Mixed covering of trees and the augmentation problem with odd diameter constraints

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Abstract. In this paper, we present a polynomial time algorithm for solving the problem of partial covering of trees with n_1 balls of radius R_1 and n_2 balls of radius R_2 $(R_1 < R_2)$ so as to maximize the total number of covered vertices. The solutions provided by this algorithm in the particular case $R_1 = R - 1, R_2 = R$ can be used to obtain for any integer $\delta > 0$ a factor $(2+\frac{1}{\delta})$ approximation algorithm for solving the following augmentation problem with odd diameter constraints $D = 2R + 1$: given a tree T, add a minimum number of new edges such that the augmented graph has diameter $\leq D$. The previous approximation algorithm of Ishii, Yamamoto, and Nagamochi (2003) has factor 8.

Key Words. Partial covering, Diameter, Augmentation problem, Dynamical programming, Approximation algorithms.

1 Introduction

In this paper, we present a polynomial time algorithm for solving the following covering problem on trees:

Problem PARTIAL MIXED COVERING: *Given a tree $T = (V, E)$ with n vertices, the non-negative integers R_1, R_2 ($R_1 < R_2$) and n_1, n_2, locate n_1 balls of radius R_1 and n_2 balls of radius R_2 so as to maximize the total number of covered vertices.*

This problem generalizes the MAXIMUM COVERAGE problem investigated by Megiddo, Zemel, and Hakimi [14], in which, given a tree T and the integers R_0 and n_0, one wish to locate n_0 balls of radius R_0 so as to maximize the total number of covered vertices. Unlike the R-DOMINATING problem on trees (asking for covering of a tree with a minimum number of balls of radius R), which is easily solvable in linear time, the existence of polynomial time algorithms for PARTIAL MIXED COVERING and MAXIMUM COVERAGE problems is nontrivial. The difficulty resides in the fact that we have to decide which vertices should be covered, which vertices should be chosen as centers, and balls of which radius should be located at those centers.

In [14], the initial motivation for studying the MAXIMUM COVERAGE problem came from the problem of locating a given number of facilities in a transportation network to cover a maximum number of customers. The more general setting of the PARTIAL MIXED COVERING problem allows to model situations in which two kinds of facilities are available, the difference between them being their range, i.e. the maximum distance between a facility and a customer supplied by it. Also, it turns out that the MAXIMUM COVERAGE and the PARTIAL MIXED COVERING problems are (polynomially solvable) special cases of the (NP-hard) general PARTIAL COVERING problem introduced by Kearns [12] and recently revisited in [9]. In the present paper, we provide yet another motivation for studying PARTIAL MIXED COVERING by deriving from it an approximation algorithm for the augmentation problem with diameter constraints which we formulate below. Notice also that several related problems can be reduced to PARTIAL MIXED COVERING. For example, running the algorithm for PARTIAL MIXED COVERING for all feasible pairs (n_1', n_2'), we obtain a polynomial time algorithm for the following problem:

Problem MIXED COVERING: *Given a tree $T = (V, E)$ with n vertices, a function f of two non-negative integer variables, the non-negative integers R_1, R_2 ($R_1 < R_2$) and n_1, n_2, find a covering (if it exists) of T with $n_1' \leq n_1$ balls of radius R_1 and $n_2' \leq n_2$ balls of radius R_2 minimizing the function $f(n_1', n_2')$.*

(If $f(n_1', n_2') = n_1'$, we obtain the problem of covering T with n_2 balls of radius R_2 and a minimum number of balls of radius R_1. In particular, if $R_1 = 0$, we get yet another formulation of the MAXIMUM COVERAGE problem.) The MIXED COVERING problem was first formulated in [3] in connection with the following augmentation problem:

Problem ADC (AUGMENTATION under DIAMETER CONSTRAINTS): *Given a graph $G = (V, E)$ with n vertices and a positive integer D, add a minimum number* OPT *of new edges E' such that the augmented graph $G' = (V, E \cup E')$ has diameter at most D.*

Due to its practical importance for improving the reliability of existing communication networks, the AUGMENTATION under DIAMETER CONSTRAINTS problem has received much attention in the literature [1, 3, 4, 5, 6, 7, 11, 13, 15]. In particular, it was shown to be NP-hard for any $D \geq 2$ and at least as difficult to approximate as SET COVER [3, 13, 15]. However, the complexity status of this problem is unknown if the input graph G is a tree. In case of paths, OPT is determined up to an additive constant error term. Namely, in this case, Chung and Garey [5] established that $(n - D - 1)/(D + 1) \leq \text{OPT} \leq (n - D + 2)/(D - 2)$, and Alon, Gyárfás and Ruszinkó [1] refined this bound by establishing that the values of OPT for the n-cycle (i.e., a path plus one additional edge) satisfy $\lfloor n/(D - 1)\rfloor - 7 \leq \text{OPT} \leq \lfloor n/(D - 1)\rfloor$ for even D and $\lfloor n/(D - 2)\rfloor - 146 \leq \text{OPT} \leq \lfloor n/(D - 2)\rfloor$ for odd D. Other lower and upper bounds for more general classes of graphs have been considered in [1, 7]; see also the survey of Chung [4] which contains further references and related problems.

For the problem ADC on trees, Chepoi and Vaxès [3] presented a factor 2 approximation algorithm for even $D = 2R$ and Ishii, Yamamoto, and Nagamochi [11] presented a factor 8 approximation algorithm for odd $D = 2R + 1$. In [3] it was conjectured that the optimal solutions provided by MIXED COVERING may be used to derive approximate feasible solutions for the problem ADC with odd D. In this paper, we prove that indeed any mixed covering of the input tree T with n_1 balls of radius $R-1$ and n_2 balls of radius R minimizing the function $f(n_1, n_2) = n_1 + \frac{n_2(n_2-1)}{2}$ can be transformed into a feasible solution for the problem ADC with $D = 2R + 1$ containing at most $(2 + \frac{1}{\delta})\text{OPT} + O(\delta^5)$ added edges for any integer $\delta > 0$, thus asymptotically matching the approximation ratio for even D. This augmentation (using at most $n_1 + \frac{n_2(n_2-1)}{2}$ new edges) is obtained by drawing an edge between any pair of centers of n_2 balls of radius R and between the center of any of n_1 balls of radius $R - 1$ and the center of some ball of radius R. Notice that the performance guarantees of all mentioned algorithms for trees should be much better, however the bottleneck in analyzing them is the difficulty of establishing better lower bounds for the minimum number of added edges; for example, the proof of the above mentioned lower bound for paths [5] is already quite involved.

The paper is organized as follows. In Section 2 we present a few necessary definitions and notations. Section 3 describes a dynamic programming algorithm for solving the PARTIAL MIXED COVERING problem. It also discusses the complexity and the correctness of the algorithm and presents some further problems which are solvable by a similar approach. Finally, in Section 4 we establish and analyze a factor $2 + \frac{1}{\delta}$ approximation algorithm for the problem ADC with odd diameter constraints.

2 Preliminaries

For a graph $G = (V, E)$, the *length* of a path between two vertices is the number of edges in this path. The *distance* $d_G(u, v)$ between two vertices u, v of G is the length of the shortest path between these vertices. The *diameter* of G is the largest distance between two vertices of G. For an integer $k \geq 0$ and a vertex $v \in V$, let $B_k(v) = \{x \in V : d_G(v, x) \leq k\}$ denote the *ball* of radius k centered at v. Set also $B_{-1}(v) = \emptyset$. The *relative radius* $rr(x, B_k(v))$ of a ball $B_k(v)$ in a vertex x of a tree $T = (V, E)$ equals $k - d_T(v, x)$. We say that a ball $B_k(v)$ is *located* in a subtree T' of T if $v \in T'$. For two vertices x, y of a tree T denote by $P(x, y)$ the unique path of T between x and y. For a subset $Q \subset V$ denote by $T(Q)$ the least subtree of T containing Q. For a vertex y in a rooted tree T with root u, any vertex $x \neq y$ on the path $P(u, y)$ is called an *ancestor* of y. If x is an ancestor of y, then y is a *descendant* of x. Denote by T_x the subtree of T rooted at the vertex x and consisting of x and all of its descendants.

