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# Online maximum $k$ -coverage\*

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## Abstract

We study an online model for the maximum  $k$ -vertex-coverage problem, where given a graph  $G = (V, E)$  and an integer  $k$ , we ask for a subset  $A \subseteq V$ , such that  $|A| = k$  and the number of edges covered by  $A$  is maximized. In our model, at each step  $i$ , a new vertex  $v_i$  is revealed, and we have to decide whether we will keep it or discard it. At any time of the process, only  $k$  vertices can be kept in memory; if at some point the current solution already contains  $k$  vertices, any inclusion of any new vertex in the solution must entail the irremediable deletion of one vertex of the current solution (a vertex not kept when revealed is irremediably deleted). We propose algorithms for several natural classes of graphs (mainly regular and bipartite), improving on an easy  $\frac{1}{2}$ -competitive ratio. We next settle a set-version of the problem, called maximum  $k$ -(set)-coverage problem. For this problem we present an algorithm that improves upon former results for the same model for small and moderate values of  $k$ .

## 1 Introduction

In the *maximum  $k$ -vertex-coverage* ( $MkVC$ ) problem we are given a graph  $G = (V, E)$  ( $|V| = n$ ,  $|E| = m$ ) and an integer  $k$ , we ask for a subset  $A \subseteq V$ , such that  $|A| = k$  and the number of edges covered by  $A$  is maximized. The  $MkVC$  problem is NP-hard, since otherwise the optimal solution for the vertex cover problem could be found in polynomial time: for each  $k$ ,  $1 \leq k \leq n$ , run the algorithm for the  $MkVC$  problem and stop when all elements are covered.

In this paper we consider the following online model for this problem: at each step  $i$ , a new vertex  $v_i$  with its adjacent edges is revealed, and we have to decide whether we will include  $v_i$  in the solution or discard it. At any time of the process, only  $k$  vertices can be kept in memory, so if at some point the

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current solution already contains  $k$  vertices, any inclusion of any new vertex in the solution must be compensated with the irremediable deletion of one vertex of the current solution. Of course, a vertex that is not kept when it is revealed is also irremediably deleted. To our knowledge, no online model for the  $MkVC$  problem has been studied until now.

A generalization of the  $MkVC$  problem is the *maximum  $k$ -(set)-coverage* ( $MkC$ ) problem, where given a universe of elements  $E = \{e_1, e_2, \dots, e_m\}$ , a collection of subsets of  $E$ ,  $S = \{S_1, S_2, \dots, S_n\}$ , and an integer  $k \leq n$ , we ask for a subcollection  $A = \{A_1, A_2, \dots, A_{|A|}\} \subseteq S$ , such that  $|A| = k$  and the number of elements of  $E$  covered by  $A$  is maximized. The online model for the  $MkC$  problem is the same as for the  $MkVC$ .

Clearly, the  $MkVC$  problem is a special case of the  $MkC$  problem where: (i) each element belongs to exactly two sets and (ii) the intersection of any two sets of  $S$  has size at most one, since multiple edges are not permitted.

The weighted generalization of the  $MkC$  problem, denoted by **WEIGHTED  $MkC$** , has been also studied in the literature. In this problem, each element  $e_i \in E$  has a non-negative weight  $w(e_i)$ , and the goal is to maximize the total weight of the elements covered by  $k$  sets.

The analogous online model for **WEIGHTED  $MkC$**  problem, where at each step  $i$  a set  $S_i \in S$  together with its elements is revealed and only  $k$  such sets can be kept in memory, has been studied in [7], where an algorithm of competitive ratio  $\frac{1}{4}$  is given. The authors in their so called *set-streaming model* assume that the set of elements is known a priori. Nevertheless, they do not use this information in the proposed algorithm.

In the classical offline setting, the  $MkC$  problem is known to be non approximable within a factor  $1 - \frac{1}{e}$  [2]. On the other hand, even for the weighted version of the problem, an approximation algorithm of ratio  $1 - (1 - \frac{1}{k})^k$  is known [5]. This ratio tends to  $1 - \frac{1}{e}$  as  $k$  increases, closing in this way the approximability question for the problem.

In [1] the inverse problem (i.e., the hitting set version of  $MkC$ ), also called *maximum coverage problem*, has been studied: given a universe of elements  $E = \{e_1, e_2, \dots, e_m\}$ , a collection of subsets of  $E$ ,  $S = \{S_1, S_2, \dots, S_n\}$ , a non-negative weight  $w(S_i)$  for each  $S_i \in S$ , and an integer  $k$ , a set  $B \subseteq E$  is sought, such that  $|B| = k$  and the total weight of the sets in  $S$  that intersect with  $B$  is maximized. It is easy to see that this version is equivalent to the **WEIGHTED  $MkC$**  modulo the interchange of the roles between set-system and ground set. An algorithm of approximation ratio  $1 - (1 - \frac{1}{p})^p$  is presented in [1] for this problem, where  $p$  is the cardinality of the largest set in  $S$ . In the case where each set has cardinality equal to two then this problem coincides with the  $MkVC$  problem; hence a  $\frac{3}{4}$  approximation ratio is implied by the algorithm in [1]. Several improvements for some restricted cases of the  $MkVC$  problem are presented in [3, 4].

In this paper we study the online model described above for both the  $MkVC$  and the  $MkC$  problems. In Section 2, we prove several negative results on the competitiveness of any algorithm for the model handled for both problems. In

Section 3, we present algorithms for regular graphs, regular bipartite graphs, trees and chains, achieving non-trivial competitive ratios, improving upon an easy  $\frac{1}{2}$  competitiveness result holding for any graph. Finally, in Section 4 the  $k$ -(set)-coverage problem is handled. For this problem, we present an algorithm that improves upon former results for the same model for small and moderate values of  $k$ .

The following notations will be used in the sequel. They are based upon the definition of the  $MkVC$  problem and are easily extendable to the  $MkC$  problem.

For any  $A \subseteq V$ , we denote by  $E(A)$  the set of edges covered by  $A$  and by  $m(A) = |E(A)|$  the number of these edges. Let  $SOL = m(A)$  be the number of edges covered by our algorithms. Moreover, we denote by  $A^* \subseteq V$  an optimal subset of vertices and by  $OPT = m(A^*)$  the number of edges covered by an optimal solution. The maximum degree (or the degree when it is regular) of the input-graph  $G = (V, E)$  is denoted by  $\Delta$ . Dealing with  $MkC$ ,  $\Delta$  denotes the cardinality of a set that contains a maximum number of elements, that is,  $\Delta = \max\{|S_i| : 1 \leq i \leq n\}$ . For a subset  $A \subseteq V$  and a vertex  $v_i \in A$ , we call *public* the edges of  $v_i$  that are shared by another vertex in  $A$  and *private* the edges of  $v_i$  that are covered just by  $v_i$  in  $A$ . Finally, as it is common in the online setting, the quality of an algorithm is measured by means of the so-called *competitive ratio* representing the ratio of the value of the solution computed by the algorithm over the optimal value of the whole instance, i.e., the value of an optimal (offline) solution of the final instance.

