A Simple Algorithm for Optimal Search Trees with Two-Way Comparisons

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Abstract

We present a simple $O(n^4)$-time algorithm for computing optimal search trees with two-way comparisons. The only previous solution to this problem, by Anderson et al., has the same running time, but is significantly more complicated and is restricted to the variant where only successful queries are allowed. Our algorithm extends directly to solve the standard full variant of the problem, which also allows unsuccessful queries and for which no polynomial-time algorithm was previously known. The correctness proof of our algorithm relies on a new structural theorem for two-way-comparison search trees.

1 Introduction

Search trees are among the most fundamental data structures for storing and accessing data. Static search trees are used in applications where the set of possible queries and their frequencies are known in advance, in which case a single tree can be precomputed and used throughout the application to handle lookup queries. Such a tree is optimal if it minimizes the expected search cost. The problem of computing optimal search trees has been studied extensively in various forms since the 1960s. Perhaps the most famous example is Knuth’s $O(n^2)$-time algorithm [16], which is widely considered one of the gems of algorithmics and has inspired the discovery of general techniques for speeding up dynamic-programming algorithms [8 §7.1].

But most optimal-search-tree results, including Knuth’s, consider only trees based on three-way comparisons [16]. In contrast, standard high-level programming languages...
implement two-way comparisons (e.g., =, ≤, <). As pointed out by Knuth himself in the second edition of “The Art of Computer Programming”, search trees based on two-way comparisons are desirable [17] §6.2.2 ex. 33. Yet they are not as well understood: there is only one published proof that optimal two-way-comparison trees can be computed in polynomial time, given decades after Knuth’s result, in a breakthrough by Andersson et al. [1], who gave an $O(n^4)$-time algorithm for the successful-queries variant, in which only searches for keys stored in the tree are supported. Their work was motivated by an application to efficient message dispatching in object-oriented programming languages, addressed earlier by Chambers and Chen [4] (see also [18]) who provided an $O(n^2)$-time heuristic (non-optimal) algorithm for constructing two-way-comparison trees. Independently, Andersson [2] had earlier looked at speeding up search trees by replacing three-way comparisons with two-way ones.

**Difficulties introduced by two-way comparisons.** Given a query $q$, the search for $q$ in a search tree starts at the root and traces a path to a leaf, at each node comparing $q$ to the node’s key, then following the edge corresponding to the outcome. A three-way comparison has three possible outcomes: the query is less than, equal to, or greater than the node’s key. In any three-way-comparison tree, the set of queries reaching any internal node $N$ is an open interval between some two keys. Each subtree solves a subproblem defined by some such inter-key interval. This leads to a simple dynamic program with $O(n^2)$ subproblems and to a simple $O(n^3)$-time algorithm, the derivation of which is a standard exercise [8, §15.5], [9, Ex. 6.20]. Knuth’s celebrated result improves the running time to $O(n^2)$.

In contrast (as noted e.g. in [1]), algorithms for two-way-comparison search trees face the following obstacle. After a few equality tests, the subproblem that remains is defined by some inter-key interval with holes — each hole corresponding to an earlier equality test. (For example, in tree (b) of Figure 1 the set of queries reaching node $\langle q < 5 \rangle$ is $[3, \infty) - \{7\}$. In tree (b) of Figure 2 the set of queries reaching node $\langle q = 5 \rangle$ is $\{1, 5, 7\}$.)

The resulting dynamic-programming formulation can have $\Theta(2^n)$ subproblems, as the query sets that can arise can contain any subset of the keys.

**The most-likely-key property.** To achieve their polynomial-time algorithm for two-way comparison trees, Anderson et al. circumvent this obstacle in two steps. First, using an elegant side-weights argument, they show that every instance has an optimal tree $T$ with what we call the **most-likely-key property:**

(MLK) In each equality-test node $N$ in $T$, the test is to some key of maximum weight among keys reaching $N$.

Their proof is for the successful-queries variant but extends directly to the unrestricted variant. The result implies that it suffices to consider only trees with the MLK property.