Define the *kth power* of a tree $T = (V, E)$ as the graph T^k having V as vertex-set and two vertices x, y are adjacent in T^k if and only if $d_T(x, y) \leq k$. For a subset of vertices Q, denote by $T^k(Q)$ the subgraph of T^k induced by Q. It is well known [2] that the kth power T^k of a tree is a chordal graph (whence $T^k(Q)$ are chordal graphs for all $Q \subseteq V$). Therefore T^k and $T^k(Q)$ are perfect graphs for all k and Q (recall that a graph G is *perfect* [2] if the minimum number of cliques necessary to cover G equals the size of the largest stable set of G). Notice that $Y \subseteq V$ is a stable set of T^k if and only if $d_T(x, y) > k$ for any $x, y \in Y$. On the other hand, a clique C of T^k consists of vertices with pairwise distances (in T) at most k. The least subtree $T(C)$ containing the set C has diameter at most k, thus its radius is either at most R if $k = 2R$ or at most $R + 1$ if $k = 2R + 1$. In the first case, $T(C)$ can be covered by a ball of radius R. In the second case, $T(C)$ can be covered by an *edge-ball* of radius R, i.e., by two balls of radius R centered at adjacent vertices of $T(C)$. As a consequence, a covering of T^{2R} or of T^{2R+1} with a minimum number of cliques corresponds to a covering of T with a minimum number of balls of radius R or of edge-balls of radius R, and, due to the perfectness of the graphs T^{2R} and T^{2R+1}, this equals the size of a largest stable set of these graphs.

A polynomial time algorithm is called an *α-factor approximation* algorithm for a minimization problem Π if for each instance I of Π, it returns a solution whose value is at most α times the optimal value $\mathrm{OPT}_\Pi(I)$ of Π on I plus a constant not depending of I; see [16].

3 Mixed covering of trees

In this section, we describe a dynamic programming algorithm for solving the PARTIAL MIXED COVERING problem on trees. Our algorithm follows the main lines of the algorithm from [14] and works in general in the following way. Root the tree T at an arbitrary vertex u. The algorithm proceeds the tree T in a upward manner, from leaves to the root, by solving larger and larger subproblems of the following type. Given the current vertex s, and the integers $0 \leq n_1' \leq n_1, 0 \leq n_2' \leq n_2$, the algorithm finds the *maximal* number of covered vertices of T_s in a partial covering using n_1' balls of radius R_1 and n_2' balls of radius R_2 located in T_s. However, the algorithm must take care of two things: (i) some ball which will be located outside T_s at some later stage and whose radius and center are yet unknown may have an impact on the covering of T_s, and (ii) we have to consider the interaction between the subtrees rooted at the neighbors of s, because some vertices of one or several such subtrees may be covered by a ball located in another subtree. To overcome these difficulties, we introduce two additional parameters r and a which take integer values in the ranges $[-1, R_2 - 1]$ and $[0, R_2]$, respectively. For fixed values of r and a, the algorithm returns the maximal number of covered vertices of T_s in a partial covering using n_1' balls of radius R_1 and n_2' balls of radius R_2 located in T_s (permanent balls), given that one additional (temporary) ball of radius r is located at s and that the relative radius of at least one of the permanent balls located in T_s is at least a. This requires the solution of a resource allocation problem, which optimally distributes the balls of radius R_1 and the balls of radius R_2 among the subtrees rooted at the neighbors of s in T_s, using for this the optimal solutions of the previously solved subproblems at each of the sons of s. Following [14], we introduce the following functions EXT and INT:

1. $\mathrm{EXT}(T_s; n_1', n_2'; r)$ is equal to the maximum number of vertices of T_s which can be covered by n_1' balls of radius R_1 and n_2' balls of radius R_2 located in T_s, given that there is one additional ball of radius r centered at s. The algorithm computes $\mathrm{EXT}(T_s; n_1', n_2'; r)$ for all $r \in \{-1, 0, \ldots, R_2 - 1\}$, $n_1' \in \{0, \ldots, n_1\}$, $n_2' \in \{0, \ldots, n_2\}$, and all rooted subtrees $T_s, s \in V$, of T.

2. $\mathrm{INT}(T_s; n_1', n_2'; a)$ is equal to the maximum number of vertices of T_s which can be covered by n_1' balls of radius R_1 and n_2' balls of radius R_2 located in T_s, given that the relative radius $rr(s, B)$ in s of one of those balls B is at least a. The algorithm computes $\mathrm{INT}(T_s; n_1', n_2'; a)$ for all $a \in \{0, \ldots, R_2\}$, $n_1' \in \{0, \ldots, n_1\}$, $n_2' \in \{0, \ldots, n_2\}$, and all rooted subtrees $T_s, s \in V$, of T.

Let $s_1, ..., s_l$ be the sons of the current vertex s. To evaluate the functions INT and EXT on the subtree T_s, we use the values of these two functions on the subtrees $T_{s_1}, \ldots, T_{s_l}$.

If a temporary ball is located at s, then the algorithm computes the optimal distribution of remaining balls in the subtrees $T_{s_1}, \ldots, T_{s_l}$, taking into account the ball centered at s. If no ball is located at s, the algorithm distributes the balls to $T_{s_1}, \ldots, T_{s_l}$ so that to maximize the number of covered vertices. In this case, certain vertices of some subtree T_{s_i} can be covered by a ball located outside T_{s_i} and, vice versa, a ball centered at a vertex of T_{s_i} may cover a vertex outside this subtree. The first case is settled by the function EXT which locates a ball of radius r at s, the second case is settled by the function INT which forces the location in T_{s_i} of a ball which covers all vertices of T_s at distance at most a from s. Finally, in order to optimally distribute the permanent balls among the subtrees $T_{s_1}, \ldots, T_{s_l}$, the algorithm uses the function ALLOC which solves the resource allocation problem with two resources [10]. Namely, ALLOC$(f_1, \ldots, f_l; p, q)$ is the value of the optimal solution of

$$\begin{array}{ll} \text{Maximize} & \sum_{i=1}^{l} f_i(p_i, q_i) \\ \text{subject to} & \sum_{i=1}^{l} p_i = p \\ & \sum_{i=1}^{l} q_i = q \\ & p_i,\ q_i \text{ are nonnegative integers.} \end{array}$$

The optimal distribution $\{(p_1, q_1), \ldots, (p_l, q_l)\}$ of the resource allocation problem is computed by the function BUILD-ALLOC$(f_1, \ldots, f_l; p, q)$.

3.1 The algorithm

Now, we present the routines for computing INT and EXT in more details. If s is a leaf, then $T_s = \{s\}$ and the valuation of INT and EXT is given by the following formulae

$$\mathrm{INT}(T_s; n_1', n_2'; a) = \begin{cases} -\infty & \text{if } n_1' = n_2' = 0 \text{ or if } n_2' = 0 \text{ and } a > R_1, \\ 1 & \text{otherwise;} \end{cases} \tag{1}$$

$$\mathrm{EXT}(T_s; n_1', n_2'; r) = \begin{cases} 0 & \text{if } n_1' = n_2' = 0 \text{ and } r = -1, \\ 1 & \text{otherwise.} \end{cases} \tag{2}$$

Suppose now that s has at least one descendant. To evaluate the functions INT and EXT on the subtree T_s, the algorithm uses the values of the functions INT and EXT on the subtrees $T_{s_1}, \ldots, T_{s_l}$. Namely, the algorithm returns $\mathrm{INT}(T_s; n_1', n_2'; a) = \max\{I_1, I_2, I_3\}$ if $a \le R_1$, $\mathrm{INT}(T_s; n_1', n_2'; a) = \max\{I_2, I_3\}$ if $a > R_1$, and $\mathrm{EXT}(T_s; n_1', n_2'; r) = \max\{E_1, E_2\}$, where I_1, I_2, I_3, E_1, and E_2 are defined in the following way:

$I_1 \;= \mathrm{ALLOC}(f_1, \ldots, f_l; n'_1 - 1, n'_2) + 1$, where $f_i(p, q) = \mathrm{EXT}(T_{s_i}; p, q; R_1 - 1)$;

$I_2 \;= \mathrm{ALLOC}(g_1, \ldots, g_l; n'_1, n'_2 - 1) + 1$, where $g_i(p, q) = \mathrm{EXT}(T_{s_i}; p, q; R_2 - 1)$;

$I_3 \;= \max\{\mathrm{ALLOC}(h_1^{ja'}, \ldots, h_l^{ja'}; n'_1, n'_2) + 1 \;:\; j \in \{1, 2, \ldots, l\}, a' \in \{a, a+1, \ldots, R_2\}\}$, where $h_i^{ja'}(p, q) = \mathrm{EXT}(T_{s_i}; p, q; a'-1)$ for $i \neq j$ and $h_j^{ja'}(p, q) = \mathrm{INT}(T_{s_j}; p, q; a'+1)$;

$E_1 \;= \mathrm{INT}(T_s; n'_1, n'_2; r + 1)$;

$E_2 \;= \mathrm{ALLOC}(f'_1, \ldots, f'_l; n'_1, n'_2) + \delta(r)$, where $\delta(r) = 1$ if $r \geq 0$ and $\delta(r) = 0$ if $r = -1$, and $f'_i(p, q) = \mathrm{EXT}(T_{s_i}; p, q; \max\{-1, r - 1\})$.

