [20]. We show the following result which implies that a $k^{1-1/p-\epsilon}$ -approximation is unlikely for ℓ_p -NORM-MULTIWAY-CUT (by setting $k = k(\epsilon)$ to be a large constant):

Theorem 6.1 *If ℓ_p -NORM-MULTIWAY-CUT admits an efficient $k^{1-1/p-\epsilon}$ -approximation algorithm for some constant $\epsilon > 0$, then for $k \geq 9^{2/\epsilon}$ MIN-SUM-EQUI-K-PARTITIONING admits a $(O(k^{2-1/p}), 3^{4/\epsilon})$ -bicriteria approximation.*

Our proof of Theorem 6.1 proceeds via the following lemma (which is the counterpart to Lemma 5.1 of [3], but for ℓ_p -NORM-MULTIWAY-CUT).

Lemma 6.1 *If ℓ_p -NORM-MULTIWAY-CUT has a polynomial-time γ -approximation algorithm, then MIN-SUM-EQUI-K-PARTITIONING has an efficient $(5k\gamma, 9\gamma k^{1/p})$ -bicriteria approximation algorithm.*

Proof We will follow the reduction designed by Bansal et al in Lemma 5.1 of [3]. Let (G, w, k) be an instance of MIN-SUM-EQUI-K-PARTITIONING, and let λ refer to the optimum objective value of MIN-SUM-EQUI-K-PARTITIONING on instance (G, w, k) . We will assume knowledge of a value $B \in [\lambda, 2\lambda]$ by binary search. We construct an instance $(G' = (V', E'), w', T)$ of ℓ_p -NORM-MULTIWAY-CUT as follows.

$$\begin{aligned} V' &:= V \cup \{t_1, \dots, t_k\}, \\ E' &:= E \cup \{t_i v : v \in V\}, \end{aligned}$$

$$T := \{t_1, \dots, t_k\},$$

$$w'(e) := \begin{cases} w(e) & \text{if } e \in E, \\ \frac{B}{n} & \text{if } e \in E' \setminus E. \end{cases}$$

We will use OPT to refer to the optimum ℓ_p -norm objective value of ℓ_p -NORM-MULTIWAY-CUT on instance (G', w', T) . The following claim completes the proof of Lemma 6.1. □

Claim 6.1 *If $\mathcal{P}' = (P'_1, \dots, P'_k)$ is a multiway cut on instance (G', w', T) with ℓ_p -norm objective value at most $\gamma \cdot \text{OPT}$, then the partition $\mathcal{P} = (P_1, \dots, P_k)$ of V defined by $P_i = P'_i \cap V$ for all $i \in [k]$ is a $(5k\gamma, 9\gamma k^{1/p})$ -bicriteria approximate optimum to MIN-SUM-EQUI-K-PARTITIONING.*

Proof Let $\mathcal{Q} = (Q_1, \dots, Q_k)$ be an optimum solution to MIN-SUM-EQUI-K-PARTITIONING on instance (G, w, k) , and let $\mathcal{Q}' = (Q'_1, \dots, Q'_k)$ be a partition of $V(G')$ obtained by $Q'_i := Q_i \cup \{t_i\}$ for each $i \in [k]$. Then, \mathcal{Q}' is a multiway cut for (G', w', T) and the ℓ_p -norm objective value of \mathcal{Q}' raised to p th power is

$$\begin{aligned} \sum_{i=1}^k w'(\delta_{G'}(Q'_i))^p &= \sum_{i=1}^k \left(w(\delta_G(Q_i)) + \frac{n}{k} \cdot (k-1) \cdot \frac{B}{n} + \frac{n}{k} \cdot (k-1) \cdot \frac{B}{n} \right)^p \\ &= \sum_{i=1}^k \left(w(\delta_G(Q_i)) + \frac{k-1}{k} \cdot 2B \right)^p \geq \text{OPT}^p, \end{aligned} \tag{8}$$

where the first $(n/k)(k-1)(B/n)$ term represents the cost of edges between Q_i and $T - \{t_i\}$, and the second $(n/k)(k-1)(B/n)$ term represents the cost of edges between $V - Q_i$ and t_i . We recall that OPT refers to the optimum ℓ_p -norm objective value of ℓ_p -NORM-MULTIWAY-CUT on instance (G', w', T) .