If all key weights are distinct, then, at any node $N$ in such a tree $T$, the set of queries that reach $N$ is an inter-key interval with the $g$ heaviest keys removed, for some $g \in \{0, 1, \ldots, n\}$. There are $O(n^3)$ such query sets (intervals with such keys removed)
so this yields a dynamic program with $O(n^3)$ subproblems, and a simple $O(n^4)$-time algorithm for instances with distinct weights.

When the weights are not distinct, however, the MLK property gives no guidance about how to break ties. Even in trees with the MLK property, exponentially many distinct query sets can arise, so the dynamic program still has exponential size. To work around this counter-intuitive obstacle, Anderson et al. then resort to a lengthy argument that establishes “thresholds on the frequency of an item that can occur as an equality comparison at the root of an optimal tree.” This significantly complicates their analysis and their algorithm.

**New results.** Our first result is a structural theorem (Theorem 1), which implies that there is an optimal tree with what we call the **refined MLK property**:

(RMLK) *In each equality-test node $N$ in $T$, the test is to the key $k_b$ of maximum weight among keys reaching $N$, breaking ties by maximizing the index $b$.*

Breaking ties by the index $b$ (above) is for convenience, in fact any tie-breaking rule will work. The proof implicitly leverages Anderson et al.’s side-weights technique, augmented by a novel (to our knowledge) perturbation argument.

The theorem implies that, whether or not the weights are distinct, it suffices to consider trees with the RMLK property. Restricting to such trees yields a simple dynamic program with $O(n^3)$ subproblems, each solvable in $O(n)$ time, yielding an $O(n^4)$-time algorithm (Corollary 1). The algorithm is essentially the algorithm for distinct-weights instances, modified to break ties arbitrarily. Our formulation is simple and easy to implement — Appendix B gives a twenty-two-line implementation in Python. This is the first polynomial-time algorithm to be proven correct for the unrestricted variant.

**Other related work.** In 1994, Spuler [19, 20] suggested an algorithm for computing optimal two-way comparison trees, based on a result by Huang et al. [15] for a different...
Figure 2: Optimal two-way-comparison trees for the *successful-queries* variant (a) using only inequalities, and (b) when ‘=’s are also permitted. Both trees have key set \{1, 3, 5, 7\}. The number below each leaf represents its frequency. The cost of tree (a) is 2, the cost of tree (b) is 1.11.

The technical obstacle introduced by non-distinct keys also arises in finding optimal so-called *binary split trees*: for distinct-weight instances, an optimal binary split tree can be found in \(O(n^4)\) time, but for arbitrary instances the best bound known is \(O(n^5)\) \[12\]. For *generalized binary split trees* (see \[15, 6\]) no correct polynomial-time algorithm is yet known.

There are various other models for comparison-based search trees in the literature. As one example, one can consider only inequality comparisons. In trees with three-way comparisons that would correspond to restricting all \(\beta_i\)’s to be 0. As noted by Knuth, this model is equivalent to alphabetic coding, solved by Gilbert and Moore in \(O(n^3)\) time \[11\], later improved to \(O(n \log n)\) time by Hu and Tucker \[13\] and, independently, by Garsia and Wachs \[10\]. As noted by Anderson et al. this is also equivalent to finding the optimal two-way-comparison tree in the successful queries case that only uses “<” comparisons.

For more background information on different models and algorithms (including approximations) for search trees we refer the reader to the 1997 tutorial by Nagaraj \[18\].

*Note:* Our conference paper \[5\] (see \[7\] for full proofs and erratum) contains related results on two-way comparison trees, including a more complicated precursor of our algorithm \[5, Theorem 1\]. Extended forms of other results from that conference paper will appear separately; see \[7\] and e.g. \[6\].
Formal problem statement. Let \( k_1, k_2, \ldots, k_n \) denote the \( n \) given keys, in increasing order. Let \( \beta_b \geq 0 \), for \( b = 1, 2, \ldots, n \), be the weight of key \( k_b \) and \( \alpha_a \geq 0 \), for \( a = 0, 1, \ldots, n \), be the weight of interval \(( k_a, k_{a+1} )\), where for convenience \( k_0 \) and \( k_{n+1} \) (which are not keys) represent \(-\infty\) and \( \infty \). In standard applications the weights represent query frequencies or probabilities. Here each weight can be an arbitrary non-negative number.