If the INT entry equals I_1 (and $I_1 > I_3$), then a permanent ball B' of radius R_1 is centered at s and the remaining $n'_1 - 1$ R_1-balls and n'_2 R-balls are optimally distributed among the subtrees $T_{s_1}, \ldots, T_{s_l}$. Notice that in this case $a \leq R_1$ and that the relative radius in s of all permanent balls located in $T_s - \{s\}$ is less than R_1, otherwise the ball B' is useless, yielding $I_1 \leq I_3$. Therefore, there are no interactions among subtrees, from which we infer that the problems associated with the subtrees are independent, i.e., a covered vertex of T_{s_i} is either covered by B' or by a ball located in T_{s_i}. This explains why in order to compute I_1 we make a call of ALLOC with parameters $f_i(p, q) = \mathrm{EXT}(T_{s_i}; p, q; R_1 - 1), i = 1, \ldots, l$. The analysis of I_2 is similar. Now suppose that the INT entry equals I_3. In this case, the partial covering of T_s is done by permanent balls located in $T_s - \{s\}$. Let T_{s_j} be the subtree which hosts a permanent ball B' maximizing $rr(s, B') =: a'$. By definition of INT, we must have $a' \geq a$. Notice that every covered vertex of some subtree T_{s_i} is necessarily covered either by the ball B' or by a permanent ball located in T_{s_i}. Since we do not know a priori neither the subtree T_{s_j} providing the maximum nor a', we should test all possibilities. Now, for given $a' \in \{a, a+1, \ldots, R_2 - 1\}$ and $j \in \{1, \ldots, l\}$, since the ball B' with relative radius a' is located in T_{s_j}, in order to distribute the permanent balls among the subtrees $T_{s_1}, \ldots, T_{s_l}$ we make a call of ALLOC with parameters $h_i^{ja'}(p, q) = \mathrm{EXT}(T_{s_i}; p, q; a' - 1)$ for each $i = 1, \ldots, l$, $i \neq j$, and $h_j^{ja'}(p, q) = \mathrm{INT}(T_{s_j}; p, q; a' + 1)$.

Analogously, if the EXT entry equals E_1, then there exists a permanent ball B' located in T_s such that $rr(s, B') > r$. As a consequence, the ball of radius r centered at s is useless, whence we can use the result for INT, thus explaining why $E_1 = \mathrm{INT}(T_s; n'_1, n'_2; r + 1)$. Finally, if $E_2 > E_1$, then we search for an optimal distribution of $n'_1 + n'_2$ permanent balls in T_s such that $rr(s, B) \leq r$ holds for each permanent ball B. Then all covered vertices of any subtree T_{s_i} are necessarily covered either by a permanent ball located in T_{s_i} or by the ball of radius r centered at s. Therefore, the problems associated with the subtrees are independent and we make a call of ALLOC with parameters $f'_i(p, q) =$

EXT($T_{s_i}; p, q; \max\{-1, r-1\}$). Notice that if $r = -1$, then the vertex s is not covered neither by a permanent ball nor by the (empty) ball of radius -1 centered at s. In all other cases, s is covered by the ball of radius $r \geq 0$ centered at s. We conclude this subsection with a formal description of the algorithm.

Algorithm PARTIAL-MIXED-COVERING

Input. A tree $T = (V, E)$ and non negative integers R_1, R_2, n_1, n_2 ($R_1 < R_2$).
Output. A set of n_1 R_1-balls and n_2 R_2-balls maximizing the total number of covered vertices

Root T at some (non-leaf) vertex u and order the vertices of T using depth first search.
Initialize the values of INT and EXT for leaves of T using (1) and (2).
for current non-leaf vertex s
  do for $n'_1 \leftarrow 0$ **to** n_1
    do for $n'_2 \leftarrow 0$ **to** n_2
      do for $a \leftarrow 0$ **to** R_2
        do $I_1 \leftarrow \mathrm{ALLOC}(f_1, \ldots, f_l, n'_1 - 1, n'_2)$
          if $a \leq R_1$
            then $I_2 \leftarrow \mathrm{ALLOC}(g_1, \ldots, g_l, n'_1, n'_2 - 1)$
            else $I_2 \leftarrow -\infty$
          $I_3 \leftarrow \max\{\mathrm{ALLOC}(h_1^{ja'}, \ldots, h_l^{ja'}, n'_1, n'_2) \ : \ a' = a, \ldots, R_2, j = 1, \ldots, l\}$
          if $I_1 = \max\{I_1, I_2, I_3\}$
            then $a^* \leftarrow R_1, j^* \leftarrow \infty$
              $A \leftarrow \text{BUILD-ALLOC}(f_1, \ldots, f_l, n'_1 - 1, n'_2)$
          if $I_2 = \max\{I_1, I_2, I_3\}$
            then $a^* \leftarrow R_2, j^* \leftarrow \infty$
              $A \leftarrow \text{BUILD-ALLOC}(g_1, \ldots, g_l, n'_1, n'_2 - 1)$
          if $I_3 = \max\{I_1, I_2, I_3\}$
            then Let a^* and j^* be the values of a' and j yielding I_3.
              $A \leftarrow \text{BUILD-ALLOC}(h_1^{j^*a^*}, \ldots, h_l^{j^*a^*}, n'_1, n'_2)$
          $\mathrm{INT}(T_s; n'_1, n'_2; a) \leftarrow \max\{I_1, I_2, I_3\}$
          $\text{S-INT}(T_s; n'_1, n'_2; a) \leftarrow (A, a^*, j^*)$
  for $n'_1 \leftarrow 0$ **to** n_1
    do for $n'_2 \leftarrow 0$ **to** n_2
      do for $r \leftarrow -1$ **to** $R_2 - 1$
        do $E_1 \leftarrow \mathrm{INT}(T_s; n'_1, n'_2; r + 1)$
          $E_2 \leftarrow \mathrm{ALLOC}(f'_1, \ldots, f'_l, n'_1, n'_2)$
          if $E_1 = \max\{E_1, E_2\}$
            then $c \leftarrow 1$
              Extract the allocation A from S-INT($T_s; n'_1, n'_2; r + 1$).
          if $E_2 = \max\{E_1, E_2\}$
            then $c \leftarrow 2$
              $A \leftarrow \text{BUILD-ALLOC}(f'_1, \ldots, f'_l, n'_1, n'_2)$
          $\mathrm{EXT}(T_s; n'_1, n'_2; r) \leftarrow \max\{E_1, E_2\}$
          $\text{S-EXT}(T_s; n'_1, n'_2; r) \leftarrow (A, c)$
return BUILD-EXT($T_u, n_1, n_2, -1$)

To restore an optimal partial covering, the algorithm keeps in the tables S-INT and S-EXT the parameters of the distributions yielding the optimal value for the functions INT and EXT, and perhaps the radius of the permanent ball centered in the current vertex (if the optimal solution requires its location). The total space for these tables is equal to $n_1 n_2 R_2 O(\sum_{v\in V} \deg(s)) = O(n_1 n_2 R_2 n)$. Using the tables S-INT and S-EXT, an optimal location is computed by recursive functions BUILD-INT and BUILD-EXT in a downward manner. Each of these functions takes as input a vertex s and a list of parameters identifying a respective INT- or EXT-problem for s, and, using the information stored in the tables S-INT and S-EXT, decides if a permanent ball (and of what radius) should be centered at s, and specifies the parameters for its recursive call at each son s_i of s. After processing the root u of the tree T, the algorithm returns BUILD-EXT$(T_u, n_1, n_2, -1)$.