In what follows, for reasons of papers length, some proofs are omitted. They can be found in appendix.

## 2 Negative results

In this section we give negative results for the online maximum  $k$ -vertex-coverage problem and their corresponding adaptations for the maximum  $k$ -coverage problem. We start with a negative result for the restrictive case where swaps are not allowed (replacement of a vertex or set that belongs to the current solution by the newly revealed vertex or set is not permitted).

**Proposition 1.** *Any deterministic online algorithm that does not allow swaps cannot achieve a competitive ratio better than:*

- $O\left(\frac{1}{(n-1)^{1/(k+1)}}\right)$  for the  $MkVC$  problem;
- $O\left(\frac{1}{m^{1/(k+1)}}\right)$  for the  $MkC$  problem.

The next negative result for the  $MkVC$  problem fits the model addressed in the paper (swaps are allowed).

**Proposition 2.** *Any deterministic online algorithm cannot achieve a competitive ratio better than  $\frac{2k}{3k-2} \simeq \frac{2}{3}$  for the  $MkVC$  problem.*

*Proof.* Assume that  $2k - 1$  vertices,  $v_1^1, v_2^1, \dots, v_{2k-1}^1$ , of degree one and  $2k - 1$  vertices,  $v_1^2, v_2^2, \dots, v_{2k-1}^2$ , of degree two are released, such that  $(v_i^1, v_i^2) \in E$ ,  $1 \leq i \leq 2k - 1$ , and that the algorithm selects  $k' \leq k$  of them. Wlog, let  $v_1^2, v_2^2, \dots, v_{k'}^2$  be the vertices selected by the algorithm. Next the vertex  $v_3$  of degree  $k'$  is released, where  $(v_i^2, v_3) \in E$ ,  $1 \leq i \leq k'$ . The solution of the algorithm at this time is  $2k'$ , while the inclusion or not of  $v_3$  does not play any role for this value. Finally,  $2k - 1 - k'$  vertices,  $v_{k'+1}^3, v_{k'+2}^3, \dots, v_{2k-1}^3$ , of degree one are released, such that  $(v_i^2, v_i^3) \in E$ ,  $k' + 1 \leq i \leq 2k - 1$ . In this last phase, the algorithm can increase its solution by at most  $k - k'$  more edges. Hence, the final solution of the algorithm is at most  $k + k'$ . The optimum solution consists of the vertices  $v_{k+1}^2, v_{k+2}^2, \dots, v_{2k-1}^2, v_3$ , and hence is of cardinality  $2(k - 1) + k'$ . In all,  $\frac{SOL}{OPT} = \frac{k+k'}{2(k-1)+k'} \leq \frac{2k}{3k-2}$ .  $\square$

An analogous result can be proved for the  $MkC$  problem. Recall that for the offline version of the  $MkC$  problem an  $1 - \frac{1}{e} \simeq 0.63$ -inapproximability result is known [2].

**Proposition 3.** *Any deterministic online algorithm cannot achieve a competitive ratio better than  $\frac{k+2\sqrt{k+1}}{2k+2\sqrt{k+1}} \simeq \frac{1}{2}$  for the  $MkC$  problem even in the case where all sets have the same cardinality.*

### 3 Maximum $k$ -vertex-coverage

In this section we deal with the online maximum  $k$ -vertex-coverage problem. Note, first, that there exists an easy  $\frac{1}{2}$ -competitive ratio for this problem. In fact, consider selecting  $k$  vertices of largest degrees. In an optimum solution all the edges are, at best, covered once, while in the solution created by this greedy algorithm, all the edges are, at worst, covered twice. Since the algorithm selects the largest degrees of the graph, the  $\frac{1}{2}$ -competitive ratio is immediately concluded.

**Proposition 4.** *There is a  $\frac{1}{2}$ -competitive ratio for the online  $MkVC$  problem.*

In the next sections we improve the  $\frac{1}{2}$ -competitive ratio for several classes of graphs. But first, we give an easy upper bound for the number of elements covered by any solution that will be used later. Its proof is straightforward.

**Proposition 5.**  $OPT \leq k\Delta$ .

#### 3.1 Regular graphs

The following preliminary result that will be used later holds for any algorithm for the  $MkVC$  problem in regular graphs.

**Proposition 6.** *Any deterministic online algorithm achieves a  $\frac{k}{n}$ -competitive ratio for the  $MkVC$  problem on regular graphs.*

Let us note that the result of Proposition 6 for the  $mkVC$  problem also holds for general graphs in the offline setting [4].

We now present an algorithm for the  $mkVC$  problem in regular graphs. Our algorithm depends on a parameter  $x$  which indicates the improvement on the current solution that a new vertex should entail, in order to be selected for inclusion in the solution. In other words, we replace a vertex of the current solution by the released one, only if the solution increases by at least  $\lceil \frac{\Delta}{x} \rceil$  edges.

---

**ALGORITHM  $mkVC\text{-R}(x)$**

```

1:  $A = \emptyset; B = \emptyset;$ 
2: for each released vertex  $v$  do
3:   if  $|A| < k$  then
4:      $A = A \cup \{v\};$ 
5:     if  $v$  increases the edges in  $B$  by at least  $\lceil \frac{\Delta}{x} \rceil$  then
6:        $B = B \cup \{v\};$ 
7:     else if  $|B| < k$  and  $v$  increases the edges in  $B$  by at least  $\lceil \frac{\Delta}{x} \rceil$  then
8:       Select a vertex  $u \in A \setminus B;$ 
9:        $A = A \cup \{v\} \setminus \{u\}; B = B \cup \{v\};$ 
10: return  $A;$ 

```

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As we will see in what follows, the best value for  $x$  is  $x = \frac{n+2k+\sqrt{4k^2+n^2}}{2n}$ , leading to the following theorem.

**Theorem 1.** **ALGORITHM  $mkVC\text{-R}$**  achieves 0.55-competitive ratio.

*Proof.* Note that  $B \subseteq A$  consists of the vertices that improve the solution by at least  $\lceil \frac{\Delta}{x} \rceil$ ;  $b$  denotes the number of these vertices, i.e.,  $b = |B|$ . We denote by  $y_1$  the number of edges with one endpoint in  $B$  and the other in  $V \setminus B$ , and by  $y_2$  the number of edges with both endpoints in  $B$ . By definition,

$$SOL \geq y_1 + y_2 = b\Delta - y_2 = \frac{b\Delta - y_1}{2} + y_1 = \frac{b\Delta + y_1}{2} \quad (1)$$

We shall handle two cases, depending on the value of  $b$  with respect to  $k$ .