As illustrated in Figure 1(b), a two-way-comparison search tree for the instance is a rooted binary tree \( T \), where each non-leaf node \( N \) represents a comparison (= or <) to some key \( k_b \). \( T \) has \( 2n + 1 \) leaves, \( n \) of which are associated one-to-one with the keys \( k_b \) and the remaining \( n + 1 \) with the intervals \(( k_a, k_{a+1} )\). Given any value \( q \), the search for \( q \) in \( T \) must correctly identify \( q \), in the following sense. The search starts at the root and proceeds down the tree, following the edges corresponding to the outcomes of the comparisons, and it terminates either at the leaf associated with a key \( k_b \), if \( q = k_b \), or at the leaf associated with the interval \(( k_a, k_{a+1} )\) containing \( q \), if \( q \) is not a key.

For each node \( N \) in \( T \), let \( T_N \) denote the subtree rooted at \( N \). The weight of subtree \( T_N \), denoted \( w_N \), is the total weight of the keys \( k_b \) and intervals \(( k_a, k_{a+1} )\) in the set of queries reaching \( N \) (these are the keys and intervals associated with the leaves of \( T_N \)).

The cost of \( T \) is the weighted depth of its leaves, that is:

\[
\text{cost}(T) = \sum_{\ell \in \text{leaves}(T)} \text{depth}_T(\ell) \times w_\ell = \sum_{N \in T-\text{leaves}(T)} w_N,
\]

where \( \text{leaves}(T) \) is the set of leaves in \( T \) and \( \text{depth}_T(\ell) \) is the number of edges on the path from the root to \( \ell \) in \( T \).

A search tree \( T \) is called optimal for a given instance if it has minimum cost among all search trees for the given instance. The objective is to compute such an optimal tree.

The successful-queries variant. The problem definition above is for the “full” variant, with unrestricted queries. In the successful-queries variant, illustrated in Figure 2, only queries to the given keys \( k_1, k_2, \ldots, k_n \) are allowed. An instance is specified by just those keys and their frequencies \( \beta_1, \beta_2, \ldots, \beta_n \), and the problem is to find the minimum cost of any tree that handles \( S = \{ k_1, \ldots, k_n \} \). (Section 2 formally defines “handles”.) While our algorithm’s description in Section 2 assumes the full variant, it can be adapted directly to the successful-queries variant, retaining its running time of \( O(n^4) \), by a straightforward modification of the dynamic program. This adapted algorithm has the same asymptotic running time as Anderson et al.’s (who considered only the successful case), but is much simpler.

Other comparison operators. Similarly, for the sake of exposition, we allow only the two comparison operators \{<, =\}. We remark without proof that it is straightforward to extend the problem, algorithm, and analysis to allow the general model with comparison set \{<, \leq, =\}, potentially getting a lower tree cost (even for the case \( n = 1 \)). Due to symmetries and equivalences, this is the only other set of operators of interest for the full variant of the problem (allowing unsuccessful queries) that includes equality tests. We note, in passing, that Hu and P. A. Tucker [14] have shown that the full variant with only
comparison operators in {<, ≤} allowed can be solved in time $O(n \log n)$ by reducing it to an instance of alphabetic coding with $2n + 1$ keys and applying the algorithm of Hu and A. C. Tucker [13].

2 The Dynamic-Programming Algorithm

Within any tree having the RMLK property, at any node $N$, the set $S$ of queries reaching $N$ has the following form: $S$ consists of the queries lying in some inter-key interval, minus some number of heaviest keys removed. The algorithm will solve one subproblem for each set $S$ of this form. Next we explicitly define these sets and their subproblems.