BUILD-INT(T_s, n'_1, n'_2, a)
if s is a leaf
  then return The set of n'_1 balls of radius R_1 and n'_2 balls of radius R_2 centered in s
  else $(A, a^*, j^*) \leftarrow$ S-INT$(T_s; n'_1, n'_2; a)$
    Let $\{(p_i, q_i) : i = 1, \ldots, l\}$ be the allocation A.
    if $j^* = \infty$
      then $\mathcal{B} \leftarrow \{B(s, a^*)\}$
        for $i \leftarrow 0$ **to** l **do** $\mathcal{B} \leftarrow \mathcal{B} \cup$ BUILD-EXT$(T_{s_i}, p_i, q_i, a^* - 1)$
      else $\mathcal{B} \leftarrow \emptyset$
        for $j \leftarrow 0$ **to** l
          do if $j = j^*$ **then** $\mathcal{B} \leftarrow \mathcal{B} \cup$ BUILD-INT$(T_{s_i}, p_j, q_j, a^* + 1)$
            else $\mathcal{B} \leftarrow \mathcal{B} \cup$ BUILD-EXT$(T_{s_i}, p_j, q_j, a^* - 1)$
    return $\mathcal{B}$

BUILD-EXT(T_s, n'_1, n'_2, r)
if s is a leaf
  then return The set of n'_1 balls of radius R_1 and n'_2 balls of radius R_2 centered in s
  else $(A, c) \leftarrow$ S-EXT$(T_s; n'_1, n'_2; r)$
    Let $\{(p_i, q_i) : i = 1, \ldots, l\}$ be the allocation A.
    if $c = 1$ **then** $\mathcal{B} \leftarrow$ BUILD-INT$(T_s, n'_1, n'_2, r + 1)$
      else (* $c = 2$ *)
        $\mathcal{B} \leftarrow \emptyset$
        for $j \leftarrow 0$ **to** l
          do $\mathcal{B} \leftarrow \mathcal{B} \cup$ BUILD-EXT$(T_{s_i}, p_i, q_i, \max\{-1, r - 1\})$
    return $\mathcal{B}$

3.2 Correctness and complexity

In this subsection, we establish the correctness and the complexity of the algorithm presented in Subsection 3.1.

Theorem 3.1 *The described algorithm correctly solves the* PARTIAL MIXED COVERING *problem in* $O(n_1^3 n_2^3 R_2^2 n^2)$ *time.*

PROOF. To prove the correctness of the algorithm, it suffices to show that all values of INT and EXT are correctly computed. This is shown by the following claims.

Claim 1 : *If* INT *and* EXT *are correctly evaluated on each of the subtrees* $T_{s_1}, \dots, T_{s_l}$, *then* INT *is correctly evaluated on* T_s.

Proof. Let $n_1' \in \{0, \dots, n_1\}$, $n_2' \in \{0, \dots, n_2\}$, and $a \in \{0, \dots, R_2\}$. Consider an optimal partial covering $\mathcal{C}^*$ of T_s with at most n_1' balls of radius R_1 and at most n_2' balls of radius R_2 located in T_s, given that the relative radius in s of some ball of $\mathcal{C}^*$ is at least a. Suppose additionally that all balls B of $\mathcal{C}^*$ are necessary, i.e., $\mathcal{C}^*$ minus B is no longer an optimal covering for the parameters n_1', n_2', and a. Let B' be a ball of $\mathcal{C}^*$ with a maximal relative radius in s and set $a' = rr(s, B')$. Notice that, if $\mathcal{C}^*$ contains a ball B centered at s, then necessarily $B' = B$. Indeed, since B cannot be removed from $\mathcal{C}^*$ without violating the optimality of this partial covering, the relative radius in s of any ball of $\mathcal{C}^*$ different from B is less than the radius of B. Since $rr(s, B)$ equals the radius of B, we conclude that $B' = B$.

First suppose that B' is centered at s. From what has been shown above, we deduce that $rr(s, B) < a'$ for any ball $B \in \mathcal{C}^*$ different from B'. Thus every vertex of $T_{s_i}, i = 1 \dots, l$, which is covered by $\mathcal{C}^*$ is covered either by a ball located in T_{s_i} or by B'. Since EXT is correctly evaluated on each of the subtrees T_{s_i}, $i \in \{1, \dots, l\}$, the number of vertices of T_{s_i} covered by $\mathcal{C}^*$ is at most $\mathrm{EXT}(T_{s_i}; p_i, q_i; a' - 1)$. Since the function ALLOC optimally distributes the remaining $n_1' + n_2' - 1$ balls, the value of I_1 (if B' has radius R_1) or of I_2 (if B' has radius R_2) is at least $|T_s \cap (\cup\{B \in \mathcal{C}^*\})|$.

Now suppose that no ball of $\mathcal{C}^*$ is centered at s and assume that the ball B' is located in the subtree T_{s_j}. Then $rr(s, B') = a' \geq a$ by definition of INT. For any $i = 1, \dots, l$, every vertex of T_{s_i} covered by $\mathcal{C}^*$ is covered either by a ball centered at T_{s_i} or by B'. Hence the number of covered by $\mathcal{C}^*$ vertices of T_{s_i} is at most $\mathrm{EXT}(T_{s_i}; p_i, q_i; a' - 1)$ for $i \neq j$ (because $rr(s_i, B') = a' - 1$) and the number of covered by $\mathcal{C}^*$ vertices of T_{s_j} is at most $\mathrm{INT}(T_{s_j}; p_j, q_j; a' + 1)$ (because $rr(s_j, B') = a' + 1$). From the optimality of ALLOC we infer that I_3 is at least $|T_s \cap (\cup\{B \in \mathcal{C}^*\})|$. □

Claim 2 : *If* INT *and* EXT *are correctly evaluated on each of the subtrees* $T_{s_1}, \dots, T_{s_l}$ *and* INT *is correctly evaluated on* T_s, *then* EXT *is correctly evaluated on* T_s.

Proof. Let $n_1' \in \{0, \dots, n_1\}$, $n_2' \in \{0, \dots, n_2\}$, and $r \in \{-1, \dots, R_2 - 1\}$. Consider an optimal partial cover $\mathcal{C}^*$ of T_s with n_1' balls of radius R_1, n_2' balls of radius R_2 located in

T_s, and an additional ball B' of radius r centered at s. Suppose additionally that all balls of $\mathcal{C}^*$ are necessary.

First suppose that there is a ball $B \in \mathcal{C}^*$ different from B' obeying $rr(s, B) > r$. Then no covered vertex of T_s is covered solely by B', therefore B' is useless. As a consequence, we conclude that the number of vertices covered by $\mathcal{C}^*$ equals $\mathrm{INT}(T_s; n'_1, n'_2; r + 1)$.

So, suppose that $rr(s, B) \leq r$ for every ball B of $\mathcal{C}^*$ different from B'. Then every covered by $\mathcal{C}^*$ vertex of each subtree $T_{s_i}, i = 1, \ldots, l$, is covered either by a ball located in T_{s_i} or by B'. If one denote by p_i and q_i the number of balls of radius R_1 and R_2 of $\mathcal{C}^*$ located in the subtree T_{s_i}, then the number of vertices of T_{s_i} covered by $\mathcal{C}^*$ is at most $\mathrm{EXT}(T_i; p_i, q_i; \max\{-1, r - 1\})$. From the optimality of ALLOC we deduce that E_2 is at least $|T_s \cap (\cup\{B \in \mathcal{C}^*\})|$, establishing the result. $\square$

To find the complexity of the algorithm, we will estimate the number of operations necessary to evaluate the functions INT and EXT in some vertex s. One computation of INT requires one call of ALLOC to evaluate I_1 and I_2 and at most $R_2\deg(s)$ to evaluate I_3. Analogously, one computation of EXT requires one call of the function ALLOC. There is a dynamic programming algorithm with complexity $O(lp^2q^2)$ for solving the resource allocation problem $\mathrm{ALLOC}(f_1, \ldots, f_l; p, q)$ with two resources (see, for example, pages 207-208 of [10]). Therefore, if INT and EXT have been evaluated on the subtrees $T_{s_1}, \ldots, T_{s_l}$, then the $2n_1n_2(R_2+1)$ values of INT and EXT for T_s can be found in $O(n_1^3n_2^3R_2^2\deg^2(s))$ time. Since $\sum_{s\in V}\deg(s) = 2n-2$, the total complexity of the algorithm is $O(n_1^3n_2^3R_2^2n^2)$. $\square$

3.3 Related problems

In this subsection, we present several further problems, related to PARTIAL MIXED COVERING, and which can be solved by adapting the algorithm described above.