Since \mathcal{P}' is a γ -approximate optimum solution to ℓ_p -NORM-MULTIWAY-CUT, we have

$$\begin{aligned} \gamma^p \cdot \text{OPT}^p &\geq \sum_{i=1}^k w'(\delta_{G'}(P'_i))^p \\ &= \sum_{i=1}^k \left(w(\delta_G(P_i)) + |P_i|(k-1)\frac{B}{n} + (n - |P_i|)\frac{B}{n} \right)^p \\ &= \sum_{i=1}^k \left(w(\delta_G(P_i)) + B + (k-2)|P_i|\frac{B}{n} \right)^p \\ &\geq k^{1-p} \left(\sum_{i=1}^k w(\delta_G(P_i)) + (2k-2)B \right)^p. \quad (\text{by Jensen's inequality}) \end{aligned} \tag{9}$$

Hence,

$$\begin{aligned}
 & k^{1-p} \left(\sum_{i=1}^k w(\delta_G(P_i)) + (2k - 2)B \right)^p \\
 & \leq \gamma^p \cdot \text{OPT}^p \\
 & \leq \gamma^p \sum_{i=1}^k \left(w(\delta_G(Q_i)) + \frac{k-1}{k} \cdot 2B \right)^p \quad (\text{by (8)}) \\
 & \leq \gamma^p \sum_{i=1}^k \left(\lambda + \frac{k-1}{k} \cdot 2B \right)^p \\
 & = \gamma^p k \left(\lambda + \frac{k-1}{k} \cdot 2B \right)^p. \tag{10}
 \end{aligned}$$

This inequality is equivalent to

$$\sum_{i=1}^k w(\delta_G(P_i)) + (2k - 2)B \leq k\gamma \left(\lambda + \frac{k-1}{k} \cdot 2B \right).$$

Using the assumption that $B \in [\lambda, 2\lambda]$, we have

$$\sum_{i=1}^k w(\delta_G(P_i)) \leq k\gamma \left(\lambda + \frac{k-1}{k} \cdot 2B \right) < k\gamma(\lambda + 2B) \leq 5k\gamma\lambda.$$

Inequalities (9) and (10) also imply that for every $j \in [k]$,

$$\begin{aligned}
 \left((k-2)|P_j| \frac{B}{n} \right)^p & \leq \sum_{i=1}^k \left(w(\delta_G(P_i)) + B + (k-2)|P_i| \frac{B}{n} \right)^p \\
 & \leq \gamma^p \cdot \text{OPT}^p \quad (\text{by (9)}) \\
 & \leq \gamma^p \cdot k \left(\lambda + \frac{k-1}{k} \cdot 2B \right)^p \quad (\text{by (10)}) \\
 & \leq \gamma^p \cdot k(3B)^p.
 \end{aligned}$$

This implies that

$$|P_j| \leq 3\gamma k^{\frac{1}{p}} \frac{n}{k-2} \leq 9\gamma k^{\frac{1}{p}} \frac{n}{k}.$$

□

We will use the following lemma from [3] to prove Theorem 6.1.

Lemma 6.2 [3] *If MIN-SUM-EQUI-K-PARTITIONING has an efficient $(\alpha, k^{1-\epsilon})$ -bicriteria approximation algorithm for some $\epsilon > 0$, then MIN-SUM-EQUI-K-PARTITIONING also has an efficient $(\alpha \log \log k, 3^{2/\epsilon})$ -bicriteria approximation algorithm.*

We now prove Theorem 6.1.

Proof of Theorem 6.1 If ℓ_p -NORM-MULTIWAY-CUT has an efficient $k^{1-1/p-\epsilon}$ -approximation algorithm for some $\epsilon > 0$, then Lemma 6.1 implies that MIN-SUM-EQUI-K-PARTITIONING has an efficient $(5k^{2-1/p-\epsilon}, 9k^{1-\epsilon})$ -bicriteria approximation algorithm. For $k \geq 9^{2/\epsilon}$, we have $9k^{1-\epsilon} \leq k^{1-\epsilon/2}$. Then, Lemma 6.2 implies that MIN-SUM-EQUI-K-PARTITIONING has a $(O(k^{2-1/p}), 3^{4/\epsilon})$ -bicriteria approximation. This completes the proof of Theorem 6.1. \square