Fix a permutation $k_{\pi(1)}, k_{\pi(2)}, \ldots, k_{\pi(n)}$ of the keys that orders them by increasing weight, breaking ties arbitrarily. For each key $k_{\pi(r)}$, call $r$ the rank of the key. A set $S$ is valid if either (i) $S$ is a singleton key, of the form $S = \{k_b\}$, or (ii) $S$ is an inter-key interval with some largest-rank keys removed — that is, for some $i, j \in \{0, 1, \ldots, n + 1\}$ with $i < j$, and $h \in \{0, 1, \ldots, n\}$, the set $S$ equals $S(i, j, h)$, defined as

$$S(i, j, h) = [k_i, k_j] - \{k_{\pi(h+1)}, k_{\pi(h+2)}, \ldots, k_{\pi(n)}, k_0\}.$$  

(These sets are not necessarily distinct — for example, it is possible that $S(i, j, h) = S(i, j, h')$ even though, say, $h \neq h'$. Removing $k_0$ is needed only to deal with the case when $i = 0$.)

Note that $S(i, j, h)$ is a union of some keys and “failure” intervals $(k_a, k_{a+1})$. Let $w(S(i, j, h))$ denote the total weight of these keys and intervals. Given a search tree $T$, if those keys and intervals are also the ones associated with $T$’s leaves, say that $T$ handles $S(i, j, h)$ (so, for each $q \in S(i, j, h)$, the tree correctly identifies $q$). Define $\text{opt}(i, j, h)$ to be the minimum cost of any tree that handles $S(i, j, h)$. Note that $S(0, n + 1, n) = (-\infty, +\infty)$, so the goal is to return $\text{opt}(0, n + 1, n)$. (If desired, an optimal tree can also be constructed in the standard fashion.)

The algorithm computes $\text{opt}(0, n + 1, n)$ using the following recurrence relation:

(a) If $j = i + 1$ and $k_i \notin S(i, j, h)$, then $S(i, j, h) = (k_i, k_{i+1})$ so $\text{opt}(i, j, h) = 0$.

(b) If $h > 0$ and $k_{\pi(h)} \notin S(i, j, h)$, then $S(i, j, h) = S(i, j, h - 1)$, so $\text{opt}(i, j, h) = \text{opt}(i, j, h - 1)$.

(c) Else, $\text{opt}(i, j, h) = \min \left\{ w(S(i, j, h)) + \text{opt}(i, j, h - 1) \ (\text{if } h > 0, \text{ else } +\infty), \right.$

$$\min \{ w(S(i, j, h)) + \text{opt}(i, b, h) + \text{opt}(b, j, h) : i < b < j \}.$$  

Correctness. For any tree $T$ that handles $S(i, j, h)$, note that $\text{cost}(T)$ is zero if $T$ is a leaf, and otherwise $\text{cost}(T) = w(S(i, j, h)) + \text{cost}(T_L) + \text{cost}(T_R)$, where $T_L$ and $T_R$ are the left and right subtrees.

By inspection, the cases in the recurrence relation are exhaustive. Likewise, Cases (a) and (b) are correct. For Case (c), the two terms in the right-hand side correspond to creating a tree whose root does either (i) an equality-test with key $k_{\pi(h)}$ (if $h > 0$ so that $S(i, j, h)$ contains some keys, with $k_{\pi(h)}$ being the one of largest rank), or (ii) an
inequality comparison with some key $k_b$. By Theorem 1 in Section 3 there is an optimal tree of this form that handles $S(i, j, h)$, so the recurrence is correct in Case (c). (Note that if $j = i + 1$, the range of the minimum in the second term is empty, and there no trees of type (ii) above. In this case, we take the minimum to be infinite, so the recurrence remains correct.)

**Running time.** The number of subproblems $S(i, j, h)$ is $O(n^3)$. For each subproblem, by inspection of the recurrence, $\text{opt}(i, j, h)$ can be computed in $O(n)$ time. Thus the running time is $O(n^4)$. This proves Corollary 1.

**Corollary 1.** There is an $O(n^4)$-time algorithm for computing optimal two-way-comparison search trees (with unrestricted queries).

### 3 Statement and Proof of Theorem 1

Fix the keys $k_1, k_2, \ldots, k_n$ and their associated weights. Let $S$ be any valid set, and consider the subproblem of finding a minimum-cost (two-way-comparison) tree that handles $S$.

**Theorem 1.** Fix any key $k_b$ of largest weight among keys in $S$. Then some optimal tree $T$ satisfies one of the following three conditions:

(i) $T$ consists of a single leaf.

(ii) The root of $T$ does a less-than comparison.