If  $b < k$  then each vertex  $v \in V \setminus B$  is not selected by **ALGORITHM  $mkVC\text{-R}(x)$**  to be in  $B$  because it is adjacent to at most  $\lceil \frac{\Delta}{x} \rceil - 1$  vertices of  $V \setminus B$ . Thus, there are at least  $\Delta - \lceil \frac{\Delta}{x} \rceil + 1$  edges that connect  $v$  with vertices in  $B$ . Summing up for all the vertices in  $V \setminus B$ , it holds that  $y_1 \geq (n - b) (\Delta - \lceil \frac{\Delta}{x} \rceil + 1)$ , and considering also (1) we get:

$$SOL \geq (n - b) \left( \Delta - \left\lceil \frac{\Delta}{x} \right\rceil + 1 \right) + y_2 \quad (2)$$

$$SOL \geq \frac{b\Delta + (n - b) (\Delta - \lceil \frac{\Delta}{x} \rceil + 1)}{2} \quad (3)$$

Using the upper bound for the optimum provided by Proposition 5 and expressions (2) and (3), respectively, we get the following ratios:

$$\begin{aligned} \frac{SOL}{OPT} &\geq \frac{(n-b)(\Delta - \lceil \frac{\Delta}{x} \rceil + 1) + y_2}{k\Delta} \\ &\geq \frac{(n-b)(x-1)}{kx} = \frac{n(x-1) - b(x-1)}{kx} \end{aligned} \quad (4)$$

$$\begin{aligned} \frac{SOL}{OPT} &\geq \frac{\frac{b\Delta + (n-b)(\Delta - \lceil \frac{\Delta}{x} \rceil + 1)}{2}}{k\Delta} \\ &\geq \frac{bx + (n-b)(x-1)}{2kx} \geq \frac{n(x-1) + b}{2kx} \end{aligned} \quad (5)$$

Observe that the righthand side of (4) decreases with  $b$  while that of (5) increases; thus, the worst case occurs when righthand sides of them are equal, that is  $\frac{n(x-1) - b(x-1)}{kx} = \frac{n(x-1) + b}{2kx} \Leftrightarrow b = \frac{n(x-1)}{2x-1}$  and hence:

$$\frac{SOL}{OPT} \geq \frac{n(x-1) + \frac{n(x-1)}{2x-1}}{2kx} = \frac{n(x-1)}{k(2x-1)} \quad (6)$$

If  $b = k$ , then trivially holds that:

$$\frac{SOL}{OPT} \geq \frac{k \lceil \frac{\Delta}{x} \rceil}{k\Delta} \geq \frac{1}{x} \quad (7)$$

Note that (6) increases with  $x$  while (7) decreases; therefore, for the worst case we have  $\frac{n(x-1)}{k(2x-1)} = \frac{1}{x} \Leftrightarrow x = \frac{n+2k+\sqrt{4k^2+n^2}}{2n}$ . In all, it holds that:

$$\frac{SOL}{OPT} \geq \frac{2n}{n+2k+\sqrt{4k^2+n^2}} \quad (8)$$

If  $k < 0.55n$ , the ratio of (8) leads to:

$$\frac{SOL}{OPT} \geq \frac{2n}{n+2(0.55n)+\sqrt{4(0.55n)^2+n^2}} = \frac{2}{2.11+\sqrt{2.21}} = 0.55$$

On the other hand, the ratio provided in Proposition 6 that holds for any algorithm, for  $k > 0.55n$ , gives  $\frac{SOL}{OPT} \geq \frac{k}{n} \geq \frac{0.55n}{n} = 0.55$ .  $\square$

Let us note that, as it can be easily derived from (8), *when  $k = o(n)$  the competitive ratio of ALGORITHM mkvc-R is asymptotical to 1.*

### 3.2 Regular bipartite graphs

A better ratio can be achieved if we further restrict in regular bipartite graphs. A key-point of such improvement is that the maximum independent set can be found in polynomial time in bipartite graphs (see for example [6]). In what follows in this section, we consider that the number of vertices,  $n$ , is known a priori.

Our **ALGORITHM  $mkvc$ -B** initializes its solution with the first  $k$  released vertices. At this point, a maximum independent set  $B$ , of size  $b \leq k$ , in the graph induced by these  $k$  vertices is found. The vertices of this independent set will surely appear in the final solution. For the remaining  $k - b$  vertices we check if they cover at least  $\frac{\frac{n\Delta}{2} - b\Delta}{\lceil \frac{n-b}{k-b} \rceil}$  edges different from those covered by the independent set  $B$ . If yes, we return the solution consisting of the  $b$  vertices of the independent set and these  $k - b$  vertices. Otherwise, we wait for the next  $k - b$  vertices and we repeat the check. In **ALGORITHM  $mkvc$ -B**,  $G[A]$  denotes the subgraph of  $G$  induced by the vertex-subset  $A$ .

---

**ALGORITHM  $mkvc$ -B**

```

1:  $A = \{\text{the first } k \text{ released vertices}\};$ 
2: Find a maximum independent set  $B \subseteq A$  in  $G[A]$ ;  $b = |B|$ ;
3: for each released vertex  $v$  do
4:   if  $|A| = k$  then
5:     if  $m(A) \geq b\Delta + \frac{\frac{n\Delta}{2} - b\Delta}{\lceil \frac{n-b}{k-b} \rceil}$  then
6:       return  $A$ ;
7:     else
8:        $A = B$ ;
9:     else
10:       $A = A \cup \{v\}$ 
11: return  $A$ ;
```

---

**Theorem 2.** **ALGORITHM  $mkvc$ -B** achieves a 0.6075-competitive ratio.

*Proof.* Let us call *batch* the set of the  $k - b$  vertices of  $A \setminus B$  in Lines 5–10 of **ALGORITHM  $mkvc$ -B**.

The solution obtained by this algorithm contains a maximum independent set of size  $b$ . Since the input graph is bipartite, it holds that  $b \geq \frac{k}{2}$ .

The number of edges of the graph uncovered by the vertices of the maximum independent set is in total  $\frac{n\Delta}{2} - b\Delta$ . Any of these edges is covered by vertices belonging to at least one of the  $\lceil \frac{n-b}{k-b} \rceil$  batches. Hence, in average, each batch covers  $\frac{\frac{n\Delta}{2} - b\Delta}{\lceil \frac{n-b}{k-b} \rceil}$  of those edges; so there exists a batch that covers at least  $\frac{\frac{n\Delta}{2} - b\Delta}{\lceil \frac{n-b}{k-b} \rceil}$  of them. Therefore, the algorithm covers in total at least  $b\Delta + \frac{\frac{n\Delta}{2} - b\Delta}{\lceil \frac{n-b}{k-b} \rceil}$  edges.