(i) A natural generalization of the PARTIAL MIXED COVERING problem is that of maximizing the number of covered vertices by n_1 balls of radius R_1, n_2 balls of radius R_2, $\ldots$, n_k balls of radius R_k, where $R_1 < R_2 < \ldots < R_k$. We call the resulting problem GENERALIZED PARTIAL MIXED COVERING. Our algorithm can be modified to solve it in $O(R_k^2n^2\Pi_{i=1}^k n_i^3)$ time (which is less than $O(n^{3k+4})$). For this, in the computation of INT we distinguish $k+1$ cases: one for each kind of permanent balls centered at s plus one dealing with the case when no permanent ball is centered at s. Then the INT entry equals to the maximum of these $k+1$ values $I_1, \ldots, I_k, I_{k+1}$. Additionally, instead of solving each time a resource allocation problem with two resources, we solve via dynamic programming a resource allocation problem with k resources (requiring $O(\deg(s)\Pi_{i=1}^k n_i^2)$ time per instance).

(ii) The problem (i) can be further generalized in the following way: given the integers $0 \leq R_1 < \ldots R_k$ and the positive integers $n_2, \ldots, n_k$, locate n_i balls of radius R_i, $i = 2, \ldots, k$, so that the remaining part of the tree can be covered with a minimum number n_1 of balls of radius R_1. To solve this problem, we can solve a sequence of GENERALIZED PARTIAL MIXED COVERING problems, one for each value of n_1 varying between 0 and n, and return the smallest value of n_1 for which the whole tree is covered. However, we can solve this problem more efficiently by modifying the algorithm presented above. For this, we modify the definition of the functions INT and EXT, and, instead of taking maxima and performing maximization in the resource allocation problem, we take minima and solve a minimization resource allocation problem. For example, $\mathrm{EXT}(T_s; n'_2, \ldots, n'_k; r)$ is defined to be the minimum number of balls of radius R_1 which, together with n'_2 balls of radius R_2, …, n'_k balls of radius R_k, and an additional ball of radius r centered at s, cover the subtree T_s ($\mathrm{EXT}(T_s; n'_2, \ldots, n'_k; a)$ is defined accordingly). The complexity of this algorithm is $O(R_k^2 n^2 \Pi_{i=2}^k n_i^3)$.

In the particular case $k = 2$, we obtain the problem of covering a tree T with n_2 balls of radius R_2 and a minimum number of balls of radius R_1, which can be solved in $O(n_2^3 n^2 R_2^2)$ time. Now, if we want to solve the MIXED COVERING problem with $R_2 = R \leq 2, R_1 = R - 1 > 0$, and $f(n_1, n_2) = n_1 + \frac{n_2(n_2-1)}{2}$, then we may suppose that $n_2 \leq \sqrt{n}$. Since the dynamic programming algorithm keeps the solutions of subproblems for all vertices (in particular for u) and all $n_2 \leq \sqrt{n}$, it suffices to select the solution minimizing f. The algorithm in this case has complexity $O(n^{3.5} R_2^2)$.

(iii) Another natural generalization of partial mixed covering problem is the following partial mixed list covering problem. Namely, additionally to the input data of (i), for each vertex s of the tree T, a list $L_s \subseteq \{1, \ldots, k\}$ is given, which defines what kinds of permanent balls are allowed to be centered at s (the vertices s with $L_s = \emptyset$ are forbidden for establishing permanent balls). Then each INT entry is again the maximum of $I_1, \ldots, I_k, I_{k+1}$, but, in this case, $I_j = -\infty$ for any $j \notin L_s$.

(iv) Several weighted versions of the PARTIAL MIXED COVERING problem and its variations (i)-(iii) can be solved using the same approach. For instance, the tree $T = (V, E)$ can be endowed with a length function $l : E \to \mathbb{R}_+$. In this case, the balls will be defined with respect to the distance induced by this length function. Also, we may want to consider that covering some vertex u induces a gain π_u and we wish to find a partial mixed covering maximizing the total gain of covered vertices. In this case, each time when the current vertex s is covered, we add π_s (instead of 1) to the value of INT or EXT.

(v) The GENERALIZED PARTIAL MIXED COVERING problem in which the balls of certain radii are replaced by edge-balls can be easily formulated in form of (iii) or (iv). For this, we subdivide every edge e of the tree $T = (V, E)$ by introducing a new vertex

s_e, thus obtaining a new tree T'. Then $d_{T'}(u,v) = 2d_T(u,v)$ for any two vertices u,v of T. Instead of n_i balls of a radius R_i consider n_i balls of radius $2R_i$ of T' and we require that they can be centered only at the vertices of T. Instead of n_i edge-balls of a radius R_i we consider n_i balls of radius $2R_i+1$ of T' which can be centered only at the new vertices of T'. Define the gain of covering any old vertex to be 1 and the gain of covering any new vertex to be 0. Then solving the resulting partial mixed covering problem on the tree T' is equivalent to solving the initial problem with balls and edge-balls on the tree T.

4 The augmentation problem with odd diameter constraints

In this section, we apply MIXED COVERING to derive a factor $2+\frac{1}{\delta}$ (for any integer $\delta > 0$) approximation algorithm for the augmentation problem with odd diameter constraints $D = 2R+1$ on trees. For this, we compute in $O(n^{3.5}R^2)$ time a mixed covering of T with n'_1 balls of radius $R-1$ and n'_2 balls of radius R which minimizes the function $f(n'_1, n'_2) = n'_1 + n'_2(n'_2-1)/2$. The augmentation algorithm returns the set F of new edges running between the centers of any pair of balls of radius R and between the center of any ball of radius $R-1$ and the center of some ball of radius R. Set ALG $:= n'_1 + \frac{n'_2(n'_2-1)}{2}$ and let OPT denote the number of edges of an optimal solution of the problem ADC for T.

Theorem 4.1 *For any integer $\delta > 0$, we have* $\mathrm{ALG} \leq (2+\frac{1}{\delta})\mathrm{OPT} + O(\delta^5)$.

PROOF. First we show that the augmented graph $H = (V, E \cup F)$ has diameter at most $2R+1$. Pick two arbitrary vertices $u, v \in V$. Suppose that in the mixed covering u belongs to a ball centered at the vertex p and v belongs to a ball centered at the vertex q. If both these balls have radius R, then p and q are connected by a new edge, therefore $d_H(u,v) \leq R+1+R = 2R+1$. If one ball has radius R and another one has radius $R-1$, then $d_H(p,q) \leq 2$, whence $d_H(u,v) \leq (R-1)+2+R = 2R+1$. Finally, if both balls centered at p and q have radius $R-1$, then $d_H(p,q) \leq 3$ according to the algorithm, therefore $d_H(u,v) \leq (R-1)+3+(R-1) = 2R+1$. This shows that F is a feasible augmentation of T.

Let E' be an optimal solution for ADC and let $G' = (V, E \cup E')$ be the augmented graph. Denote by C_1 the set of end-vertices of edges from E' and by $\mathcal{B}_1$ the set of balls of radius $R-1$ centered at vertices of C_1. Set $n_1 := |C_1|$. Let $Q := V - \bigcup\{B : B \in \mathcal{B}_1\}$. Consider the graph $T^{2R+1}(Q)$. As we noticed in Section 2, $T^{2R+1}(Q)$ is a perfect graph, therefore the size of a maximum stable set $Y \subseteq Q$ of $T^{2R+1}(Q)$ equals the minimum number n_2 of edge-balls of radius R covering the set Q. Denote by $\mathcal{B}_2$ the family of edge-balls in this covering. The *cluster* of a vertex $x \in Q$ is the set $C_x = \{c \in C_1 : d_T(x,c) = R\}$. All vertices of $C_1 \setminus C_x$ are at distance $> R$ from x, therefore, if the cluster C_x is empty, then

x must be at distance $\leq 2R+1$ in T from all vertices of Q. Notice also that two clusters C_x and C_y are disjoint provided $d_T(x,y) \geq 2R+1$.