(iii) The root of $T$ does an equality test to $k_b$.

**Proof.** The rest of this section gives the proof. It uses the following lemma as a black box.

**Lemma 1.** Some optimal tree $T$ satisfies one of the following three conditions:

(a) $T$ consists of a single leaf.

(b) The root of $T$ does a less-than comparison.

(c) The root of $T$ does an equality test to some key $k_b$ of maximum weight among keys in $S$.

(Anderson et al. prove essentially the same lemma (their Corollary 3), but for the successful-queries variant. Their proof uses their side-weights technique and extends essentially unchanged to our setting. For completeness we include the proof in Appendix 1.)

Now, for each integer $i \geq 1$, define a new instance with the same set of keys but with the weight of key $k_b$ increased by $1/i$. That is, let $\beta_b^i = \beta_b + 1/i$ and $\beta_a^i = \beta_a$ for all $a \neq b$. By Lemma 1 there is an optimal tree $T^i$ for the modified instance such
that $T^i$ satisfies conditions (a)–(c) of Lemma 1 for the modified instance, and therefore satisfies conditions (i)–(iii) of the theorem for the original instance. Let tree $T$ be such that $T = T^i$ for infinitely many $i$. (Such a $T$ must exist, because the possible trees for each modified instance are the same, and there are finitely many of them.) Then $T$ is a possible tree for the original instance and has the desired property. To finish we show that $T$ is optimal for the original instance.

Let $T^*$ be an optimal tree for the original instance. For all $i$ such that $T = T^i$,

$$\text{cost}(T, \beta) = \text{cost}(T, \beta^i) - \text{depth}_T(k_b)/i \quad \text{by definition of } \beta^i \text{'s and Equation (11)}$$

$$\leq \text{cost}(T, \beta^i)$$

$$= \text{cost}(T^i, \beta^i) \quad \text{because } T = T^i \text{ for this } i$$

$$\leq \text{cost}(T^*, \beta^i) \quad \text{because } T^i \text{ is optimal for } \beta^i \text{'s}$$

$$= \text{cost}(T^*, \beta) + \text{depth}_{T^*}(k_b)/i \quad \text{by definition of } \beta^i \text{'s and Equation (11)}.$$

(In this derivation, arguments $\beta$ or $\beta^i$ in the cost function indicate which key weights are used to compute it.) This holds as $i \to \infty$, so $\text{cost}(T, \beta) \leq \text{cost}(T^*, \beta)$, and thus $T$ is also optimal for the original instance. $\square$

Remarks. The correct statement of Theorem 1 is somewhat delicate. Consider the following statement: If there is an optimal tree whose root does an equality test, then, for any maximum-weight key $k$ in $S$, there is an optimal tree whose root does an equality test to $k$. This seemingly similar statement is false. Figure 3 gives a counterexample.

In that figure, keys 1 and 2 both have maximum weight among keys reaching the right child of the root. Trees (a) and (b) are optimal. In (a), the right child of the root does an equality test to key 1, but no optimal subtree for that subproblem (the subproblem with valid set $S = [1, \infty)$) has an equality test to key 2 at the root — tree (c) has minimum cost among trees with such a subtree. This counter-example also holds in the problem variant where all comparisons in $\{<, \leq, =\}$ are allowed.$^1$

The correct statement of Theorem 1 perturbs the weights infinitesimally, giving $k_b$ the unique largest weight and (via Lemma 1) forcing any root equality test to be to key $k_b$. It then observes that, because the minimum tree cost is a continuous function of the weights, any tree that is optimal for the (infinitesimally!) perturbed instance must also be optimal for the original instance. This infinitesimal perturbation can make all previously optimal trees with equality tests at the root non-optimal. (For example, in Figure 3 after an infinitesimal increase in $\beta_2$, only tree (b) would be optimal.)