Using Proposition 5, we get  $\frac{SOL}{OPT} \geq \frac{b\Delta + \frac{\frac{n\Delta}{2} - b\Delta}{\lceil \frac{n-b}{k-b} \rceil}}{k\Delta} = \frac{b + \frac{\frac{n}{2} - b}{\lceil \frac{n-b}{k-b} \rceil}}{k}$  and since this quantity increases with  $b$  it holds that:

$$\frac{SOL}{OPT} \geq \frac{\frac{k}{2} + \frac{\frac{n}{2} - \frac{k}{2}}{\lceil \frac{n - \frac{k}{2}}{k - \frac{k}{2}} \rceil}}{k} = \frac{k + \frac{n-k}{\lceil \frac{2n-k}{k} \rceil}}{2k} \quad (9)$$



If  $k \leq 0.6075n$ , then (9) leads to  $\frac{SOL}{OPT} \geq 0.6075$ . Otherwise, using Proposition 6 we get the same ratio and the theorem is concluded.  $\square$

Note that by (9), **ALGORITHM mkvc-B** achieves a competitive ratio asymptotical to  $\frac{3}{4}$  when  $k = o(n)$ .

### 3.3 Trees and chains

In this section we give algorithms that further improve the competitive ratios for the  $mkvc$  problem in trees and chains. Dealing with trees the following result holds.

**Proposition 7.** *The  $mkvc$  problem can be solved within  $(1 - \frac{k-1}{\Delta^*})$ -competitive ratio in trees, where  $\Delta^*$  is the sum of the  $k$  largest degrees in the tree. The ratio is tight.*

Note that, if the number of vertices of degree greater than 1 is  $r < k$  then our algorithm finds an optimum solution using just  $r$  vertices, since the edges that are adjacent to the leaves are covered by their other endpoints.

Furthermore, in the case where all the internal vertices of the tree have the same degree  $\Delta$ , the ratio provided by Proposition 7 becomes  $(1 - \frac{k-1}{k\Delta})$ . This ratio is better than the ratio proved for regular bipartite graphs in Theorem 2 for any  $\Delta \geq 3$ , but it is worse for  $\Delta = 2$ , i.e., in the case where the input graph is a chain.

An improvement for the  $mkvc$  problem in chains follows. The main idea of the algorithm is to partition the solution,  $A$ , into two disjoint parts, whose size is dynamically adjusted: the set  $B$  of vertices that contribute two edges in  $E(A)$  and the set  $C$  of vertices that contribute one edge in  $E(A)$ . Thus,  $A = B \cup C$  and  $B \cap C = \emptyset$ .

---

#### **ALGORITHM mkvc-C**

- 1:  $A = \emptyset$ ;  $B = \emptyset$ ;  $C = \emptyset$ ; In any step  $A \equiv B \cup C$ ;
  - 2: **for** each released vertex  $v$  **do**
  - 3:   **if**  $|B| \leq k$  **and**  $v$  adds two new edges to the solution **then**
  - 4:     **if**  $|A| = k$  **then**
  - 5:       Delete an arbitrary vertex from  $C$ ;
  - 6:        $B = B \cup \{v\}$ ;
  - 7:   **else if**  $|A| < k$  **and**  $v$  adds one new edge to the solution **then**
  - 8:      $C = C \cup \{v\}$ ;
  - 9:     **if** the inclusion of  $v$  in  $A$  has as a result three consecutive vertices to appear in  $A$  **then**
  - 10:       Move  $v$  from  $C$  to  $B$ ; Remove the middle vertex from  $A$ ;
  - 11: **return**  $A$ ;
- 

**Proposition 8.** *For the  $mkvc$  problem in chains, **ALGORITHM mkvc-C** returns the (offline) optimum, if  $k < \lceil \frac{n}{3} \rceil$  or  $k \geq \lceil \frac{2n}{3} \rceil$ , and achieves a 0.75-competitive ratio, if  $\lceil \frac{n}{3} \rceil \leq k < \lceil \frac{2n}{3} \rceil$ .*

## 4 Maximum $k$ -(set)-coverage

In this section we present **ALGORITHM  $mkc$**  for the online maximum  $k$ -(set)-coverage problem. It initializes by selecting the first  $k$  released sets. Then, considering that the current solution  $A_j$  covers  $m(A_j)$  elements, the algorithm replaces a set  $Q \in A_j$  by the new released set  $P$ , only if the number of elements covered is increased by at least  $\frac{m(A_j)}{k}$ . We prove that **ALGORITHM  $mkc$**  achieves competitive ratio strictly greater than  $\frac{1}{4}$  but that tends to  $\frac{1}{4}$  as  $k$  increases. Recall that the algorithm presented in [7] achieves also an  $\frac{1}{4}$ -competitive ratio. However, our analysis is tight and gives better results for moderately large values of  $k$ .

---

### **ALGORITHM $mkc$**

- 1:  $j = 1$ ;  $A_j = \{\text{the first } k \text{ released sets}\}$ ;
  - 2: **for** each released set  $P$  **do**
  - 3:   Find the set  $Q \in A_j$  that covers privately the smallest number of elements in  $A_j$ ;
  - 4:   **if**  $m(A_j \setminus \{Q\} \cup \{P\}) > m(A_j) + \frac{m(A_j)}{k}$  **then**
  - 5:      $j = j + 1$ ;  $A_j = A_{j-1} \setminus \{Q\} \cup \{P\}$ ;
- 

To analyze **ALGORITHM  $mkc$** , let  $A_z$  be the final solution obtained, i.e.,  $SOL = m(A_z)$ . Fix, also, an optimum solution  $A^*$ .

We consider the following two types of events that may happen during the execution of the algorithm upon arrival of a set  $P$ : (a)  $P \in A^*$  and **ALGORITHM  $mkc$**  does not select it, and (b)  $P \notin A^*$  and **ALGORITHM  $mkc$**  discards  $Q \in A^*$  in order to insert  $P$  into its current solution. Clearly, at most  $k$  such events may happen in total. However, not all events happen in a different current solution; let  $\ell \leq k$  be the number of the different current solutions when events happen. Let  $A_{j_i}$ ,  $1 \leq i \leq \ell$ ,  $1 \leq j_i \leq z$ , be the  $i$ -th of these current solutions, and  $k_i$ ,  $1 \leq i \leq \ell$ , be the number of events occurred in  $A_{j_i}$ .

We will now provide an upper bound to the value  $OPT = m(A^*)$  as function of the states  $A_{j_i}$ ,  $1 \leq i \leq \ell$ . Consider that the  $s$ -th,  $1 \leq s \leq k$ , event happens in  $j_i$ . Let  $P_s$  be the new set that arrives at  $j_i$  and  $Q_s \in A_{j_i}$  be the set that covers privately the smallest number of elements in  $A_{j_i}$ . Let, also,  $\tilde{Q}_s \subseteq Q_s$  be the set of private elements of  $Q_s$  in  $A_{j_i}$ .