Claim 1 : *If two vertices $x, y \in Q$ are not adjacent in $T^{2R+1}(Q)$, then there exists at least one added edge running between the clusters C_x and C_y. In particular,* $\mathrm{OPT} = |E'| \geq \frac{n_2(n_2-1)}{2}$.

Proof. Consider a path P of length $\leq 2R+1$ connecting the vertices x and y in the augmented graph G'. Since $d_T(x,y) > 2R+1$, this path will necessarily use one or several new edges. Denote by x' and y' the closest to x and y, respectively, vertices of $P \cap C_1$. Since $d_T(x,x') \geq R$, $d_T(y,y') \geq R$, and $C_x \cap C_y = \emptyset$, we conclude that $x' \in C_x, y' \in C_y$, and x' and y' must be connected by an edge of E'.

Since any two vertices x, y of Y are not adjacent in $T^{2R+1}(Q)$, there exists at least one new edge connecting the clusters C_x and C_y. Since the clusters of vertices of Y are pairwise disjoint and $|Y| = n_2$, these new edges are pairwise distinct and therefore there exist at least $\frac{n_2(n_2-1)}{2}$ such edges. □

Claim 2 : $\mathrm{OPT} \geq \frac{n_1}{2}$.

Proof. By definition, $n_1 = |C_1|$, where C_1 is the set of end-vertices of edges of E'. Obviously $|E'| \geq \frac{|C_1|}{2}$, the worst case occurring when E' is a perfect matching on C_1. □

Next we will refine the lower bounds for OPT provided by Claims 1 and 2. Also from the covering $\mathcal{B}_1 \cup \mathcal{B}_2$ of T with n_1 balls of radius $R-1$ and n_2 edge-balls of radius R, we will derive a feasible solution for the mixed covering problem, thus establishing an upper bound on the number ALG of edges added by the algorithm. Let n_0 be the minimum number of balls of radius R of T covering the set Q and let $\mathcal{B}_0$ denote the set of balls from this covering. Since every edge-ball of radius R can be covered by two balls of radius R, we obtain the following inequality:

Claim 3 : $n_0 \leq 2n_2$.

For a fixed integer $\Delta \geq \delta + 2$, we perform the following operations with the balls of the collection $\mathcal{B}_0$. Initialize $\mathcal{A} := \emptyset$, and $\mathcal{C} := \mathcal{B}_0$. Test each ball B of the current collection $\mathcal{C}$, and if the set $(B \cap Q) \setminus ((\bigcup\{B' \in \mathcal{C} : B' \neq B\}) \cup (\bigcup\{B' \in \mathcal{A}\}))$ can be covered by at most Δ balls of radius $R-1$, then add these balls to $\mathcal{A}$ and remove B from $\mathcal{C}$. Repeat this operation until no ball of the current collection $\mathcal{C}$ can be further replaced by Δ balls of radius $R-1$. Set $\alpha := |\mathcal{A}|$ and $\gamma := |\mathcal{C}|$. Notice that $\mathcal{B}_1 \cup \mathcal{A}$ and $\mathcal{C}$ constitute a mixed covering of the tree T with $n_1 + \alpha$ balls of radius $R-1$ and γ balls of radius R.

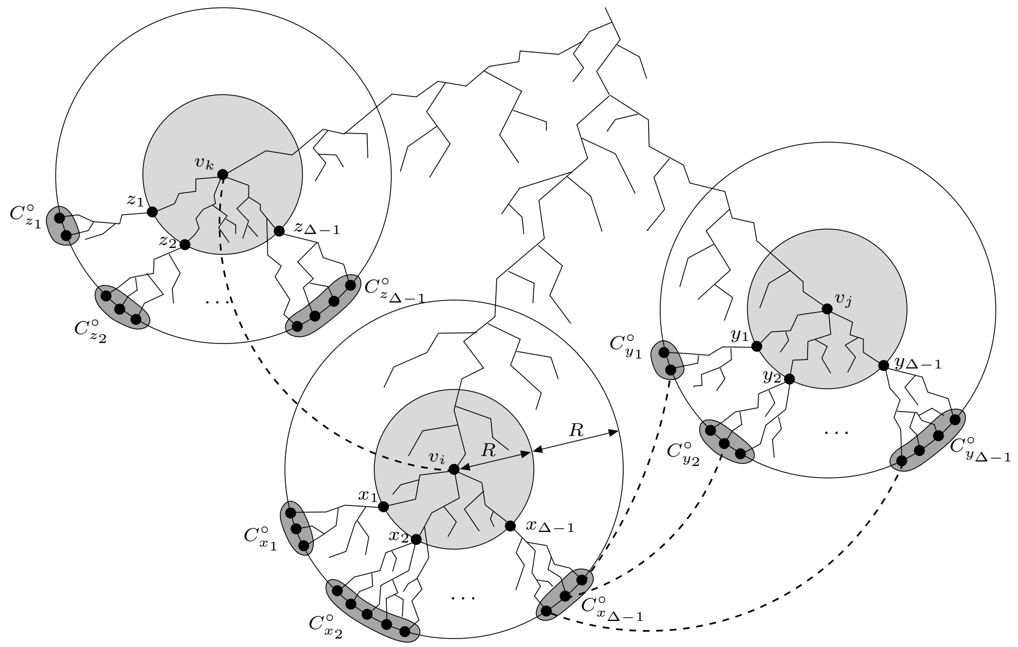

Figure 1.

Claim 4: *Every ball $B = B_R(v)$ of the resulting collection $\mathcal{C}$ contains Δ vertices*

$$x_1, \ldots, x_\Delta \in B^+ := (B \cap Q) \setminus ((\bigcup\{B' \in \mathcal{C} : B' \neq B\}) \cup (\bigcup\{B' \in \mathcal{A}\}))$$

such that $d_T(x_i, v) = R$ for all $i = 1, \ldots, \Delta$ and $d_T(x_i, x_j) = 2R$ for all distinct $i, j \in \{1, \ldots, \Delta\}$.

Proof. Since the ball B survived the last test, the minimum number of balls of radius $R-1$ necessary to cover the set B^+ is at least $\Delta+1$. Due to the duality between packing and covering with balls in trees (applied to the perfect graph $T^{2R-2}(B^+)$), B^+ contains $\Delta+1$ vertices $S = \{x_1, \ldots, x_\Delta, x_{\Delta+1}\}$ forming a stable set of $T^{2R-2}(B^+)$. Since $2R-1 \leq d_T(x_i, x_j) \leq d_T(x_i, v) + d_T(v, x_j) \leq R + R = 2R$, we conclude that at least Δ vertices, say $x_1, \ldots, x_\Delta$, must be located at distance R from the center v of B, moreover, the paths connecting v with these vertices pairwise intersect only in v, otherwise we will find two vertices of S having distance $\leq 2R-2$. Thus $d_T(x_i, x_j) = 2R$ for arbitrary $i, j \in \{1, \ldots, \Delta\}$, $i \neq j$. $\square$

Suppose that the balls of the collection $\mathcal{C}$ are ordered $B_1, \ldots, B_\gamma$ in the following way: root the tree $T(Q)$ at the center of an arbitrary ball of $\mathcal{C}$ and for two balls $B, B' \in \mathcal{C}$ set $B = B_i, B' = B_j$, where $i < j$, provided the center of the ball B is closer or at the same distance to the root than the center of the ball B', breaking ties arbitrarily. Denote by

v_i the center of the ball B_i, $i = 1, \dots, \gamma$. Consider the set of Δ vertices $x_1, \dots, x_\Delta$ of the ball B_i described in Claim 4. Since the paths $P(v_i, x_1), \dots, P(v_i, x_\Delta)$ pairwise intersect solely in the vertex v_i, at most one such path, say $P(v_i, x_\Delta)$, may pass via the father of v_i. Therefore the set $F_i := \{x_1, \dots, x_{\Delta-1}\}$ consists solely of descendants of the vertex v_i. On the other hand, if $y \in F_j, j < i$, then $d_T(v_i, y) > R = d_T(v_j, y)$, whence y cannot be a descendant of v_i. For each vertex $x \in F_i$, set $C_x^\circ := C_x - \{v_i\}$. For an illustration of this and other notions introduced above, see Fig. 1.