The RMLK property from the introduction is a corollary of the theorem. To prove it, let $T$ be the optimal tree for $S$, with root as specified in the theorem. Then (using induction on $n$) replace each of the two subtrees by an optimal one (for its subproblem)

$^1$For the successful-queries variant no such counter-example is possible; that is, it can be shown that if some optimal tree has an equality test at the root, then for any maximum-weight key there is an optimal tree that uses this key in the equality-test at the root. (A proof can be provided upon request.) This stronger property, however, does not seem to have any algorithmic implications.
Figure 3: Trees for a full instance with $K = \{1, 2\}$, $(\beta_1, \beta_2) = (0.01, 0.01)$ and $(\alpha_0, \alpha_1, \alpha_2) = (0.9, 0.08, 0)$. As discussed in the remarks, the subproblem with valid set $S = [1, \infty)$ has an optimal solution with an equality test to key 1, but none with an equality test to key 2, even though $\beta_1 = \beta_2$.

having the RMLK property. The resulting tree $T'$ is also an optimal tree for $S$, and has the RMLK property.

In fact, there must be an optimal tree with the following stronger property. At any node $N$ that does an equality test, if there are $d$ largest-weight keys in the search space reaching $N$, then the top $d$ nodes on the right spine in $T_N$ do equality tests to those $d$ keys in some order (and permuting that order arbitrarily preserves optimality). This can be shown directly by a careful treatment of ties arising in Anderson et al.’s side-weights argument.

References


Appendices

A Proof of Lemma \[1\]

This appendix proves Lemma 1. First we state Anderson et al.'s side-weights lemma [1, Lemma 2], and prove it in our setting. (The proof is identical.) Fix the keys \(k_1, k_2, \ldots, k_n\) and associated weights. Let \(S\) be any valid set and consider the problem of finding a minimum-cost tree among trees that handle \(S\). Let \(T\) be any optimal tree for this problem.

Define the side-weight of any node \(N\) in \(T\), denoted \(sw(N)\), as follows. If \(N\) is a leaf, then \(sw(N) = 0\). If \(N\) does an equality test to a key, then \(sw(N)\) is the weight of that key. Otherwise (\(N\) is an inequality-comparison node) \(sw(N) = \min\{w_L, w_R\}\), where \(L\) and \(R\) are \(N\)'s left and right children. The utility lemma states that side-weights are monotone along paths from the root:

**Lemma 2 ([1, Lemma 2]).** If \(P\) is the parent of \(N\) in \(T\), then \(sw(P) \geq sw(N)\).

*Proof.* If \(N\) is a leaf, it has side-weight 0 and the property holds, so assume that \(N\) is not a leaf. There are four cases according to whether at each of \(P\) or \(N\) the comparison done is an equality or an inequality.

**Case 1:** Both \(P\) and \(N\) are inequality comparisons. Assume without loss of generality that \(N\) is the right child of \(P\). Let \(T_1\) be the subtree rooted at the child of \(P\) that is not \(N\). Let \(T_2\) and \(T_3\) be the subtrees rooted, respectively, at the left and right children of \(N\). Let \(\mu_i\) denote the weight of \(T_i\) (for \(i \in \{1, 2, 3\}\)). Now \(sw(P) = \min\{\mu_1, \mu_2 + \mu_3\}\) and \(sw(N) = \min\{\mu_2, \mu_3\}\). Now rotate \(N\) upward along the edge to its parent. While \(T_2\) stays at the same depth, \(T_1\) moves down and \(T_3\) moves up, each by one level. The increase in cost is \(\mu_1 - \mu_3\). This must be non-negative, as \(T\) is optimal, so \(\mu_1 \geq \mu_3\). This implies \(\min\{\mu_1, \mu_2 + \mu_3\} \geq \min\{\mu_2, \mu_3\}\), that is, \(sw(P) \geq sw(N)\).

**Case 2:** Node \(P\) is an inequality comparison and \(N\) is an equality comparison. Let \(T_1\), of weight, say, \(\mu_1\), be the subtree rooted at the child of \(P\) which is not \(N\). Let \(\mu_2\) be the weight of the left subtree, say \(T_2\), of \(N\) (consisting of a leaf for the equality-test key of \(N\)). Let \(\mu_3\) be the weight of the right subtree, say \(T_3\) of \(N\). We have \(sw(P) = \min\{\mu_1, \mu_2 + \mu_3\}\) and \(sw(N) = \mu_2\). Modify \(T\) by moving the equality test at \(N\) just above \(P\). Tree \(T_1\) moves down one level, \(T_3\) stays at the same level, yet \(T_2\), of weight \(\mu_2\), moves up one level. The net increase in cost is \(\mu_1 - \mu_2\). Since \(T\) is optimal, this is non-negative, so \(\mu_1 \geq \mu_2\). Thus \(\min\{\mu_1, \mu_2 + \mu_3\} \geq \mu_2\). That is, \(sw(P) \geq sw(N)\).