If the event is of type (a) then  $P_s \in A^*$  is not selected and it covers a subset of the elements in  $E(A_{j_i} \setminus \{Q_s\})$  plus its private elements,  $\tilde{P}_s \subseteq P_s$ , in  $E(A_{j_i} \setminus \{Q_s\} \cup \{P_s\})$ . Note that it is  $m(\tilde{P}_s) \leq m(\tilde{Q}_s) + \frac{m(A_{j_i})}{k}$ , otherwise  $P_s$  should be selected by the algorithm. Moreover,  $m(\tilde{Q}_s) \leq \frac{m(A_{j_i})}{k}$ , since  $Q_s$  has the smallest private part in  $A_{j_i}$ , and hence  $m(\tilde{P}_s) \leq \frac{2m(A_{j_i})}{k}$ . If the event is of type (b) then  $Q_s \in A^*$  is removed and the elements covered by  $Q_s$  are a subset of  $E(A_{j_i})$ . In all, in the worst case we have:

$$E(A^*) \subseteq \bigcup_{i=1}^{\ell} E(A_{j_i}) \cup \bigcup_{s=1}^k \tilde{P}_s \subseteq E(A_{j_\ell}) \cup \bigcup_{i=2}^{\ell} [E(A_{j_{i-1}}) \setminus E(A_{j_i})] \cup \bigcup_{s=1}^k \tilde{P}_s$$

Therefore, for the value of the optimum solution  $A^*$  we have:

$$OPT \leq m(A_{j_\ell}) + \sum_{i=2}^{\ell} m(E(A_{j_{i-1}}) \setminus E(A_{j_i})) + \sum_{i=1}^{\ell} \left( k_i \frac{2m(A_{j_i})}{k} \right)$$

**Claim 1.**  $m(E(A_{j_{i-1}}) \setminus E(A_{j_i})) \leq \frac{|A_{j_{i-1}} \setminus A_{j_i}|}{k} m(A_{j_{i-1}})$ ,  $2 \leq i \leq \ell$ .

Using Claim 1 and since  $m(A_{j_\ell}) > m(A_{j_i})$ ,  $1 \leq i \leq \ell - 1$ , and  $\sum_{i=1}^{\ell} k_i = k$  we get:

$$OPT \leq m(A_{j_\ell}) + \sum_{i=2}^{\ell} \frac{|A_{j_{i-1}} \setminus A_{j_i}|}{k} m(A_{j_{i-1}}) + \sum_{i=1}^{\ell-1} \frac{2m(A_{j_i})}{k} + (k - \ell + 1) \frac{2m(A_{j_\ell})}{k}$$

By definition, it holds that  $j_\ell \leq z$  and hence  $m(A_{j_\ell}) \leq m(A_z) = SOL$ . Moreover, by **ALGORITHM mkc**,  $m(A_{j_\ell}) \geq \left(1 + \frac{1}{k}\right)^{j_\ell - j_i} m(A_{j_i})$ . Thus, we have:

$$\begin{aligned} \frac{SOL}{OPT} &\geq \frac{1}{1 + \frac{1}{k} \sum_{i=2}^{\ell} \frac{j_i - j_{i-1} + 2}{\left(1 + \frac{1}{k}\right)^{j_\ell - j_{i-1}}} + \frac{2(k - \ell + 1)}{k}} \\ &= \frac{1}{3 + \frac{1}{k} \sum_{i=2}^{\ell} \frac{j_i - j_{i-1} + 2}{\left(1 + \frac{1}{k}\right)^{j_\ell - j_{i-1}}} - \frac{2(\ell - 1)}{k}} \end{aligned} \quad (10)$$

**Claim 2.** For any  $\ell \geq 2$ , it holds that  $\sum_{i=2}^{\ell} \frac{j_i - j_{i-1} + 2}{\left(1 + \frac{1}{k}\right)^{j_\ell - j_{i-1}}} \leq \frac{g(\ell)}{\ln\left(1 + \frac{1}{k}\right)}$ , where  $g(\ell) = \frac{\left(1 + \frac{1}{k}\right)^2}{e} \cdot e^{g(\ell-1)}$  and  $g(2) = \frac{\left(1 + \frac{1}{k}\right)^2}{e}$ .

Using Claim 2 and (10), we get  $\frac{SOL}{OPT} \geq \frac{1}{3 + \frac{1}{k} \left[ \frac{g(\ell)}{\ln\left(1 + \frac{1}{k}\right)} - 2(\ell - 1) \right]}$ , where  $g(\ell) = \frac{\left(1 + \frac{1}{k}\right)^2}{e} \cdot e^{g(\ell-1)}$  and  $g(2) = \frac{\left(1 + \frac{1}{k}\right)^2}{e}$ . This quantity is minimized for an  $\ell = o(k)$ . The ratio  $r$  achieved by **ALGORITHM mkc** for different values of  $k$  is shown in Table 1.

$k$	2	3	5	10	30	50	100	300	500	1000
$r$	0.333	0.324	0.314	0.300	0.282	0.275	0.268	0.261	0.258	0.256

Table 1: Approximation ratio of **ALGORITHM mkc**

To see that the ratio achieved by **ALGORITHM mkc** is always greater than  $\frac{1}{4}$ , consider the following expression for the ratio, slightly less fine than (10):

$$\frac{SOL}{OPT} \geq \frac{1}{3 + \frac{1}{k} \sum_{i=2}^{\ell} \frac{j_i - j_{i-1}}{\left(1 + \frac{1}{k}\right)^{j_\ell - j_{i-1}}} + \frac{1}{k} \sum_{i=2}^{\ell} \left( \frac{2}{\left(1 + \frac{1}{k}\right)^{j_\ell - j_{i-1}}} - 2 \right)}$$

Note first that if  $\ell = 1$  then both sums on the denominator are zero and hence we have a  $\frac{1}{3}$ -competitive ratio. If  $\ell \geq 2$  we have the following analysis. For the first sum, by a similar analysis as in Claim 2 we can prove that  $\sum_{i=2}^{\ell} \frac{j_i - j_{i-1}}{\left(1 + \frac{1}{k}\right)^{j_\ell - j_{i-1}}} \leq \frac{g(\ell)}{\ln\left(1 + \frac{1}{k}\right)}$ , where  $g(\ell) = \frac{1}{e} \cdot e^{g(\ell-1)}$  and  $g(2) = \frac{1}{e}$ . It is easy to see by a simple

induction that  $g(\ell) \leq 1$  for any  $\ell \geq 2$  and hence  $\sum_{i=2}^{\ell} \frac{j_i - j_{i-1}}{\left(1 + \frac{1}{k}\right)^{j_\ell - j_{i-1}}} \leq \frac{1}{\ln\left(1 + \frac{1}{k}\right)} \leq k$ . For the second sum, we have:

$$\sum_{i=2}^{\ell} \left( \frac{2}{\left(1 + \frac{1}{k}\right)^{j_\ell - j_{i-1}}} - 2 \right) \leq \sum_{i=2}^{\ell} \left( \frac{2}{\left(1 + \frac{1}{k}\right)} - 2 \right) = \frac{2k}{k+1} - 2 = -\frac{2}{k+1}$$

Therefore, using these bounds to the ratio we get:

$$\frac{SOL}{OPT} \geq \frac{1}{4 - \frac{2}{k(k+1)}} = \frac{1}{4} + \frac{1}{4} \frac{1}{2k(k+1)-1}$$

It is hopefully clear from the previous discussion, that the analysis of **ALGORITHM M $k$ C** works also for the **WEIGHTED M $k$ C** problem, up to the assumption that  $m(\cdot)$  in **ALGORITHM M $k$ C** is the total weight of the elements and not their number.