Claim 5 : *For all* $i, j \in \{1, \dots, \gamma\}, j < i$, *if* $x \in F_i$ *and* $y \in F_j$, *then* $d_T(x, y) \geq 2R + 1$. *If* $d_T(x, y) = 2R + 1$, *then* v_j *is the closest to* v_i *center of a ball of* $\mathcal{C}$ *which is an ancestor of* v_i.

Proof. Let z be the nearest common ancestor of v_i and v_j. If $z \notin \{v_i, v_j\}$, then z is also the nearest common ancestor of x and y, because x is a descendant of v_i and y is a descendant of v_j. In this case, we deduce that

$$d_T(x, y) = d_T(x, v_i) + d_T(v_i, z) + d(z, v_j) + d_T(v_j, y) \geq R + 1 + 1 + R > 2R + 1.$$

On the other hand, if $z \in \{v_i, v_j\}$, then $z = v_j$, because $j < i$. Since y is not a descendant of v_i, we conclude that $v_i \in P(x, y)$, and

$$d_T(x, y) = d_T(x, v_i) + d_T(v_i, y) \geq R + (R + 1) = 2R + 1.$$

Hence, if $d_T(x, y) = 2R + 1$, then v_i is a descendant of v_j and it remains to show that no other center v_k of a ball of $\mathcal{C}$ can be located on the path between v_i and v_j. Suppose the contrary: then, since y is not a descendant of v_k, the vertices v_i and v_k belong to the path $P(x, y)$, hence

$$d_T(x, v_i) + d_T(v_i, v_k) + d_T(v_k, y) \geq R + 1 + (R + 1) = 2R + 2,$$

yielding a contradiction. $\square$

For each $i = 2, \dots, \gamma$, denote by $v_{i'}$ (if it exists) the closest to v_i center of a ball of $\mathcal{C}$ which is an ancestor of v_i. The following assertion is an immediate consequence of Claim 5.

Claim 6 : *For any* $i = 2, \dots, \gamma$, *if* $j, k \in \{1, \dots, i-1\} - \{i'\}$, $j \neq k$, *and* $x \in F_i, y \in F_j$, *and* $z \in F_k$, *then* $d_T(x, y) > 2R + 1, d_T(x, z) > 2R + 1$, *and* $d_T(y, z) \geq 2R + 1$. *In particular, the clusters* C_x, C_y, *and* C_z *are pairwise disjoint.*

We apply this claim to provide new lower bounds for OPT. For $i = 2, \dots, \gamma$, set $\Gamma_i = \bigcup_{j<i} \bigcup_{y \in F_j} C_y$. Let β_i be the number of edges of the optimal solution E' running between the vertex v_i and the clusters from Γ_i. For a vertex $x \in F_i$, denote by $\kappa(x)$

the number of edges of E' running between the cluster C_x and the clusters from Γ_i. Let $\kappa_i = \min\{\kappa(x) : x \in F_i\}$. Notice that $\kappa_i \geq \beta_i$, because v_i belongs to all clusters $C_x, x \in F_i$. On the other hand, $\kappa_i \geq i-2$, because there is an added edge between C_x and every cluster C_y such that $d_T(x,y) > 2R+1$ (see Fig. 1 for an illustration), and, by Claim 6, Γ_i contains at least $i-2$ pairwise disjoint clusters C_y, obeying $d_T(x,y) > 2R+1$.

Claim 7 : OPT $\geq (\Delta-1)\sum_{i=2}^{\gamma}\kappa_i - (\Delta-2)\sum_{i=2}^{\gamma}\beta_i$.

Proof. For each $i = 2, ..., \gamma$, there are β_i edges of E' between v_i and vertices of Γ_i, therefore, for any vertex $x \in F_i$ at least $\kappa_i - \beta_i$ edges of E' run between C_x° and Γ_i. Since the clusters of the $\Delta-1$ vertices from F_i pairwise intersect solely in v_i, at least $(\Delta-1)(\kappa_i-\beta_i)$ distinct edges of E' run between $\bigcup_{x\in F_i} C_x^\circ$ and Γ_i. Moreover, since the clusters of any two vertices $y \in F_j$ and $z \in F_k, j \neq k$, are disjoint, thus we obtain

$$\text{OPT} \geq \sum_{i=2}^{\gamma}[\beta_i + (\Delta-1)(\kappa_i-\beta_i)] = (\Delta-1)\sum_{i=2}^{\gamma}\kappa_i - (\Delta-2)\sum_{i=2}^{\gamma}\beta_i. \qquad \square$$

Claim 8 : OPT $\geq \frac{n_1-\beta+\sum_{i=2}^{\gamma}\beta_i}{2}$*, where β is the number of vertices $v_i, i = 2, \ldots, \gamma$, such that $\beta_i > 0$.*

Proof. First notice that for any vertex v_i with $\beta_i > 0$, the β_i edges of E' incident to v_i form a star with $\beta_i + 1$ vertices. Each of the $n_1 - \sum_{i=2}^{\gamma}\beta_i - \beta$ remaining vertices of C_1 is incident to an added edge, yielding at least $\frac{n_1-\sum_{i=2}^{\gamma}\beta_i-\beta}{2}$ other edges of E', the worst case being a perfect matching. This shows that

$$\text{OPT} \geq \sum_{i=2}^{\gamma}\beta_i + \frac{n_1-\sum_{i=2}^{\gamma}\beta_i-\beta}{2} = \frac{n_1+\sum_{i=2}^{\gamma}\beta_i-\beta}{2}. \qquad \square$$

Claim 9 : ALG $\leq n_1 - \beta + \Delta(n_0-\gamma) + \frac{\gamma(\gamma-1)}{2}$.

Proof. Notice that every v_i with $\beta_i > 0$ is a center of a ball of radius R of $\mathcal{C}$ and a center of a ball of radius $R-1$ of $\mathcal{B}_1$. Remove those β balls from $\mathcal{B}_1$. Together with $\mathcal{A}$ and $\mathcal{C}$, the resulting collection $\mathcal{B}_1$ form a mixed covering of T with γ balls of radius R and at most $n_1 - \beta + \Delta(n_0-\gamma)$ balls of radius $R-1$. This covering gives rise to a feasible solution of the augmentation problem using at most $n_1 - \beta + \Delta(n_0-\gamma) + \frac{\gamma(\gamma-1)}{2}$ new edges. $\square$

First assume that $\gamma \geq 2$. From Claims 7 and 8 we obtain

$$(2+\frac{1}{\delta})\text{OPT} \geq n_1 - \beta + \sum_{i=2}^{\gamma}\beta_i + \frac{\Delta-1}{\delta}\sum_{i=2}^{\gamma}\kappa_i - \frac{\Delta-2}{\delta}\sum_{i=2}^{\gamma}\beta_i$$

$$= n_1 - \beta + (\frac{\Delta - 2}{\delta} - 1)\sum_{i=2}^{\gamma}(\kappa_i - \beta_i) + (1 + \frac{1}{\delta})\sum_{i=2}^{\gamma}\kappa_i$$

$$\geq n_1 - \beta + (1 + \frac{1}{\delta})\sum_{i=2}^{\gamma}\kappa_i,$$

where the last inequality follows from $\kappa_i \geq \beta_i$ for $i = 2, \ldots, \gamma$ and $\Delta \geq \delta + 2$. Since $\kappa_i \geq i - 2$ for $i = 2, \ldots, \gamma$, we conclude that

$$(2 + \frac{1}{\delta})\text{OPT} \geq n_1 - \beta + (1 + \frac{1}{\delta})\frac{(\gamma - 1)(\gamma - 2)}{2}.$$

In order to ensure $(2 + \frac{1}{\delta})\text{OPT} \geq \text{ALG}$, in view of last inequality and Claim 9 it suffices to show that

$$n_1 - \beta + (1 + \frac{1}{\delta})\frac{(\gamma - 1)(\gamma - 2)}{2} \geq n_1 - \beta + \Delta(n_0 - \gamma) + \frac{\gamma(\gamma - 1)}{2}.$$

Taking $\Delta = \delta + 2$, after some elementary transformations this inequality can be rewritten as $f(\gamma) = \gamma^2 + b\gamma - c \geq 0$, where $b = 2\delta^2 + 2\delta - 3$ and $c = (2\delta^2 + 4\delta)n_0 - 2\delta - 2$. Since $b > 0$ because δ is a positive integer, the inequality $f(\gamma) > 0$ holds if $c \leq 0$. Now, if $c > 0$, the inequality holds for any $\gamma \geq \gamma_0$, where γ_0 is the largest solution of the quadratic equation $f(\gamma) = 0$ (the exact value of γ_0 will be given below). Therefore, if $\gamma \geq \gamma_0$, then $\text{ALG} \leq (2 + \frac{1}{\delta})\text{OPT}$.