**Case 3:** Both \(P\) and \(N\) are equality comparisons. Swap the comparisons in \(P\) and \(N\). The increase in cost is \(sw(P) - sw(N)\). Since \(T\) is optimal, this is non-negative, so \(sw(P) \geq sw(N)\).

**Case 4:** Node \(P\) is an equality comparison and \(N\) is an inequality comparison. Let \(\mu_1\) be the weight of \(P\)'s left-subtree \(T_1\) (consisting of a leaf for the equality-test key of \(P\)). Let \(T_2\) and \(T_3\) of weight \(\mu_2\) and \(\mu_3\), respectively, be the subtrees hanging off \(N\). Rotate \(N\) upward, moving the equality-test node \(P\) just above the appropriate child of \(N\). Then exactly one of \(T_2\) and \(T_3\) moves up while \(T_1\) moves down. The increase in cost is either...
Since $T$ is optimal, the increase is non-negative, so $\mu_1 \geq \min\{\mu_2, \mu_3\}$. That is, $sw(P) \geq sw(N)$.

Next is the proof of Lemma 1. The proof is essentially the same as Anderson et al.'s proof of [1, Corollary 3], which is for successful-queries instances but extends directly to our setting.

**Lemma 1.** Some optimal tree $T$ satisfies one of the following three conditions:

(a) $T$ consists of a single leaf.

(b) The root of $T$ does a less-than comparison.

(c) The root of $T$ does an equality test to some key $k_b$ of maximum weight among keys in $S$.

**Proof.** Let $T$ be an optimal tree with root $N$. In the case that $N$ does an inequality comparison, we are done. So assume $N$ does an equality test to some key $k_b$, necessarily in $S$. Let $k_a$ be any largest-weight key in $S$. The parent of the leaf for $k_a$ is some node $P$ in $T$. First, if $P$ doesn’t do an equality test to $k_a$, replace the comparison at $P$ by an equality test to $k_a$, without otherwise changing $T$. This preserves correctness and optimality. Now, in the case that $P$ is the root $N$, we are done. In the remaining case, by Lemma 2 along the path from $N$ to $P$, the side-weights are non-increasing, so $\beta_b = sw(N) \geq sw(P) = \beta_a$, and $k_b$ is a largest-weight key in $S$.

**B Source code (Python)**

See Figure 4.
#!/usr/bin/env python3.7
from math import inf # (inf is infinity)
from functools import lru_cache
memoize = lru_cache(maxsize=None)

def opt_cost(K, alpha, beta):
    "Return min cost of any 2-way-comparison tree for (K, alpha, beta)."

    K_by_wt = sorted(range(1, len(K)+1), key=lambda i: beta[i-1])
    @memoize
    def opt(i, j, h):
        K_in_A_ijh = [b for b in K_by_wt[:h] if i <= b < j]
        w_ijh = sum(alpha[i:j] + [beta[k-1] for k in K_in_A_ijh])

        return (0 if j == i+1 and i not in K_in_A_ijh else opt(i, j, h-1) if h > 0 and K_by_wt[h-1] not in K_in_A_ijh else min(w_ijh + opt(i, j, h-1) if h > 0 else inf,
            min(w_ijh + opt(i, b, h) + opt(b, j, h) for b in range(i+1, j))
            if i+1 < j else inf)

    return opt(0, len(K)+1, len(K))

# Run the algorithm on the instance from Figure 1(b).
K = [1, 3, 5, 7]
beta = [0.08, 0.02, 0.08, 0.33]
alpha = [0.15, 0.21, 0.08, 0.04, 0.01]
assert opt_cost(K, alpha, beta) == 2.8

Figure 4: Python code for the algorithm for the full variant.