We conclude this section by providing a tight example for the ratio achieved by **ALGORITHM M $k$ C**. The idea of the example is strongly based upon the proof given above, which indicates the “critical” values of  $\ell$  and  $j_i$ ,  $1 \leq i \leq \ell$ . For simplicity, we will consider the case where  $k = 3$ , but it is easy to extend our example for any  $k$ , by appropriately choosing the values of  $\ell$  and  $j_i$ .

For  $k = 3$ , one can see that the ratio of **ALGORITHM M $k$ C** is minimized when  $\ell = 2$  and  $j_2 - j_1 = 1$ . Hence, consider the scenario shown in Figure 1. Let

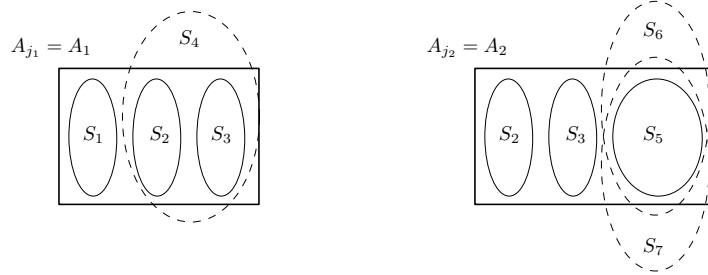


Figure 1: A tight example for **ALGORITHM M $k$ C** when  $k = 3$ .

$A_1 = \{S_1, S_2, S_3\}$  be the solution after the first three sets have been released. These sets are disjoint and each one covers  $c$  elements. Next, the set  $S_4$  appears, which covers  $2c - \epsilon$  new elements plus all elements in  $S_2$  and  $S_3$ . The algorithm does not select  $S_4$ , since it may choose  $S_1$  as a candidate for swapping; in this case the new solution would cover  $m(\{S_2, S_3, S_4\}) = 4c - \epsilon$  elements which is smaller than  $m(A_1) + \frac{m(A_1)}{k} = 4c$ . Then, the set  $S_5$  is released, that is disjoint with the previous sets and covers  $2c$  elements. Thus, the algorithm replaces  $S_1$  by  $S_5$ , and the new solution is  $A_2 = \{S_2, S_3, S_5\}$ . Finally,  $S_6$  and  $S_7$  are released, each covering the elements in  $S_5$  plus  $\frac{8c}{3} - \epsilon$  new elements. **ALGORITHM M $k$ C** does not select none of them, since they do not satisfy the algorithm’s criterion.

So, the final solution,  $A_2$ , covers  $m(S_2) + m(S_3) + m(S_5) = 4c$  elements. The optimal solution consists of sets  $S_4$ ,  $S_6$  and  $S_7$  and covers  $OPT = (4c - \epsilon) +$

$(2c + \frac{8c}{3} - \epsilon) + (\frac{8c}{3} - \epsilon) = \frac{34c}{3} - \epsilon$  elements. Therefore, the ratio achieved by **ALGORITHM  $m\bar{k}c$**  is  $\frac{SOL}{OPT} = \frac{4c}{\frac{34c}{3} - \epsilon} \simeq \frac{12}{34} = 0.353$ .

Note that the gap between this ratio and the ratio 0.324 given in Table 1 is due to the fact that the elements of  $S_1$  do not appear to the optimal solution. Indeed, if  $S_1$  was included to the optimal, then  $OPT = \frac{37c}{3} - \epsilon$  and  $\frac{SOL}{OPT} = \frac{4c}{\frac{37c}{3} - \epsilon} \simeq \frac{12}{37} = 0.324$ . This gap decreases as  $k \rightarrow \infty$ .

## 5 Conclusions

There exist several interesting questions arising from the results presented in this paper. The first of them is to improve the easy  $\frac{1}{2}$ -competitive ratio for  $m\bar{k}VC$  in general graphs and the (less easy) worst-case  $\frac{1}{4}$ -competitive ratio in set systems. Another open question is to provide tighter upper bounds for the on-line model handled in regular graphs. We still do not see how one can improve the analysis of **ALGORITHM  $m\bar{k}c$**  in the case of equal cardinalities, or how to tighten the upper bound of Proposition 3 in Section 2, in order to match (or to get closer to) the competitive ratio of **ALGORITHM  $m\bar{k}c$** . Let us note that an algorithm in the spirit of **ALGORITHM  $m\bar{k}VC$ -R** of Section 3.1 for the case of equal-cardinality sets, only achieves ratio  $\frac{1}{\sqrt{k}}$ .

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## A Appendix

### A.1 Proof of Proposition 1

**Proposition 1.** *Any deterministic online algorithm that does not allow swaps cannot achieve a competitive ratio better than*

- $O\left(\frac{1}{(n-1)^{1/(k+1)}}\right)$  for the *MkVC* problem
- $O\left(\frac{1}{m^{1/(k+1)}}\right)$  for the *MkC* problem.

*Proof.* For the *MkVC* problem, let  $k \ll n$  and consider the following scenario. In step  $i$ ,  $1 \leq i \leq k$ , the central vertex  $v_i$  of a star with  $d(v_i) = (n-1)^{i/(k+1)}$  is released. If the algorithm rejects  $v_i$ , then the remaining  $\sum_{j=1}^i (n-1)^{j/(k+1)}$  vertices of the  $i$  stars plus  $n - i - \sum_{j=1}^i (n-1)^{j/(k+1)}$  singleton vertices are released. If the algorithm selects  $v_i$ , then vertex  $v_{i+1}$ , with  $d(v_{i+1}) = (n-1)^{(i+1)/(k+1)}$  is released. If after the  $k$ -th vertex the algorithm has selected all the  $k$  released vertices then a new vertex  $v_{k+1}$  with degree  $d(v_{k+1}) = n-1$  is released; finally the remaining vertices of the stars and  $n - k - \sum_{j=1}^k (n-1)^{j/(k+1)}$  singleton vertices are released.

If after the step  $i$  the algorithm has rejected  $v_i$ , then only vertices of degree at most one are released. Hence, the algorithm covers  $k - (i-1) + \sum_{j=1}^{i-1} (n-1)^{j/(k+1)}$  edges, while the optimum solution covers  $k - i + \sum_{j=1}^i (n-1)^{j/(k+1)}$  edges. Thus:

$$\begin{aligned} \frac{SOL}{OPT} &= \frac{k - (i-1) + \sum_{j=1}^{i-1} (n-1)^{j/(k+1)}}{k - i + \sum_{j=1}^i (n-1)^{j/(k+1)}} \\ &= \frac{k - (i-1) + \frac{(n-1)^{i/(k+1)} - (n-1)^{1/(k+1)}}{(n-1)^{1/(k+1)} - 1}}{k - i + \frac{(n-1)^{(i+1)/(k+1)} - (n-1)^{1/(k+1)}}{(n-1)^{1/(k+1)} - 1}} = O\left(\frac{1}{(n-1)^{1/(k+1)}}\right) \end{aligned}$$

If the algorithm has selected all the  $k$  first released vertices, then it covers exactly  $\sum_{j=1}^k (n-1)^{j/(k+1)}$  elements, while the optimum solution (that includes  $v_{k+1}$  which is never selected by the online algorithm) covers  $n-1$  elements. Hence:

$$\begin{aligned} \frac{SOL}{OPT} &= \frac{\sum_{j=1}^k (n-1)^{j/(k+1)}}{n-1} = \frac{(n-1)^{(k+1)/(k+1)} - (n-1)^{1/(k+1)}}{(n-1)^{1/(k+1)} - 1} \\ &= O\left(\frac{1}{(n-1)^{1/(k+1)}}\right) \end{aligned}$$

that concludes the proof.