Now, suppose that $\gamma \leq \gamma_0$. Using the lower bounds for OPT established in Claims 1 and 2, we obtain $(2 + \frac{1}{\delta})\text{OPT} \geq n_1 + \frac{n_2^2}{2\delta} - \frac{n_2}{2\delta}$. On the other hand, from Claim 9 we deduce that $\text{ALG} \leq n_1 + n_0\Delta + \frac{\gamma^2}{2}$. Since $n_0 \leq 2n_2$ by Claim 3,, $\Delta = \delta + 2$, and $\gamma \leq \gamma_0$, we have $\text{ALG} \leq n_1 + 2(\delta + 2)n_2 + \frac{\gamma_0^2}{2}$. Now, by definition of γ_0,

$$\frac{\gamma_0^2}{2} = \frac{[\sqrt{(2\delta^2 + 2\delta - 3)^2 + 8((\delta^2 + 2\delta)n_0 - \delta - 1)} - (2\delta^2 + 2\delta - 3)]^2}{8}$$

$$< \frac{(2\delta^2 + 2\delta - 3)^2 + 8((\delta^2 + 2\delta)n_0 - \delta - 1) + (2\delta^2 + 2\delta - 3)^2}{8}$$

$$< \delta^2(\delta + 1)^2 + 2\delta(\delta + 2)n_2 - \delta - 1,$$

because $(p - q)^2 \leq p^2 + q^2$ if $p, q \geq 0$, $2\delta^2 + 2\delta - 3 < 2\delta(\delta + 1)$, and $n_0 \leq 2n_2$. Comparing the lower bound for $(2 + \frac{1}{\delta})\text{OPT}$ with the upper bounds for ALG and $\frac{\gamma_0^2}{2}$, the desired inequality $\text{ALG} \leq (2 + \frac{1}{\delta})\text{OPT}$ holds if

$$n_2^2 - n_2 \geq 2\delta^3(\delta + 1)^2 + 4\delta^2(\delta + 2)n_2 + 4\delta(\delta + 2)n_2 - 2\delta^2 - 2\delta,$$

i.e., provided

$$g(n_2) = n_2^2 - n_2(4\delta^3 + 12\delta^2 + 8\delta + 1) - (2\delta^5 + 4\delta^4 + 2\delta^3 - 2\delta^2 - 2\delta) \geq 0.$$

As a result, we conclude that ALG $\leq (2 + \frac{1}{\delta})$OPT holds for all n_2 larger or equal to the largest solution n_2^+ of the quadratic equation $g(n_2) = 0$, otherwise, if $n_2 < n_2^+$, we have ALG $- (2 + \frac{1}{\delta})$OPT $\leq -\frac{g(n_2)}{2\delta} \leq -\frac{g(2\delta^3+6\delta^2+4\delta+\frac{1}{2})}{2\delta} = O(\delta^5)$. The case $\gamma = 0$ or 1 can be settled in a similar way, using the inequalities $(2 + \frac{1}{\delta})$OPT $\geq n_1 + \frac{n_2^2}{2\delta} - \frac{n_2}{2\delta}$ and ALG $\leq n_1 + 2n_2(\delta + 2)$. Therefore, in all cases we obtain that ALG $\leq (2+\frac{1}{\delta})OPT+O(\delta^5)$, concluding the proof of the theorem. □

Remark 1. In particular cases $\delta = 1, 2$, and 3, we obtain the following values for the additive error between ALG and $(2 + \frac{1}{\delta})$OPT: 80, 621, 2560, respectively.

Remark 2. Using the factor $2+\frac{1}{\delta}$ augmentation algorithm for ADC analyzed in Theorem 4.1 and the biconnectivity augmentation algorithm of Eswaran and Tarjan [8], we obtain a factor $3 + \frac{1}{\delta}$ approximation algorithm for the problem ADC on trees and odd diameters D with an additional requirement that the resulting augmented graph is biconnected (this problem is known to be NP-hard on trees [3]). A polynomial time algorithm with performance guarantee 4 is presented in [11]; see also [3] for related results.

Remark 3. Notice that the optimal solutions of the problem ADC with $D = 2R+1$ may contain cliques of arbitrary size k. For this, consider the tree T consisting of a star S in which the center v is adjacent to its tips $v_1, \ldots, v_k$, plus k other pairwise disjoint stars $S_1, \ldots, S_k$, where the star S_i consists of $K \gg k$ paths of length R pairwise intersecting only in the central vertex v_i. The diameter of T is $2R + 2$, and, in order to decrease it to $2R + 1$, the best way is to add an edge between any pair of centers of the stars.

Remark 4. In case of $D = 2R$, the factor 2 approximation algorithm for the problem ADC on trees given in [3] computes an optimal mixed covering with one ball of radius R and a minimum number of balls of radius $R - 1$, and draws an edge between the center of each $(R - 1)$-ball and the center of the R-ball. We conjecture that *this algorithm for even D as well as the algorithm for odd D analyzed in this paper actually are optimal up to an additive constant error term.*

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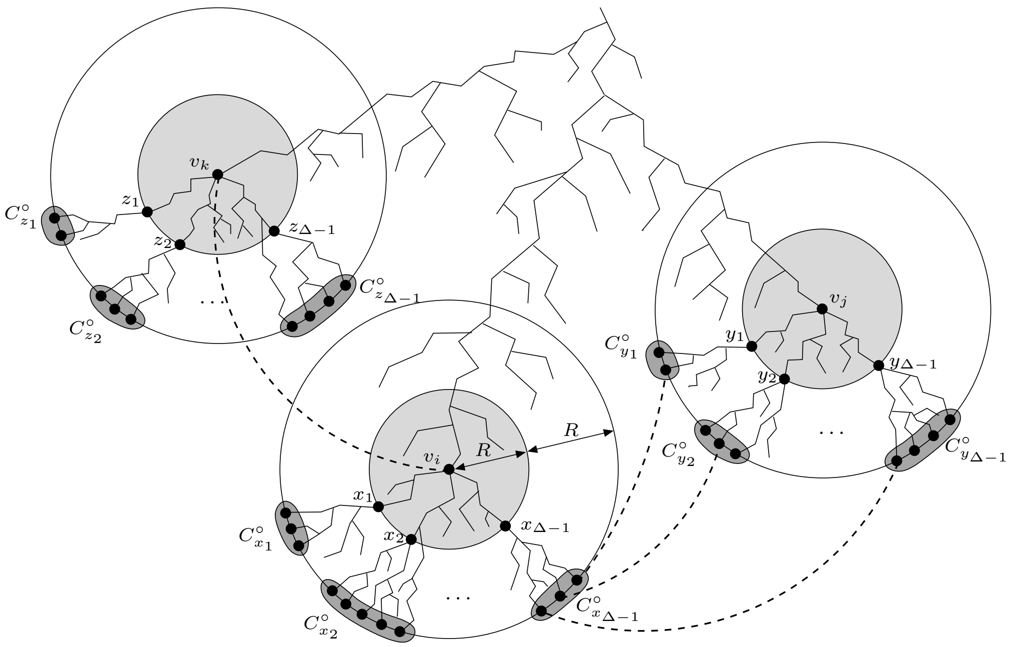

v_k
z_1
z_2
$z_{\Delta-1}$
$C^\circ_{z_1}$
$C^\circ_{z_2}$
$C^\circ_{z_{\Delta-1}}$
v_i
R
R
x_1
x_2
$x_{\Delta-1}$
$C^\circ_{x_1}$
$C^\circ_{x_2}$
$C^\circ_{x_{\Delta-1}}$
v_j
y_1
y_2
$y_{\Delta-1}$
$C^\circ_{y_1}$
$C^\circ_{y_2}$
$C^\circ_{y_{\Delta-1}}$

Figure 1.