6.1 A Trivial $O(k^{1-1/p})$ -Approximation via Isolating Cuts

In this section, we show a trivial approximation algorithm for ℓ_p -NORM-MULTIWAY-CUT that achieves an approximation factor of $(k^{p-1} + 1)^{1/p}$. Our algorithm follows a high-level 3-step recipe similar to the one in Sect. 3—namely, (I) isolate terminals, (II) uncross, and (III) aggregate. Given an instance (G, w, T) of ℓ_p -NORM-MULTIWAY-CUT, let the set T of terminals be $\{t_1, \dots, t_k\}$. For each $i \in [k]$, we compute a minimum $(t_i, T - t_i)$ -cut, say $(S_i, V - S_i)$. The sets S_1, \dots, S_k can be uncrossed via posimodularity to ensure that each $(S_i, V - S_i)$ is still a minimum $(t_i, T - t_i)$ -cut and moreover $S_i \cap S_j = \emptyset$ for all distinct $i, j \in [k]$. Let $R := V - \cup_{i=1}^k S_i$, and we assume without loss of generality that $w(\delta(S_1)) = \max_{i \in [k]} w(\delta(S_i))$. We will show that the multiway cut $(S_1 \cup R, S_2, S_3, \dots, S_k)$ is a $(k^{p-1} + 1)^{1/p}$ -approximation for ℓ_p -NORM-MULTIWAY-CUT. We note that this same algorithm achieves a 2-approximation factor for MULTIWAY-CUT [13].

Let (P_1, \dots, P_k) denote an optimum solution for ℓ_p -NORM-MULTIWAY-CUT. Since $(S_i, V - S_i)$ is a min $(t_i, T - t_i)$ -cut, we have that $w(\delta(S_i)) \leq w(\delta(P_i))$. We also note that $w(\delta(S_1 \cup R)) = w(\delta(\cup_{i=2}^k S_i)) \leq \sum_{i=2}^k w(\delta(S_i))$. Let us consider the p th power of the ℓ_p -norm objective value of $(S_1 \cup R, S_2, \dots, S_k)$:

$$\begin{aligned} & w(\delta(S_1 \cup R))^p + \sum_{i=2}^k w(\delta(S_i))^p \\ & \leq \left(\sum_{i=2}^k w(\delta(S_i)) \right)^p + \sum_{i=2}^k w(\delta(S_i))^p \\ & \leq (k-1)^{p-1} \left(\sum_{i=2}^k w(\delta(S_i))^p \right) + \sum_{i=2}^k w(\delta(S_i))^p \quad (\text{by Jensen}) \\ & = ((k-1)^{p-1} + 1) \sum_{i=2}^k w(\delta(S_i))^p \\ & \leq ((k-1)^{p-1} + 1) \left(1 - \frac{1}{k} \right) \sum_{i=1}^k w(\delta(S_i))^p \quad (\text{by choice of } S_1) \end{aligned}$$

$$\leq (k^{p-1} + 1) \sum_{i=1}^k w(\delta(P_i))^p.$$

Hence, the ℓ_p -norm objective value of $(S_1 \cup R, S_2, \dots, S_k)$ is within a $(k^{p-1} + 1)^{1/p}$ -factor of the optimum ℓ_p -norm objective value.

7 Conclusion

In this work, we introduced ℓ_p -NORM-MULTIWAY-CUT for $p \geq 1$ as a unified generalization of MULTIWAY-CUT and MIN-MAX-MULTIWAY-CUT. We showed that ℓ_p -NORM-MULTIWAY-CUT is NP-hard for constant number of terminals or in planar graphs for every $p \geq 1$. The natural convex program for ℓ_p -NORM-MULTIWAY-CUT has an integrality gap of $\Omega(k^{1-1/p})$ and the problem is $(k^{1-1/p-\epsilon})$ -inapproximable for any constant $\epsilon > 0$ and $k > k(\epsilon)$ assuming the small set expansion hypothesis, where k is the number of terminals in the input instance. The inapproximability result suggests that a dependence on n in the approximation factor is unavoidable if we would like to obtain an approximation factor that is better than the trivial $O(k^{1-1/p})$ -factor. On the algorithmic side, we gave an $O(\sqrt{\log^3 n \log k})$ -approximation (i.e., an $O(\log^2 n)$ -approximation), where n is the number of vertices in the input graph. Our results suggest that the approximability behaviour of ℓ_p -NORM-MULTIWAY-CUT exhibits a sharp transition from $p = 1$ to $p > 1$. Our work raises several open questions. We mention a couple of them:

1. Can we achieve an $O(\log n)$ -approximation for ℓ_p -NORM-MULTIWAY-CUT for every $p \geq 1$? We recall that when $p = \infty$, the current best approximation factor is indeed $O(\log n)$ [3].
2. Is there a polynomial-time algorithm for ℓ_p -NORM-MULTIWAY-CUT for any given p that achieves an approximation factor that smoothly interpolates between the best upper bounds on the approximation factors for $p = 1$ and $p = \infty$ —e.g., is there an $O(\log^{1-1/p} n)$ -approximation?

Acknowledgements Karthik would like to thank Ali Bibak for initial discussions on parts of this work.