For the *MkC* problem the proof is similar. In this case, in phase  $i$ , if all previously released sets are selected by the algorithm then a set of cardinality  $m^{i/(k+1)}$  is released.  $\square$

### A.2 Proof of Proposition 3

**Proposition 3.** *Any deterministic online algorithm cannot achieve a competitive ratio better than  $\frac{k+2\sqrt{k+1}}{2k+2\sqrt{k+1}} \simeq \frac{1}{2}$  for the *MkC* problem even in the case where*

all sets have the same cardinality.

*Proof.* A  $r$ -sunflower is a set system of regular sets of size  $\Delta$  with a common intersection of size  $r$ ; the sets of a sunflower are called *petals*.

Consider the following scenario. The adversary starts by sending  $\frac{\Delta(p-1)}{p}$ -sunflower petals where  $\frac{\Delta}{p} \in \mathbb{N}$ , while the algorithm keeps  $k'$  of them; it continues so until the first time  $\tau$  where there are  $k - \lfloor \frac{k'}{p} \rfloor$  rejected sets. Notice that this will be always the case for some  $\tau \leq \lfloor \frac{k(2p-1)}{p} \rfloor$ .

Then the adversary starts sending disjoint sets, each one matching private parts of  $p$  petals in the solution, until the maximum number of private parts have been matched.

The solution of the algorithm will cover at most  $\frac{(k'+p-1)}{p}\Delta$  elements, while the optimum will cover at least  $\lfloor \frac{k'}{p} \rfloor \Delta$  elements by the matching sets plus  $\lfloor \frac{k - \frac{k'}{p} + p - 1}{p} \rfloor \Delta$  elements by rejected petals. Thus, the ratio is bounded above by  $\frac{k'+p-1}{k+(p-1)\frac{k'}{p}+p-1}$  where  $0 \leq k' \leq k$ , which is less than or equal to the simplified expression  $\frac{pk+p(p-1)}{(2p-1)k+p(p-1)}$ . This expression is minimized when  $p = \sqrt{k} + 1$ , that is  $\frac{(\sqrt{k}+1)k+(\sqrt{k}+1)\sqrt{k}}{(2(\sqrt{k}+1)-1)k+(\sqrt{k}+1)\sqrt{k}} = \frac{k+2\sqrt{k}+1}{2k+2\sqrt{k}+1}$ , which for  $k$  large enough tends asymptotically to  $\frac{1}{2}$ .  $\square$

### A.3 Proof of Proposition 6

**Proposition 6.** *Any deterministic online algorithm achieves a  $\frac{k}{n}$ -competitive ratio for the  $MkVC$  problem on regular graphs.*

*Proof.* An optimum solution covers at most all the edges of the graph (recall that  $|E| = m$ ), that is  $OPT \leq m = \frac{n\Delta}{2}$ . On the other hand, any solution covers  $k\Delta$  edges, some of them possibly twice, i.e., at least  $\frac{k\Delta}{2}$  edges, that is  $SOL \geq \frac{k\Delta}{2}$ . We so get  $\frac{SOL}{OPT} \geq \frac{k}{n}$ .  $\square$

### A.4 Proof of Proposition 7

**Proposition 7.** *The  $MkVC$  problem can be solved within  $(1 - \frac{k-1}{\Delta^*})$ -competitive ratio in trees, where  $\Delta^*$  is the sum of the  $k$  largest degrees in the tree. The ratio is tight.*

*Proof.* An upper bound for the optimum solution is  $OPT \leq \Delta^*$ , that is the case where  $k$  non-adjacent vertices of the largest degree are selected.

Consider the algorithm that selects the  $k$  vertices of the largest degrees. These  $k$  vertices cover  $\Delta^*$  edges, some of them possibly twice. It is easy to see that the number of such edges is maximized when the subgraph induced by the  $k$  selected vertices is connected. In this case, there are  $k - 1$  edges covered twice.

Hence, the total number of covered edges is  $\Delta^* - (k - 1)$ , while at most  $\Delta^*$  edges can be covered by any solution.  $\square$

## A.5 Proof of Proposition 8

**Proposition 8.** *For the  $\text{MkVC}$  problem in chains,  $\text{ALGORITHM MkVC-C}$  returns the (offline) optimum, if  $k < \lceil \frac{n}{3} \rceil$  or  $k \geq \lceil \frac{2n}{3} \rceil$ , and achieves a 0.75-competitive ratio, if  $\lceil \frac{n}{3} \rceil \leq k < \lceil \frac{2n}{3} \rceil$ .*

*Proof.* Since there do not exist three consecutive vertices in the solution obtained by the algorithm, each vertex in  $B$  has at most one adjacent vertex in  $C$ , and hence  $|B| \geq |C|$ .

If  $k < \lceil \frac{n}{3} \rceil$  then assume, for contradiction, that in the final solution it holds that  $C \neq \emptyset$  and let  $v \in C$ . Then,  $SOL = 2|B| + |C| \leq 2(k - 1) + 1 = 2k - 1 < 2 \lceil \frac{n}{3} \rceil - 1$ . Thus, the non-covered edges of the input graph are at least  $n - 2 \lceil \frac{n}{3} \rceil + 2$ . Since the number of connected components in the solution of the algorithm is at most  $k - 1 < \lceil \frac{n}{3} \rceil - 1$ , there are two adjacent edges not covered by the algorithm; let  $u$  the common vertex of these edges. But the algorithm in Lines 3–6 should have removed  $v$  and add  $u$  in  $A$ , a contradiction. Thus,  $C = \emptyset$  which means that all the  $k$  vertices of the solution cover privately exactly two edges, as in the optimum.

If  $k \geq \lceil \frac{2n}{3} \rceil$ , then the algorithm returns a solution containing all the vertices of the graph. Indeed, since  $|B| \geq |C|$  we have  $SOL = 2|B| + |C| \geq 2 \lceil \frac{n}{3} \rceil + \lfloor \frac{n}{3} \rfloor = n$ . Hence, the optimum solution is obtained by the algorithm.