References

1. Ahmadi, S., Khuller, S., Saha, B.: Min–Max Correlation Clustering via Multicut, pp. 13–26. Integer Programming and Combinatorial Optimization, IPCO, Atlanta (2019)
2. Angelidakis, H., Makarychev, Y., Manurangsi, P.: An Improved Integrality Gap for the Călinescu–Karloff–Rabani Relaxation for Multiway Cut, pp. 39–50. Integer Programming and Combinatorial Optimization, IPCO, Atlanta (2017)
3. Bansal, N., Feige, U., Krauthgamer, R., Makarychev, K., Nagarajan, V., Naor, J., Schwartz, R.: Min–max graph partitioning and small set expansion. *SIAM J. Comput.* **43**(2), 872–904 (2014)
4. Bérczi, K., Chandrasekaran, K., Király, T., Madan, V.: Improving the integrality gap for multiway cut. *Math. Program.* **183**, 171–193 (2020)
5. Buchbinder, N., Naor, J., Schwartz, R.: Simplex partitioning via exponential clocks and the multiway cut problem. In: Proceedings of the Forty-Fifth Annual ACM Symposium on Theory of Computing, STOC, pp. 535–544 (2013)

6. Buchbinder, N., Schwartz, R., Weizman, B.: Simplex transformations and the multiway cut problem. In: Proceedings of the Twenty-Eighth Annual ACM-SIAM Symposium on Discrete Algorithms, SODA, pp. 2400–2410 (2017)
7. Călinescu, G., Karloff, H., Rabani, Y.: An improved approximation algorithm for multiway cut. *J. Comput. Syst. Sci.* **60**(3), 564–574 (2000)
8. Chandrasekaran, K., Chekuri, C.: Min–max partitioning of hypergraphs and symmetric submodular functions. In: Proceedings of the Thirty-Second Annual ACM-SIAM Symposium on Discrete Algorithms, SODA, pp. 1026–1038 (2021)
9. Charikar, M., Gupta, N., Schwartz, R.: Local Guarantees in Graph Cuts and Clustering, pp. 136–147. *Integer Programming and Combinatorial Optimization, IPCO, Atlanta* (2017)
10. Chekuri, C., Ene, A.: Submodular cost allocation problem and applications. In: *International Colloquium on Automata, Languages and Programming, ICALP*, pp. 354–366 (2011)
11. Cheung, K., Cunningham, W., Tang, L.: Optimal 3-terminal cuts and linear programming. *Math. Program.* **106**(1), 1–23 (2006)
12. Chvátal, V.: Recognizing decomposable graphs. *J. Graph Theory* **8**, 51–53 (1984)
13. Dahlhaus, E., Johnson, D., Papadimitriou, C., Seymour, P., Yannakakis, M.: The complexity of multi-terminal cuts. *SIAM J. Comput.* **23**(4), 864–894 (1994)
14. Garey, M.R., Johnson, D.S.: *Computers and Intractability: A Guide to the Theory of NP-Completeness*. Series of Books in the Mathematical Sciences, W.H. Freeman, San Francisco (1979)
15. Kalhan, S., Makarychev, K., Zhou, T.: Correlation clustering with local objectives. *Adv. Neural Inf. Process. Syst.* **32**, 9346–9355 (2019)
16. Karger, D., Klein, P., Stein, C., Thorup, M., Young, N.: Rounding algorithms for a geometric embedding of minimum multiway cut. *Math. Oper. Res.* **29**(3), 436–461 (2004)
17. Manokaran, R., Naor, J., Raghavendra, P., Schwartz, R.: SDP gaps and UGC hardness for multiway cut, 0-extension, and metric labeling. In: *Proceedings of the Fortieth Annual ACM Symposium on Theory of Computing, STOC*, pp. 11–20 (2008)
18. Marx, D.: Parameterized graph separation problems. *Theor. Comput. Sci.* **351**(3), 394–406 (2006)
19. Puleo, G., Milenkovic, O.: Correlation clustering and biclustering with locally bounded errors. *IEEE Trans. Inf. Theory* **64**, 4105–4119 (2018)
20. Raghavendra, P., Steurer, D., Tulsiani, M.: Reductions between expansion problems. In: *IEEE Conference on Computational Complexity, CCC*, pp. 64–73 (2012)
21. Sharma, A., Vondrák, J.: Multiway cut, pairwise realizable distributions, and descending thresholds. In: *Proceedings of the Forty-Sixth Annual ACM Symposium on Theory of Computing, STOC*, pp. 724–733 (2014)
22. Svitkina, Z., Tardos, É.: Min–max multiway cut. In: *Approximation, Randomization, and Combinatorial Optimization. Algorithms and Techniques, APPROX*, pp. 207–218 (2004)

Publisher's Note Springer Nature remains neutral with regard to jurisdictional claims in published maps and institutional affiliations.