If  $\lceil \frac{n}{3} \rceil \leq k < \lceil \frac{2n}{3} \rceil$ , then for the solution created by the algorithm we have  $SOL = 2|B| + |C|$ , while  $OPT \leq 2k = 2|B| + 2|C|$ . Since  $|B| \geq |C|$ , we get:

$$\frac{SOL}{OPT} \geq \frac{2|B| + |C|}{2|B| + 2|C|} \geq \frac{2|B| + |B|}{2|B| + 2|B|} = \frac{3}{4} = 0.75$$

that completes the proof.  $\square$

## A.6 Proof of Claim 1

**Claim 1.**  $m(E(A_{j_{i-1}}) \setminus E(A_{j_i})) \leq \frac{|A_{j_{i-1}} \setminus A_{j_i}|}{k} m(A_{j_{i-1}})$ ,  $2 \leq i \leq \ell$ .

*Proof.* Let  $Q_r$ ,  $1 \leq r \leq |A_{j_{i-1}} \setminus A_{j_i}|$ , be the  $r$ -th set that is removed from  $A_{j_{i-1}}$  between the events  $i - 1$  and  $i$ , considering only the sets that exist in  $A_{j_{i-1}}$ . Let, also,  $\tilde{Q}_r$  be the private part of  $Q_r$  just before it is removed. We will show that, for any  $p$ ,  $1 \leq p \leq |A_{j_{i-1}} \setminus A_{j_i}|$ , it holds that  $\sum_{r=1}^p q_r \leq \frac{p}{k} m(A_{j_{i-1}})$ , and thus:

$$m(E(A_{j_{i-1}}) \setminus E(A_{j_i})) = \sum_{r=1}^{|A_{j_{i-1}} \setminus A_{j_i}|} q_r \leq \frac{|A_{j_{i-1}} \setminus A_{j_i}|}{k} m(A_{j_{i-1}})$$

Assume for a contradiction that for the first time after the removal of the set  $Q_p$  it holds that  $\sum_{r=1}^p q_r > \frac{p}{k} m(A_{j_{i-1}})$ , hence,  $\sum_{r=1}^{p-1} q_r \leq \frac{p-1}{k} m(A_{j_{i-1}})$ . Clearly,



$q_p > \frac{m(A_{j_{i-1}})}{k}$ . Moreover, following **ALGORITHM mkC**  $Q_p$  has the smallest private part between the sets belonging in the solution when  $Q_p$  is selected to be removed. Thus, the  $k-p$  sets of  $A_{j_{i-1}}$  which are still in the solution have private parts of size greater than  $(k-p)\frac{m(A_{j_{i-1}})}{k}$  in total. Consequently:

$$\begin{aligned} m(A_{j_{i-1}}) &> \sum_{r=1}^p q_r + (k-p)\frac{m(A_{j_{i-1}})}{k} \\ &> \frac{p}{k}m(A_{j_{i-1}}) + (k-p)\frac{m(A_{j_{i-1}})}{k} = m(A_{j_{i-1}}) \end{aligned}$$

a contradiction. Therefore, there is no  $p$  such that  $\sum_{r=1}^p q_r > \frac{p}{k}m(A_{j_{i-1}})$ , and the claim is proved.  $\square$

## A.7 Proof of Claim 2

**Claim 2.** For any  $\ell \geq 2$ , it holds that  $\sum_{i=2}^{\ell} \frac{j_i - j_{i-1} + 2}{(1 + \frac{1}{k})^{j_i - j_{i-1}}} \leq \frac{g(\ell)}{\ln(1 + \frac{1}{k})}$ , where  $g(\ell) = \frac{(1 + \frac{1}{k})^2}{e} \cdot e^{g(\ell-1)}$  and  $g(2) = \frac{(1 + \frac{1}{k})^2}{e}$ .

*Proof.* Set  $d_i = j_i - j_{i-1}$ . Consider the function  $f_{\ell}(d) = \sum_{i=2}^{\ell} \frac{d_i + 2}{(1 + \frac{1}{k})^{\sum_{j=i}^{\ell} d_j}}$ . We will prove the claim by induction to  $\ell$ .

For  $\ell = 2$  we have  $f_2(d) = \sum_{i=2}^2 \frac{d_i + 2}{(1 + \frac{1}{k})^{\sum_{j=i}^2 d_j}} = \frac{d_2 + 2}{(1 + \frac{1}{k})^{d_2}}$ , where  $\frac{\partial f_2(d)}{\partial d_2} = \frac{1 - (d_2 + 2)\ln(1 + \frac{1}{k})}{(1 + \frac{1}{k})^{d_2}}$ . The global maximum is attained for  $d_2 + 2 = \frac{1}{\ln(1 + \frac{1}{k})}$ . Thus:  
 $f_2(d) \leq \frac{1}{\ln(1 + \frac{1}{k})} \cdot \frac{1}{(1 + \frac{1}{k})^{\frac{1}{\ln(1 + \frac{1}{k})} - 2}} = \frac{1}{\ln(1 + \frac{1}{k})} \cdot \frac{(1 + \frac{1}{k})^2}{e}$ .

Assume that the statement is true for  $\ell - 1$ .  
For  $\ell$ , we have:

$$\begin{aligned} f_{\ell}(d) &= \sum_{i=2}^{\ell} \frac{d_i + 2}{(1 + \frac{1}{k})^{\sum_{j=i}^{\ell} d_j}} = \frac{d_{\ell} + 2}{(1 + \frac{1}{k})^{d_{\ell}}} + \sum_{i=2}^{\ell-1} \frac{d_i + 2}{(1 + \frac{1}{k})^{\sum_{j=i}^{\ell} d_j}} \\ &= \frac{d_{\ell} + 2}{(1 + \frac{1}{k})^{d_{\ell}}} + \frac{1}{(1 + \frac{1}{k})^{d_{\ell}}} \cdot f_{\ell-1}(d) \end{aligned}$$

where  $\frac{\partial f_{\ell}(d)}{\partial d_{\ell}} = \frac{1 - (d_{\ell} + 2 + f_{\ell-1}(d))\ln(1 + \frac{1}{k})}{(1 + \frac{1}{k})^{d_{\ell}}}$ . The global maximum is attained for

$d_\ell + 2 + f_{\ell-1}(d) = \frac{1}{\ln(1+\frac{1}{k})}$ . Thus:

$$\begin{aligned} f_\ell(d) &\leq \frac{1}{\ln(1+\frac{1}{k})} \cdot \frac{1}{(1+\frac{1}{k})^{\frac{1}{\ln(1+\frac{1}{k})} - 2 - f_{\ell-1}(d)}} \\ &\leq \frac{1}{\ln(1+\frac{1}{k})} \cdot \frac{(1+\frac{1}{k})^2}{e} \cdot \left(1+\frac{1}{k}\right)^{\frac{g(\ell-1)}{\ln(1+\frac{1}{k})}} \\ &= \frac{1}{\ln(1+\frac{1}{k})} \cdot \frac{(1+\frac{1}{k})^2}{e} \cdot e^{g(\ell-1)} \end{aligned}$$

and the claim follows.  $\square$